VEMU INSTITUTE OF TECHNOLOGY

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LECTURE NOTES

DEPARTMENT OF COMPUTER SCIENCE AND ENGINEERING

SUBJECT NAME: COMPILER DESIGN (20A05601T)

BRANCH: CSE

YEAR AND SEMESTER: III – II

COURSE: B. TECH

REGULATION: R20

UNIT – I

Introduction Language Processing, Structure of a compiler, the Evaluation of Programming language, The Science of building a Compiler application of Compiler Technology. Programming Language Basics.

Lexical Analysis-:The role of lexical analysis buffering, specification of tokens. Recognitions of tokens the lexical analyzer generator lexical

UNIT -1

TRANSLATOR

A translator is a program that takes as input a program written in one language and produces as output a program in another language. Beside program translation, the translator performs another very important role, the error-detection. Any violation of HLL specification would be detected and reported to the programmers. Important role of translator are:

1 Translating the HLL program input into an equivalent machine language program.
2 Providing diagnostic messages wherever the programmer violates specification of the HLL.

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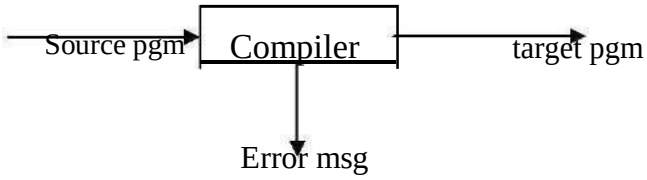
1 Translating the hll program input into an equivalent ml program.
2 Providing diagnostic messages wherever the programmer violates specification of the hll.

TYPE OF TRANSLATORS:-

a. **Compiler**
b. **Interpreter**
c. **Preprocessor**

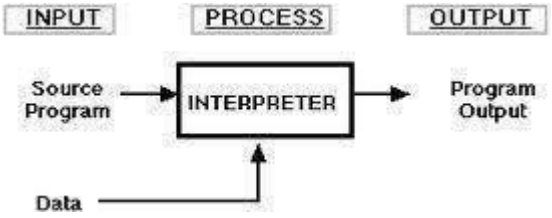
Compiler

Compiler is a translator program that translates a program written in (HLL) the source program and translate it into an equivalent program in (MLL) the target program. As an important part of a compiler is error showing to the programmer.



Executing a program written n HLL programming language is basically of two parts. The source program must first be compiled and translated into a object program. Then the resulting object program is loaded into a memory executed.

Interpreter*:* An interpreter is a program that appears to execute a source program as if it were machine language.



Languages such as BASIC, SNOBOL, LISP can be translated using interpreters. JAVA also uses interpreter. The process of interpretation can be carried out in following phases.

1. Lexical analysis
2. Synatx analysis
3. Semantic analysis
4. Direct Execution

Advantages:

Modification of user program can be easily made and implemented as execution proceeds.

5. Type of object that denotes a various may change dynamically.

Debugging a program and finding errors is simplified task for a program used for interpretation.

The interpreter for the language makes it machine independent.

Disadvantages:

The execution of the program is *slower*.

Memory consumption is more.

OVERVIEW OF LANGUAGE PROCESSING SYSTEM

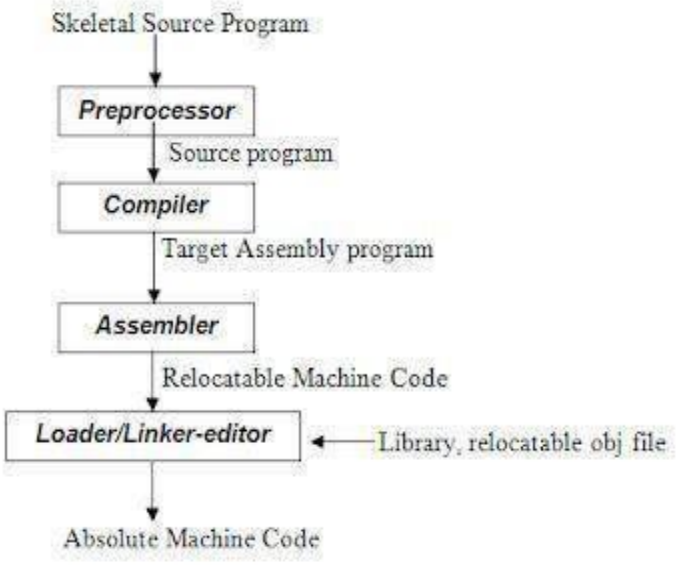


Fig 1.1 Language –processing System

Preprocessor

A preprocessor produce input to compilers. They may perform the following functions.

1. *Macro processing:* A preprocessor may allow a user to define macros that are short hands for longer constructs.
2. *File inclusion:* A preprocessor may include header files into the program text.
3. *Rational preprocessor:* these preprocessors augment older languages with more modern flow-of-control and data structuring facilities.
4. *Language Extensions:* These preprocessor attempts to add capabilities to the language by certain amounts to build-in macro

Assembler*:* programmers found it difficult to write or read programs in machine language. They begin to use a mnemonic (symbols) for each machine instruction, which they would subsequently translate into machine language. Such a mnemonic machine language is now called an assembly language. Programs known as assembler were written to automate the translation of assembly language in to machine language. The input to an assembler program is called source program, the output is a machine language translation (object program).

Loader and Link-editor:

Once the assembler procedures an object program, that program must be placed into memory and executed. The assembler could place the object program directly in memory and transfer control to it, thereby causing the machine language program to be execute. This would waste core by leaving the assembler in memory while the user‟s program was being executed. Also the programmer would have to retranslate his program with each execution, thus wasting translation time. To overcome this problems of wasted translation time and memory. System programmers developed another component called loader

"A loader is a program that places programs into memory and prepares them for execution." It would be more efficient if subroutines could be translated into object form the loader could"relocate" directly behind the user‟s program. The task of adjusting programs othey may be placed in arbitrary core locations is called relocation.

STRUCTURE OF A COMPILER

Phases of a compiler: A compiler operates in phases. A phase is a logically interrelated operation that takes source program in one representation and produces output in another representation. The phases of a compiler are shown in below

There are two phases of compilation.

a. Analysis (Machine Independent/Language Dependent)
b. Synthesis(Machine Dependent/Language independent)

Compilation process is partitioned into no-of-sub processes called p**hases'**.

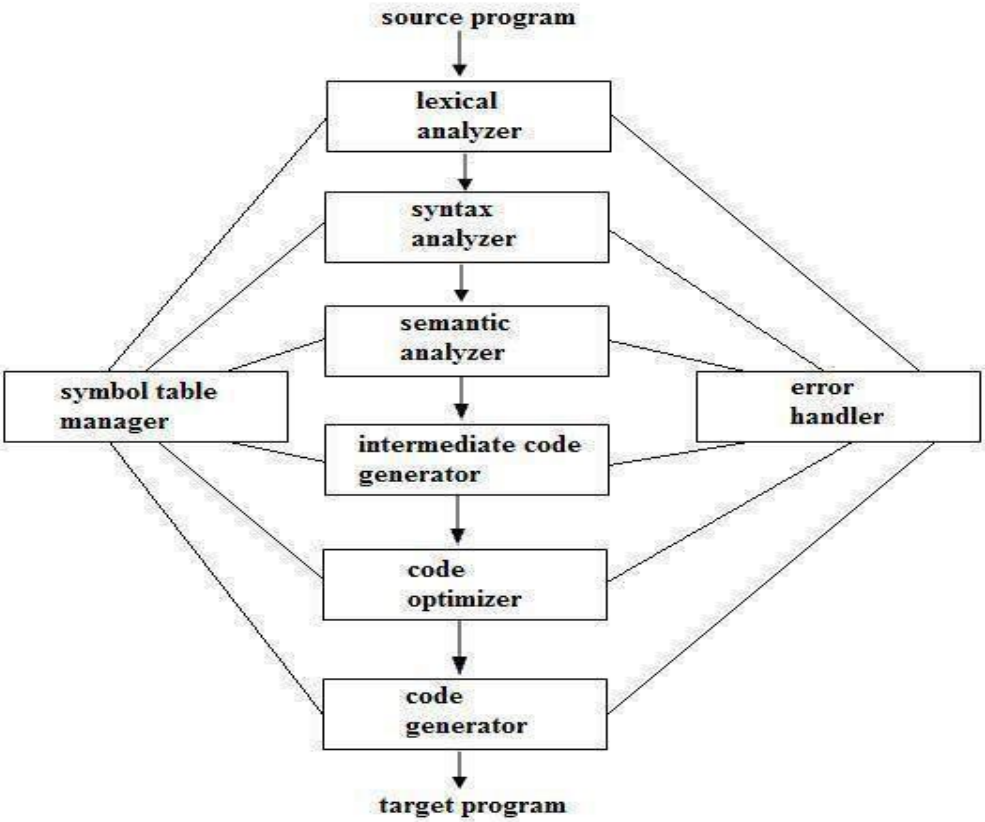


Fig 1.5 Phases of a compiler

Lexical Analysis:-

LA or Scanners reads the source program one character at a time, carving the source program into a sequence of automic units called **tokens.**

Syntax Analysis:-

The second stage of translation is called Syntax analysis or parsing. In this phase expressions, statements, declarations etc… are identified by using the results of lexical analysis. Syntax analysis is aided by using techniques based on formal grammar of the programming language.

Intermediate Code Generations:-

An intermediate representation of the final machine language code is produced. This phase bridges the analysis and synthesis phases of translation.

Code Optimization :-

This is optional phase described to improve the intermediate code so that the output runs faster and takes less space.

Code Generation:-

The last phase of translation is code generation. A number of optimizations to **reduce the length of machine language program** are carried out during this phase. The output of the code generator is the machine language program of the specified computer.

Table Management (or) Book-keeping:-

This is the portion to **keep the names** used by the program and records essential information about each. The data structure used to record this information called a „Symbol Table".

Error Handlers:-

It is invoked when a flaw error in the source program is detected.

The output of **LA** is a stream of tokens, which is passed to the next phase, the syntax analyzer or parser. The SA groups the tokens together into syntactic structure called as **expression.** Expression may further be combined to form statements. The syntactic structure can be regarded as a tree whose leaves are the token called as parse trees.

The parser has two functions. It checks if the tokens from lexical analyzer, occur in pattern that are permitted by the specification for the source language. It also imposes on tokens a tree-like structure that is used by the sub-sequent phases of the compiler.

Example, if a program contains the expression **A+/B** after lexical analysis this expression might appear to the syntax analyzer as the token sequence **id+/id.** On seeing the /, the syntax analyzer should detect an error situation, because the presence of these two adjacent binary operators violates the formulations rule of an expression.

Syntax analysis is to make explicit the hierarchical structure of the incoming token stream by **identifying which parts of the token stream should be grouped**.

Example, (A/B*C has two possible interpretations.)

1, divide A by B and then multiply by C or

2, multiply B by C and then use the result to divide A.

each of these two interpretations can be represented in terms of a parse tree.

Intermediate Code Generation:-

The intermediate code generation uses the structure produced by the syntax analyzer to create a stream of simple instructions. Many styles of intermediate code are possible. One common style uses instruction with one operator and a small number of operands.

The output of the syntax analyzer is some representation of a parse tree. the intermediate code generation phase transforms this parse tree into an intermediate language representation of the source program.

Code Optimization

This is optional phase described to improve the intermediate code so that the output runs faster and takes less space. Its output is another intermediate code program that does the some job as the original, but in a way that saves time and / or spaces.

1, Local Optimization:-

There are local transformations that can be applied to a program to make an improvement. For example,

If **A** > **B** goto **L2**

Goto **L3**

L2 :

This can be replaced by a single statement

If **A < B** goto **L3**

Another important local optimization is the elimination of common sub-expressions

A := B + C + D
E := B + C + F

Might be evaluated as

T1 := B + C

A := T1 + D
E := T1 + F

Take this advantage of the common sub-expressions **B + C.**

2, Loop Optimization:-

Another important source of optimization concerns about **increasing the speed of loops**. A typical loop improvement is to move a computation that produces the same result each time around the loop to a point, in the program just before the loop is entered.

Code Generator :-

Code Generator produces the object code by deciding on the memory locations for data, selecting code to access each datum and selecting the registers in which each computation is to be done. Many computers have only a few high speed registers in which computations can be performed quickly. A good code generator would attempt to utilize registers as efficiently as possible.

Table Management OR Book-keeping :-

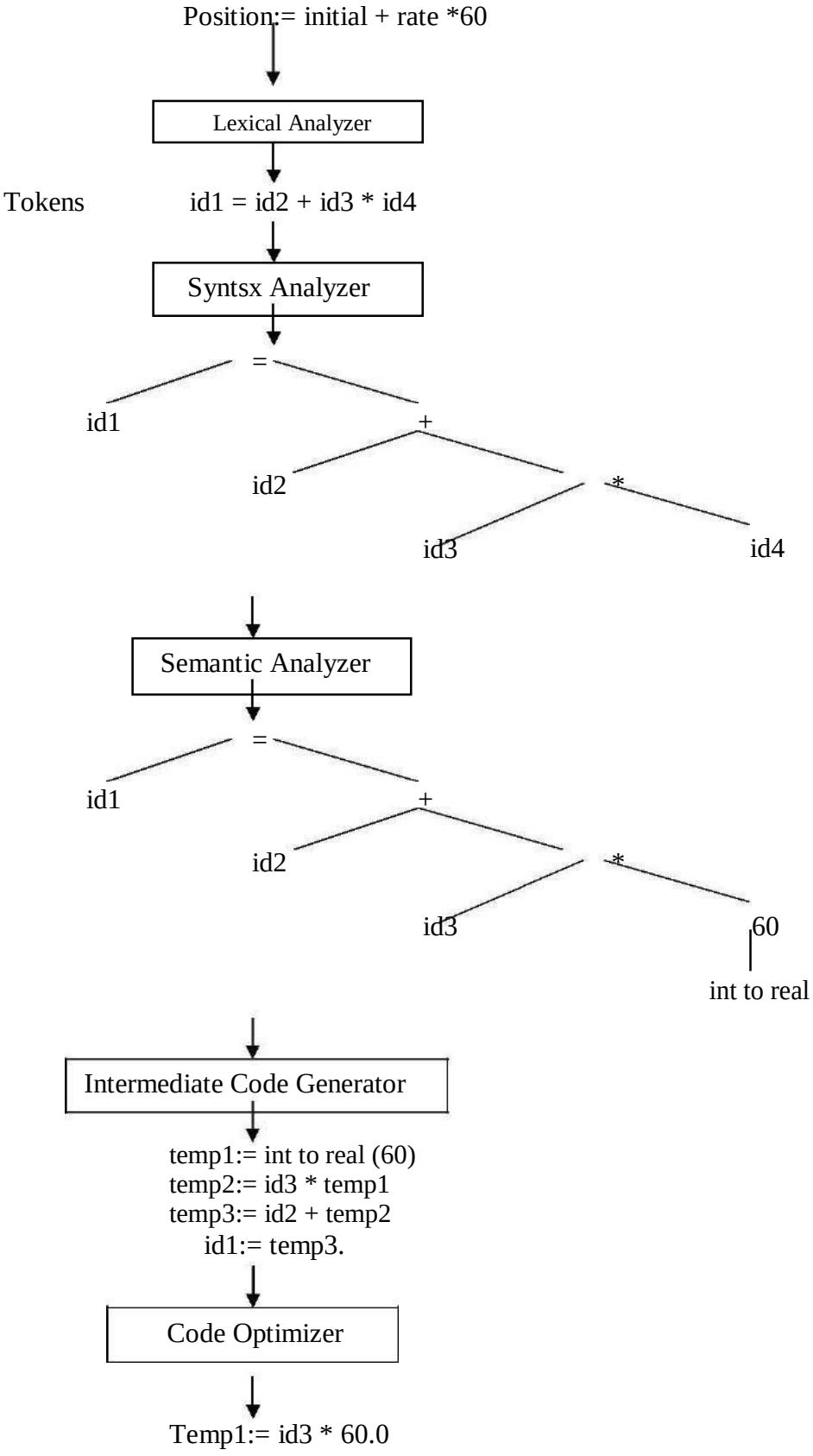
A compiler needs to collect information about all the data objects that appear in the source program. The information about data objects is collected by the early phases of the compiler-lexical and syntactic analyzers. The data structure used to record this information is called as Symbol Table.

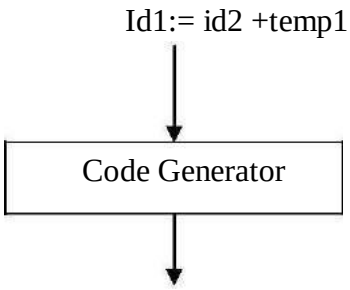
Error Handing :-

One of the most important functions of a compiler is the detection and reporting of errors in the source program. The error message should allow the programmer to determine exactly where the errors have occurred. Errors may occur in all or the phases of a compiler.

Whenever a phase of the compiler discovers an error, it must report the error to the error handler, which issues an appropriate diagnostic msg. Both of the table-management and error-Handling routines interact with all phases of the compiler.

Example:


Position:= initial + rate *60
Lexical Analyzer
Tokens
id1 = id2 + id3 * id4
Syntsx Analyzer
=
id1
+
id2
*
id3
id4
Semantic Analyzer
=
id1
+
id2
*
id3
60
int to real
Intermediate Code Generator
temp1:= int to real (60)
temp2:= id3 * temp1
temp3:= id2 + temp2
id1:= temp3.
Code Optimizer
Temp1:= id3 * 60.0




```
MOVF id3, r2
MULF *60.0, r2
MOVF   id2, r2
ADDF   r2, r1
MOVF   r1, id1
```

Evolution of Programming languages

The **history of programming languages** spans from documentation of early mechanical computers to modern tools for software development. Early programming languages were highly specialized, relying on mathematical notation

The move to Higher Level Languages

The first step towards more people friendly programming languages was the development of mnemonic assembly languages in the early 1950's.The instructions in assembly languages were just mnemonic representations of machine instructions.
A major step towards higher level languages was made in the later half of the 1950's with the development of FORTRAN for scientific computation, Cobol for business data Processing and Lisp for symbolic computation.

In the following decades many more languages were created with innovative features to help make programming easier, more natural, and more robust.

Languages can also be classified in variety of ways.

Classification by Generation: Ist generation are the machine languages, 2nd generation are the assembly languages, 3rd generation are the higher level languages like Fortran,cobol, Lisp,C etc.4th generation are the languages designed for specific application like NOMAD,SQL,POST The term fifth generation language has been applied to logic and the constraint based language like prolog and OPS5.

Classification by the use: imperative languages in which your program specifies How computation is to be done the declarative for languages in which your program specifies what computation is to be done.

Examples:

Imperative languages: C,C++,C#,Java.
Declarative languages: ML, Haskell, Prolog

Object oriented language is one that supports Object oriented programming, a Programming style in which a program consists of a collection of objects that interact with one another.

Examples: Simula 67, small talk, C ++, Java,Ruby

Scripting languages are interpreted languages with high level operators designed for "gluing together" computations These computations originally called Scripts

Example: JavaScript, Perl, PHP, python, Ruby, TCL

The Science of building a Compiler

A compiler must accept all source programs that conform to the specification of the language; the set of source programs is infinite and any program can be very large, consisting of possibly millions of lines of code. Any transformation performed by the compiler while translating a source program must preserve the meaning of the program being compiled. Compiler writers thus have influence over not just the compilers they create, but all the programs that their compilers compile. This leverage makes writing compilers particularly rewarding; however, it also makes compiler development challenging.

Modelling in compiler design and implementation: The study of compilers is mainly a study of how we design the right mathematical models and choose the right algorithms.Some of most fundamental models are finite-state machines and regular expressions.These models are useful for de-scribing the lexical units of programs (keywords, identifiers, and such) and for describing the algorithms used by the compiler to recognize those units. Also among the most fundamental models are context-free grammars, used to describe the syntactic structure of programming languages such as the nesting of parentheses or control constructs. Similarly, trees are an important model for representing the structure of programs and their translation into object code.

The science of code optimization: The term "optimization" in compiler design refers to the attempts that a com-piler makes to produce code that is more efficient than the obvious code. In modern times, the optimization of code that a compiler performs has become both more important and more complex. It is more complex because processor architectures have become more complex, yielding more opportunities to improve the way code executes. It is more important because massively par-allel computers require substantial optimization, or their performance suffers by orders of magnitude.

Compiler optimizations must meet the following design objectives:

1. The optimization must be correct, that is, preserve the meaning of the compiled program,
2. The optimization must improve the performance of many programs,
3. The compilation time must be kept reasonable, and
4. The engineering effort required must be manageable.

Thus, in studying compilers, we learn not only how to build a compiler, but also the general methodology of solving complex and open-ended problems.

Applications of Compiler Technology

Compiler design impacts several other areas of computer science.

Implementation of high-level programming language: A high-level programming language defines a programming abstraction: the programmer expresses an algorithm using the language, and the compiler

must translate that program to the target language. higher-level programming languages are easier to program in, but are less efficient, that is, the target programs run more slowly. Programmers using a low-level language have more control over a computation and can, in principle, produce more efficient code.

Language features that have stimulated significant advances in compiler technology.

Practically all common programming languages, including C, Fortran and Cobol, support user-defined aggregate data types, such as arrays and structures, and high-level control flow, such as loops and procedure invocations. If we just take each high-level construct or data-access operation and translate it directly to machine code, the result would be very inefficient. A body of compiler optimizations, known as *data-flow optimizations*, has been developed to analyze the flow of data through the program and removes redundancies across these constructs. They are effective in generating code that resembles code written by a skilled programmer at a lower level.

Object orientation was first introduced in Simula in 1967, and has been incorporated in languages such as Smalltalk, C + + , C # , and Java. The key ideas behind object orientation are

1. Data abstraction and
2. Inheritance of properties,

Java has many features that make programming easier, many of which have been introduced previously in other languages. Compiler optimizations have been developed to reduce the overhead, for example, by eliminating unnecessary range checks and by allocating objects that are not accessible beyond a procedure on the stack instead of the heap. Effective algorithms also have been developed to minimize the overhead of garbage collection.

In dynamic optimization, it is important to minimize the compilation time as it is part of the execution overhead. A common technique used is to only compile and optimize those parts of the program that will be frequently executed.

Optimizations for Computer Architecture: high-performance systems take advantage of the same two basic techniques: *parallelism* and *memory hierarchies.* Parallelism can be found at several levels: at the *instruction level*, where multiple operations are executed simultaneously and at the *processor level*, where different threads of the same application are run on different processors. Memory hierarchies are a response to the basic limitation that we can build very fast storage or very large storage, but not storage that is both fast and large.

Design of New Computer Architectures: in modern computer architecture development, compilers are developed in the processor-design stage, and compiled code, running on simulators, is used to evaluate the proposed architectural features. One of the best known examples of how compilers influenced the design of computer architecture was the invention of the RISC (Reduced Instruction-Set Computer) architecture.

Compiler optimizations often can reduce these instructions to a small number of simpler operations by eliminating the redundancies across complex instructions. Thus, it is desirable to build simple instruction sets; compilers can use them effectively and the hardware is much easier to optimize. Most general-purpose processor architectures, including PowerPC, SPARC, MIPS, Alpha, and PA-RISC, are based on the RISC concept.

Specialized Architectures Over the last three decades, many architectural concepts have been proposed. They include data flow machines, vector machines, VLIW (Very Long Instruction Word) machines, SIMD (Single Instruction, Multiple Data) arrays of processors, systolic arrays, multiprocessors with shared memory, and multiprocessors with distributed memory. The development of each of these architectural concepts was accompanied by the research and development of corresponding compiler technology.

Program Translations: The following are some of the important applications of program-translation techniques.

Binary Translation: Compiler technology can be used to translate the binary code for one machine to that of another, allowing a machine to run programs originally compiled for another instruction set. Binary translation technology has been used by various computer companies to increase the availability of software for their machines.

Hardware Synthesis: Not only is most software written in high-level languages; even hardware de-signs are mostly described in high-level hardware description languages like Verilog and VHDL. Hardware designs are typically described at the register trans-fer level (RTL), where variables represent registers and expressions represent combinational logic.

Database Query Interpreters: Besides specifying software and hardware, languages are useful in many other applications. For example, query languages, especially SQL (Structured Query Language), are used to search databases. Database queries consist of predicates containing relational and boolean operators. They can be interpreted or com-piled into commands to search a database for records satisfying that predicate.

Programming Language Basics:

1 The Static/Dynamic Distinction

2 Environments and States

3 Static Scope and Block Structure

4 Explicit Access Control

5 Dynamic Scope

6 Parameter Passing Mechanisms

The Static/Dynamic Distinction: Among the most important issues that we face when designing a compiler for a language is what decisions can the compiler make about a program. If a language uses a policy that allows the compiler to decide an issue, then we say that the language uses a *static* policy or that the issue can be decided at *compile time.* On the other hand, a policy that only allows a decision to be made when we execute the program is said to be a *dynamic policy.* One issue is the scope of declarations. The *scope* of a declaration of *x* is the region of the program in which uses of *x* refer to this declaration. A language uses *static scope* or *lexical scope* if it is possible to determine the scope of a declaration by looking only at the program. Otherwise, the language uses *dynamic scope.* With dynamic scope, as the program runs, the same use of *x* could refer to any of several different declarations of *x.*

Environments and States:

The *environment* is a mapping from names to locations in the store. Since variables refer to locations, we could alternatively define an environment as a mapping from names to variables.

The *state* is a mapping from locations in store to their values. That is, the state maps 1-values to their corresponding r-values, in the terminology of C. Environments change according to the scope rules of a language.

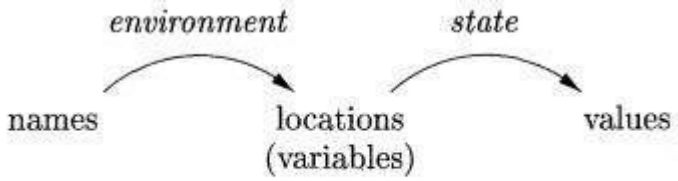


Figure 1.8: Two-stage mapping from names to values

Static Scope and Block Structure

Most languages, including C and its family, use static scope. we consider static-scope rules for a language with blocks, where a *block* is a grouping of declarations and statements. C uses braces { and } to delimit a block; the alternative use of **begin** and **end** for the same purpose dates back to Algol.

A C program consists of a sequence of top-level declarations of variables and functions.Functions may have variable declarations within them, where variables include local variables and parameters. The scope of each such declaration is restricted to the function in which it appears. The scope of a top-level declaration of a name *x* consists of the entire program that follows, with the exception of those statements that lie within a function that also has a declaration of *x.*

A block is a sequence of declarations followed by a sequence of statements, all surrounded by braces. a declaration *D* "belongs" to a block *B* if *B* is the most closely nested block containing *D;* that is, *D* is located within *B,* but not within any block that is nested within *B.* The static-scope rule for variable declarations in a block-structured lan-guages is as follows. If declaration *D* of name *x* belongs to block *B,* then the scope of D is all of B, except for any blocks B' nested to any depth within J5, in which x is redeclared. Here, x is redeclared in B' if some other declaration D' of the same name x belongs to B'.

An equivalent way to express this rule is to focus on a use of a name x. Let Bi, i?2, • • • , Bk be all the blocks that surround this use of x, with Bk the smallest, nested within Bk-i, which is nested within Bk-2, and so on. Search for the largest i such that there is a declaration of x belonging to B^. This use of x refers to the declaration in B{. Alternatively, this use of x is within the scope of the declaration in Bi.

Explicit Access Control

Through the use of keywords like **public, private,** and **protected,** object-oriented languages such as C + + or Java provide explicit control over access to member names in a superclass. These keywords support *encapsulation* by restricting access. Thus, private names are purposely given a scope that includes only the method declarations and definitions associated with that class and any "friend" classes (the C + + term). Protected names are accessible to subclasses. Public names are accessible from outside the class.

Dynamic Scope

Any scoping policy is dynamic if it is based on factor(s) that can be known only when the program executes. The term *dynamic scope,* however, usually refers to the following policy: a use of a name *x* refers to the declaration of *x* in the most recently called procedure with such a declaration. Dynamic scoping of this type appears only in special situations. We shall consider two ex-amples of

dynamic policies: macro expansion in the C preprocessor and method resolution in object-oriented programming.

Declarations and Definitions

Declarations tell us about the types of things, while definitions tell us about their values. Thus, i n t i is a declaration of i, while i = 1 is a definition of i.

The difference is more significant when we deal with methods or other procedures. In C + + , a method is declared in a class definition, by giving the types of the arguments and result of the method (often called the signature for the method. The method is then defined, i.e., the code for executing the method is given, in another place. Similarly, it is common to define a C function in one file and declare it in other files where the function is used.

Parameter Passing Mechanisms

In this section, we shall consider how the *actual parameters* (the parameters used in the call of a procedure) are associated with the *formal parameters* (those used in the procedure definition). Which mechanism is used determines how the calling-sequence code treats parameters. The great majority of languages use either "call-by-value," or "call-by-reference," or both.

Call - by - Value

In *call-by-value,* the actual parameter is evaluated (if it is an expression) or copied (if it is a variable). The value is placed in the location belonging to the corresponding formal parameter of the called procedure. This method is used in C and Java, and is a common option in C + + , as well as in most other languages. Call-by-value has the effect that all computation involving the formal parameters done by the called procedure is local to that procedure, and the actual parameters themselves cannot be changed.

Note, however, that in C we can pass a pointer to a variable to allow that variable to be changed by the callee. Likewise, array names passed as param eters in C, C + + , or Java give the called procedure what is in effect a pointer or reference to the array itself. Thus, if a is the name of an array of the calling procedure, and it is passed by value to corresponding formal parameter *x,* then an assignment such as x [i] = 2 really changes the array element a[2]. The reason is that, although *x* gets a copy of the value of *a,* that value is really a pointer to the beginning of the area of the store where the array named *a* is located.

Similarly, in Java, many variables are really references, or pointers, to the things they stand for. This observation applies to arrays, strings, and objects of all classes. Even though Java uses call-by-value exclusively, whenever we pass the name of an object to a called procedure, the value received by that procedure is in effect a pointer to the object. Thus, the called procedure is able to affect the value of the object itself.

Call - by - Reference

In *call-by-reference,* the address of the actual parameter is passed to the callee as the value of the corresponding formal parameter. Uses of the formal parameter in the code of the callee are implemented by following this pointer to the location indicated by the caller. Changes to the formal parameter thus appear as changes to the actual parameter.

If the actual parameter is an expression, however, then the expression is evaluated before the call, and its value stored in a location of its own. Changes to the formal parameter change this location, but can have no effect on the data of the caller.

Call-by-reference is used for "ref" parameters in C + + and is an option in many other languages. It is almost essential when the formal parameter is a large object, array, or structure. The reason is that strict call-by-value requires that the caller copy the entire actual parameter into the space belonging to the corresponding formal parameter. This copying gets expensive when the parameter is large. As we noted when discussing call-by-value, languages such as Java solve the problem of passing arrays, strings, or other objects by copying only a reference to those objects. The effect is that Java behaves as if it used call-by-reference for anything other than a basic type such as an integer or real.

Call - by - Name

A third mechanism — call-by-name — was used in the early programming language Algol **60.** It requires that the callee execute as if the actual parameter were substituted literally for the formal parameter in the code of the callee, as if the formal parameter were a macro standing for the actual parameter (with renaming of local names in the called procedure, to keep them distinct). When the actual parameter is an expression rather than a variable, some unintuitive behaviors occur, which is one reason this mechanism is not favored today.

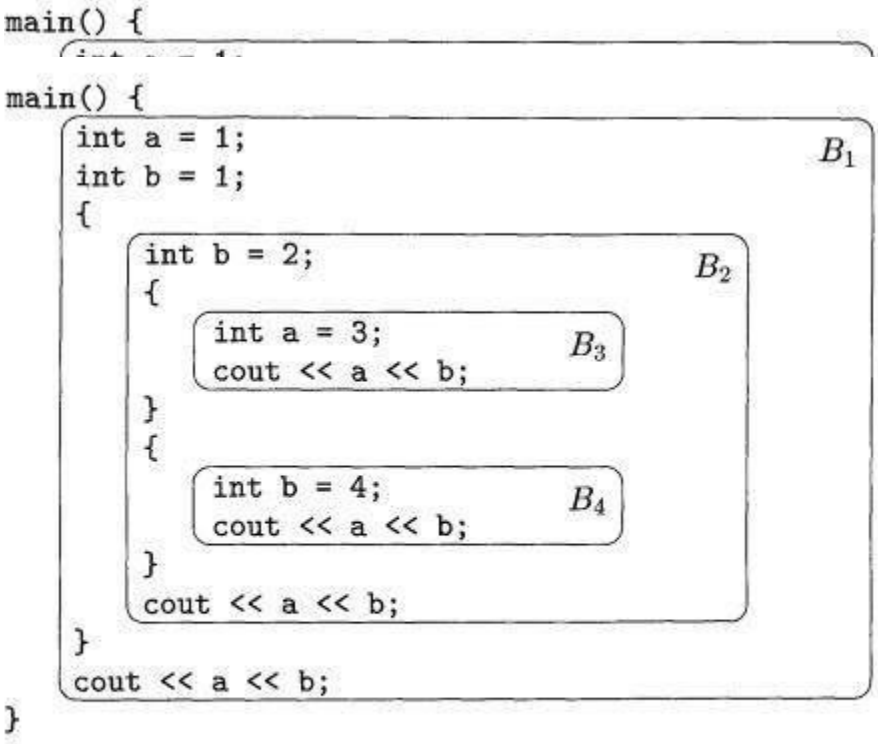


Figure 1.10: Blocks in a C++ program

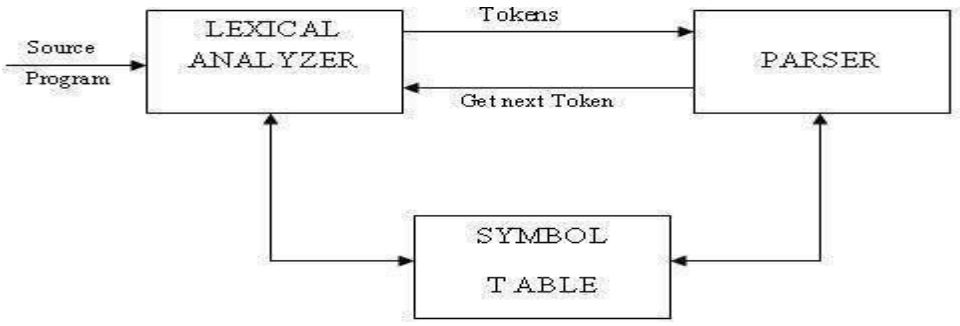
LEXICAL ANALYSIS

OVER VIEW OF LEXICAL ANALYSIS

- To identify the tokens we need some method of describing the possible tokens that can appear in the input stream. For this purpose we introduce regular expression, a notation that can be used to describe essentially all the tokens of programming language.
- Secondly , having decided what the tokens are, we need some mechanism to recognize these in the input stream. This is done by the token recognizers, which are designed using transition diagrams and finite automata.

ROLE OF LEXICAL ANALYZER

the LA is the first phase of a compiler. It main task is to read the input character and produce as output a sequence of tokens that the parser uses for syntax analysis.



Upon receiving a get next token command form the parser, the lexical analyzer reads the input character until it can identify the next token. The LA return to the parser representation for the token it has found. The representation will be an integer code, if the token is a simple construct such as parenthesis, comma or colon.

LA may also perform certain secondary tasks as the user interface. One such task is striping out from the source program the commands and white spaces in the form of blank, tab and new line characters. Another is correlating error message from the compiler with the source program.

LEXICAL ANALYSIS VS PARSING:

Lexical analysis	Parsing
A Scanner simply turns an input String (say a file) into a list of tokens. These tokens represent things like identifiers, parentheses, operators etc. The lexical analyzer (the "lexer") parses individual symbols from the source code file into tokens. From there, the "parser" proper turns those whole tokens into sentences of your grammar	A parser converts this list of tokens into a Tree-like object to represent how the tokens fit together to form a cohesive whole (sometimes referred to as a sentence). A parser does not give the nodes any meaning beyond structural cohesion. The next thing to do is extract meaning from this structure (sometimes called contextual analysis).

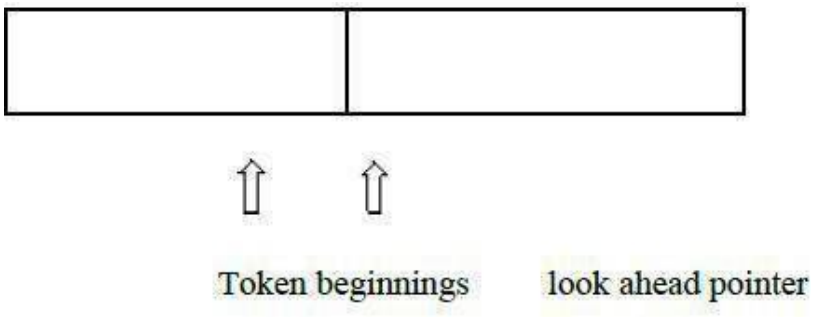
INPUT BUFFERING

The LA scans the characters of the source pgm one at a time to discover tokens. Because of large amount of time can be consumed scanning characters, specialized buffering techniques have been developed to reduce the amount of overhead required to process an input character.

Buffering techniques:

1. Buffer pairs
2. Sentinels

The lexical analyzer scans the characters of the source program one a t a time to discover tokens. Often, however, many characters beyond the next token many have to be examined before the next token itself can be determined. For this and other reasons, it is desirable for thelexical analyzer to read its input from an input buffer. Figure shows a buffer divided into two haves of, say 100 characters each. One pointer marks the beginning of the token being discovered. A look ahead pointer scans ahead of the beginning point, until the token is discovered .we view the position of each pointer as being between the character last read and thecharacter next to be read. In practice each buffering scheme adopts one convention either apointer is at the symbol last read or the symbol it is ready to read.



Token beginnings look ahead pointerThe distance which the lookahead pointer may have to travel past the actual token may belarge. For example, in a PL/I program we may see: DECALRE (ARG1, ARG2… ARG *n*) Without knowing whether DECLARE is a keyword or

an array name until we see the character that follows the right parenthesis. In either case, the token itself ends at the second E. If the look ahead pointer travels beyond the buffer half in which it began, the other half must be loaded with the next characters from the source file. Since the buffer shown in above figure is of limited size there is an implied constraint on how much look ahead can be used before the next token is discovered. In the above example, ifthe look ahead traveled to the left half and all the way through the left half to the middle, we could not reload the right half, because we would lose characters that had not yet been groupedinto tokens. While we can make the buffer larger if we chose or use another buffering scheme,we cannot ignore the fact that overhead is limited.

TOKEN, LEXEME, PATTERN:

Token: Token is a sequence of characters that can be treated as a single logical entity. Typical tokens are,

1) Identifiers 2) keywords 3) operators 4) special symbols 5)constants

Pattern: A set of strings in the input for which the same token is produced as output. This set of strings is described by a rule called a pattern associated with the token.

Lexeme: A lexeme is a sequence of characters in the source program that is matched by the pattern for a token.

Example:

Description of token

Token	**lexeme**	**pattern**
const	const	const
if	if	If
relation	<,<=,= ,< >,>=,>	< or <= or = or < > or >= or letter followed by letters & digit
i	pi	any numeric constant
nun	3.14	any character b/w “and “except"
literal	"core"	pattern

A pattern is a rule describing the set of lexemes that can represent a particular token in source program.

LEXICAL ERRORS:

Lexical errors are the errors thrown by your lexer when unable to continue. Which means that there's no way to recognise a *lexeme* as a valid *token* for you lexer. Syntax errors, on the other side, will be thrown by your scanner when a given set of **already** recognised valid tokens don't match any of the right sides of your grammar rules. simple panic-mode error handling system requires that we return to a high-level parsing function when a parsing or lexical error is detected.

Error-recovery actions are:

i. Delete one character from the remaining input.
ii. Insert a missing character in to the remaining input.
iii. Replace a character by another character.
iv. Transpose two adjacent characters.

DIFFERENCE BETWEEN COMPILER AND INTERPRETER

A compiler converts the high level instruction into machine language while an interpreter converts the high level instruction into an intermediate form.

Before execution, entire program is executed by the compiler whereas after translating the first line, an interpreter then executes it and so on.

List of errors is created by the compiler after the compilation process while an interpreter stops translating after the first error.

An independent executable file is created by the compiler whereas interpreter is required by an interpreted program each time.

The compiler produce object code whereas interpreter does not produce object code. In the process of compilation the program is analyzed only once and then the code is generated whereas source program is interpreted every time it is to be executed and every time the source program is analyzed. hence interpreter is less efficient than compiler.

Examples of interpreter: A *UPS Debugger* is basically a graphical source level debugger but it contains built in C interpreter which can handle multiple source files.

Example of compiler: *Borland c compiler* or Turbo C compiler compiles the programs written in C or C++.

REGULAR EXPRESSIONS

Regular expression is a formula that describes a possible set of string.

Component of regular expression..

X	**the character x**
.	**any character, usually accept a new line**
[x y z]	**any of the characters x, y, z, …..**
R?	**a R or nothing (=optionally as R)**
R*	**zero or more occurrences…..**
R+	**one or more occurrences ……**
R1R2	**an R1 followed by an R2**
R2R1	**either an R1 or an R2.**

A token is either a single string or one of a collection of strings of a certain type. If we view the set of strings in each token class as an language, we can use the regular-expression notation to describe tokens.

Consider an identifier, which is defined to be a letter followed by zero or more letters or digits. In regular expression notation we would write.

Identifier = letter (letter | digit)*

Here are the rules that define the regular expression over alphabet .

- o is a regular expression denoting { € }, that is, the language containing only the empty string.
- o For each „a‟ in ∑, is a regular expression denoting { a }, the language with only one string consisting of the single symbol „a‟ .
- o If R and S are regular expressions, then

(R) | (S) means LrULs
R.S means Lr.Ls
R* denotes Lr*

REGULAR DEFINITIONS

For notational convenience, we may wish to give names to regular expressions and to define regular expressions using these names as if they were symbols.

Identifiers are the set or string of letters and digits beginning with a letter. The following regular definition provides a precise specification for this class of string.

Example-1,

Ab*|cd? Is equivalent to (a(b*)) | (c(d?))

Pascal identifier

Letter - A | B | ……| Z | a | b |……| z|

Digits - 0 | 1 | 2 | …. | 9

letter (letter / digit)* I

Recognition of tokens:

We learn how to express pattern using regular expressions. Now, we must study how to take the patterns for all the needed tokens and build a piece of code that examins the input string and finds a prefix that is a lexeme matching one of the patterns.

```
Stmt  ->  if expr then stmt
        | If expr then else stmt
        | є
Expr -->  term relop term
        |term

Term -->id
```

For relop ,we use the comparison operations of languages like Pascal or SQL where = is "equals" and < > is "not equals" because it presents an interesting structure of lexemes. The terminal of grammar, which are if, then , else, relop ,id and numbers are the names of tokens as far as the lexical analyzer is concerned, the patterns for the tokens are described using regular definitions.

```
digit     -->[0,9]
digits    -->digit+
number -->digit(.digit)?(e.[+-]?digits)?
letter    -->[A-Z,a-z]
id        -->letter(letter/digit)*
if        --> if
then      -->then
else      -->else
relop     --></>/<=/>=/==/< >
```

In addition, we assign the lexical analyzer the job stripping out white space, by recognizing the "token" we defined by:

ws --> (blank/tab/newline)$^{+}$

Here, blank, tab and newline are abstract symbols that we use to express the ASCII characters of the same names. Token ws is different from the other tokens in that ,when we recognize it, we do not return it to parser ,but rather restart the lexical analysis from the character that follows the white space . It is the following token that gets returned to the parser.

Lexeme	Token Name	Attribute Value
Any ws	_	_
if	if	_
then	then	_
else	else	_
Any Id	id	pointer to table entry
Any number	number	pointer to table entry
<	relop	LT

<=	relop	LE
=	relop	ET
< >	relop	NE

TRANSITION DIAGRAM:

Transition Diagram has a collection of nodes or circles, called states. Each state represents a condition that could occur during the process of scanning the input looking for a lexeme that matches one of several patterns .

Edges are directed from one state of the transition diagram to another. each edge is labeled by a symbol or set of symbols.

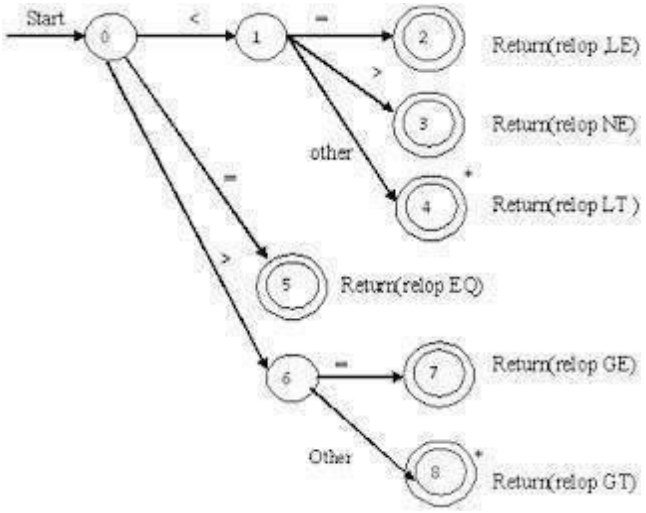
If we are in one state s, and the next input symbol is a, we look for an edge out of state s labeled by a. if we find such an edge ,we advance the forward pointer and enter the state of the transition diagram to which that edge leads.

Some important conventions about transition diagrams are

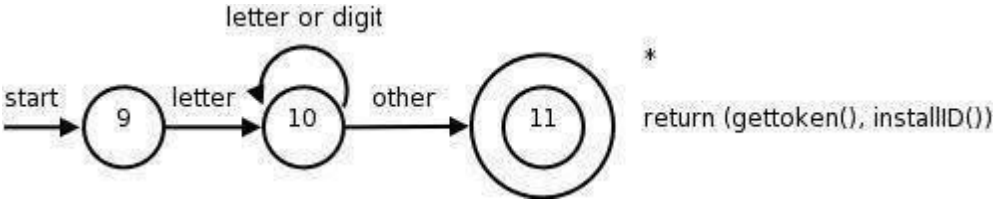
1. Certain states are said to be accepting or final .These states indicates that a lexeme has been found, although the actual lexeme may not consist of all positions b/w the lexeme Begin and forward pointers we always indicate an accepting state by a double circle.

2. In addition, if it is necessary to return the forward pointer one position, then we shall additionally place a * near that accepting state.

3. One state is designed the state ,or initial state ., it is indicated by an edge labeled "start" entering from nowhere .the transition diagram always begins in the state before any input symbols have been used.



As an intermediate step in the construction of a LA, we first produce a stylized flowchart, called a transition diagram. Position in a transition diagram, are drawn as circles and are called as states.



The above TD for an identifier, defined to be a letter followed by any no of letters or digits.A sequence of transition diagram can be converted into program to look for the tokens specified by the diagrams. Each state gets a segment of code.

If	=	if
Then	=	then
Else	=	else
Relop	=	< | <= | = | > | >=
Id	=	letter (letter | digit) *|
Num	=	digit |

2.10 AUTOMATA

An automation is defined as a system where information is transmitted and used for performing some functions without direct participation of man.

1, an automation in which the output depends only on the input is **called an automation without memory.**

2, an automation in which the output depends on the input and state also is **called as automation with memory**.

3, an automation in which the output depends only on the state of the machine is **called a Moore machine**.

3, an automation in which the output depends on the state and input at any instant of time is **called a mealy machine**.

DESCRIPTION OF AUTOMATA

1, an automata has a mechanism to read input from input tape,

2, any language is recognized by some automation, Hence these automation are basically language „acceptors" or „language recognizers".

Types of Finite Automata

- Deterministic Automata
- Non-Deterministic Automata.

DETERMINISTIC AUTOMATA

A deterministic finite automata has at most one transition from each state on any input. A DFA is a special case of a NFA in which:-

1, it has no transitions on input € ,

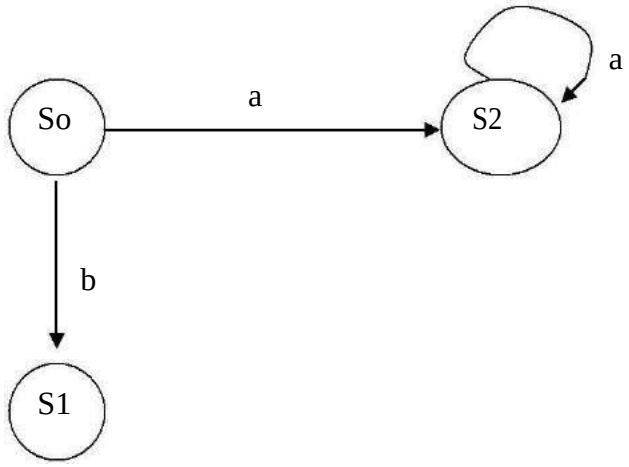
2, each input symbol has at most one transition from any state.

DFA formally defined by 5 tuple notation M = (Q, ∑, δ, qo, F), where
Q is a finite „set of states", which is non empty.
∑ is „input alphabets", indicates input set.
qo is an „initial state" and qo is in Q ie, qo, ∑, Q
F is a set of „Final states",
δ is a „transmission function" or mapping function, using this function the next state can be determined.

The regular expression is converted into minimized DFA by the following procedure:

Regular expression → NFA → DFA → Minimized DFA

The Finite Automata is called DFA if there is only one path for a specific input from current state to next state.



From state S0 for input „a" there is only one path going to S2. similarly from S0 there is only one path for input going to S1.

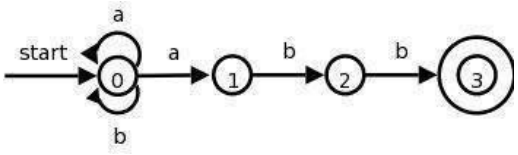
NONDETERMINISTIC AUTOMATA

- A NFA is a mathematical model that consists of
  - A set of states S.
  - A set of input symbols ∑.
  - A transition for move from one state to an other.
  - A state so that is distinguished as the start (or initial) state.
  - A set of states F distinguished as accepting (or final) state.
  - A number of transition to a single symbol.

- A NFA can be diagrammatically represented by a labeled directed graph, called a transition graph, In which the nodes are the states and the labeled edges represent the transition function.

- This graph looks like a transition diagram, but the same character can label two or more transitions out of one state and edges can be labeled by the special symbol € as well as by input symbols.

- The transition graph for an NFA that recognizes the language (a | b) * abb is shown



DEFINITION OF CFG

It involves four quantities.

CFG contain terminals, N-T, start symbol and production.

- Terminal are basic symbols form which string are formed.
- N-terminals are synthetic variables that denote sets of strings
- In a Grammar, one N-T are distinguished as the start symbol, and the set of string it denotes is the language defined by the grammar.
- The production of the grammar specify the manor in which the terminal and N-T can be combined to form strings.
- Each production consists of a N-T, followed by an arrow, followed by a string of one terminal and terminals.

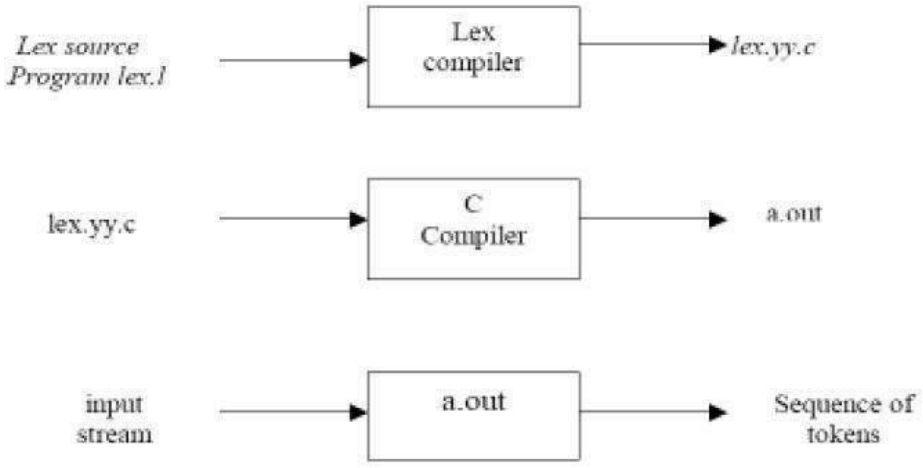
DEFINITION OF SYMBOL TABLE

- An extensible array of records.
- The identifier and the associated records contains collected information about the identifier.

  FUNCTION identify (Identifier name)

  RETURNING a pointer to identifier information contains
- The actual string
- A macro definition A
- keyword definition
- A list of type, variable & function definition
- A list of structure and union name definition
- A list of structure and union field selected definitions.

Creating a lexical analyzer with Lex



Lex specifications:

A Lex program (the .l file) consists of three parts:

declarations
%%
translation rules
%%
auxiliary procedures

1. The *declarations* section includes declarations of variables,manifest constants(A manifest constant is an identifier that is declared to represent a constant e.g. *# define PIE 3.14*), and regular definitions.
2. The *translation rules* of a Lex program are statements of the form :

   p1 *{action 1}*
   p2 *{action 2}*
   p3 *{action 3}*
   ... ...
   ... ...

   where each *p* is a regular expression and each *action* is a program fragment describing what action the lexical analyzer should take when a pattern *p* matches a lexeme. In Lex the actions are written in C.

3. The third section holds whatever *auxiliary procedures* are needed by the *actions.*Alternatively these procedures can be compiled separately and loaded with the lexical analyzer.

Note: You can refer to a sample lex program given in page no. 109 of chapter 3 of the book: *Compilers: Principles, Techniques, and Tools* by Aho, Sethi & Ullman for more clarity.

UNIT –II

Syntax Analysis-:The Role of a parser, Context free Grammars, Writing A grammar, top down parsing bottom up parsing, Introduction to Lr Parser.

UNIT -2

SYNTAX ANALYSIS

ROLE OF THE PARSER

Parser obtains a string of tokens from the lexical analyzer and verifies that it can be generated by the language for the source program. The parser should report any syntax errors in an intelligible fashion. The two types of parsers employed are:

1.Top down parser: which build parse trees from top(root) to bottom(leaves)

2.Bottom up parser: which build parse trees from leaves and work up the root.

Therefore there are two types of parsing methods– top-down parsing and bottom-up parsing

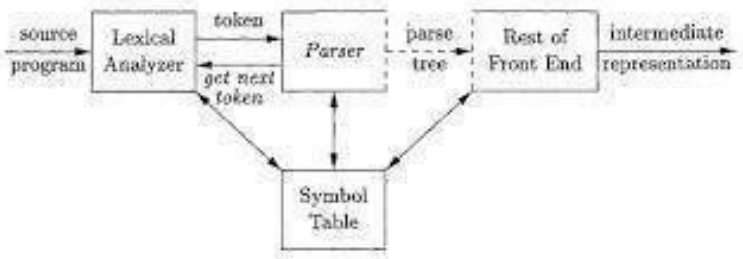


Figure 4.1: Position of parser in compiler model

Context free Grammars(CFG)

CFG is used to specify the syntax of a language. A grammar naturally describes the hierarchical structure of most program-ming language constructs.

Formal Definition of Grammars

A context-free grammar has four components:

1. A set of terminal symbols, sometimes referred to as "tokens." The terminals are the elementary symbols of the language defined by the grammar.
2. A set of nonterminals, sometimes called "syntactic variables." Each non-terminal represents a set of strings of terminals, in a manner we shall describe.
3. A set of productions, where each production consists of a nonterminal, called the head or left side of the production, an arrow, and a sequence of terminals and1or nonterminals, called the *body* or *right side* of the production. The intuitive intent of a production is to specify one of the written forms of a construct; if the head nonterminal represents a construct, then the body

represents a written form of the construct.

.

4. A designation of one of the nonterminals as the *start* symbol.

Production is for a nonterminal if the nonterminal is the head of the production. A string of terminals is a sequence of zero or more terminals. The string of zero terminals, written as ***E* ,** is called the empty string.

Derivations

A grammar derives strings by beginning with the start symbol and repeatedly replacing a nonterminal by the body of a production for that nonterminal. The terminal strings that can be derived from the start symbol form the *language* defined by the grammar.

Leftmost and Rightmost Derivation of a String

- **Leftmost derivation** − A leftmost derivation is obtained by applying production to the leftmost variable in each step.
- **Rightmost derivation** − A rightmost derivation is obtained by applying production to the rightmost variable in each step.
- **Example**

Let any set of production rules in a CFG be

$$X \rightarrow X+X \mid X*X \mid X \mid a$$

over an alphabet {a}.

The leftmost derivation for the string **"a+a*a"** is

$$X \rightarrow X+X \rightarrow a+X \rightarrow a + X*X \rightarrow a+a*X \rightarrow a+a*a$$

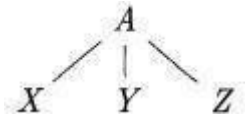
The rightmost derivation for the above string **"a+a*a"** is

$$X \rightarrow X*X \rightarrow X*a \rightarrow X+X*a \rightarrow X+a*a \rightarrow a+a*a$$

Derivation or Yield of a Tree

The derivation or the yield of a parse tree is the final string obtained by concatenating the labels of the leaves of the tree from left to right, ignoring the Nulls. However, if all the leaves are Null, derivation is Null.

parse tree pictorially shows how the start symbol of a grammar derives a string in the language. If nonterminal A has a production A → XYZ , then a parse tree may have an interior node labeled A with three children labeled X, Y, and Z, from left to right:



Given a context-free grammar, a *parse tree* according to the grammar is a tree with the following properties:

1. The root is labeled by the start symbol.
2. Each leaf is labeled by a terminal or by e.
3. Each interior node is labeled by a nonterminal

.

If A is the nonterminal labeling some interior node and X I , Xz, . . . ,Xn are the labels of the children of that node from left to right, then there must be a production A → X1X2 . . Xn . Here, X1, X2,. . . , Xn, each stand for a symbol that is either a terminal or a nonterminal. As a special case, if A → **c** is a production, then a node labeled A may have a single child labeled ***E***

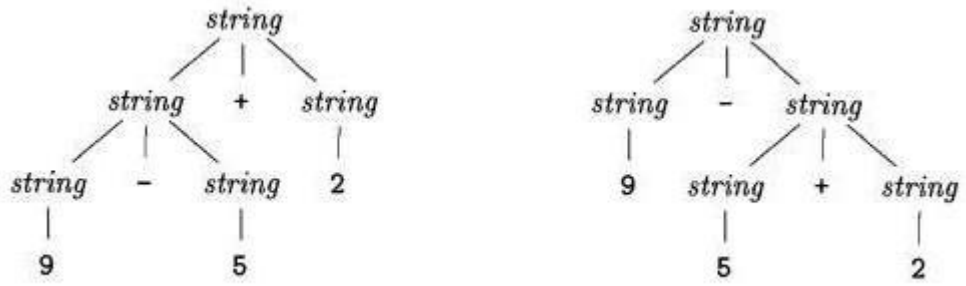
Ambiguity

A grammar can have more than one parse tree generating a given string of terminals. Such a grammar is said to be *ambiguous.* To show that a grammar is ambiguous, all we need to do is find a terminal string that is the yield of more than one parse tree. Since a string with more than one parse tree usually has more than one meaning, we need to design unambiguous grammars for compiling applications, or to use ambiguous grammars with additional rules to resolve the ambiguities.

Example :: Suppose we used a single nonterminal *string* and did not distinguish between digits and lists,

string → *string* + *string* | *string* − *string* | 0 | 1 | 2 | 3 | 4 | 5 | 6 | 7 | 8 | 9

Fig. shows that an expression like 9-5+2 has more than one parse tree with this grammar. The two trees for 9-5+2 correspond to the two ways of parenthesizing the expression: (9-5) +2 and 9- (5+2) . This second parenthesization gives the expression the unexpected value 2 rather than the customary value 6.



Two parse trees for 9-5+2

Verifying the language generated by a grammar

The set of all strings that can be derived from a grammar is said to be the language generated from that grammar. A language generated by a grammar **G** is a subset formally defined by

$$L(G)=\{W|W \in \sum^*, S \Rightarrow_G \mathbf{W}\}$$

If **L(G1) = L(G2)**, the Grammar **G1** is equivalent to the Grammar **G2**.

Example

If there is a grammar

G: N = {S, A, B} T = {a, b} P = {S → AB, A → a, B → b}

Here **S** produces **AB**, and we can replace **A** by **a**, and **B** by **b**. Here, the only accepted string is **ab**, i.e.,

L(G) = {ab}

.

Writing a grammar

A *grammar* consists of a number of *productions*. Each production has an abstract symbol called a *nonterminal* as its *left-hand side*, and a sequence of one or more nonterminal and *terminal* symbols as its *right-hand side*. For each grammar, the terminal symbols are drawn from a specified *alphabet*.

Starting from a sentence consisting of a single distinguished nonterminal, called the *goal symbol*, a given context-free grammar specifies a *language*, namely, the set of possible sequences of terminal symbols that can result from repeatedly replacing any nonterminal in the sequence with a right-hand side of a production for which the nonterminal is the left-hand side.

There are **four categories** in writing a grammar :

1. Lexical Vs Syntax Analysis
2. Eliminating ambiguous grammar.
3. Eliminating left-recursion
4. Left-factoring.

Each parsing method can handle grammars only of a certain form hence, the initial grammar may have to be rewritten to make it parsable

1. Lexical Vs Syntax Analysis

Reasons for using the regular expression to define the lexical syntax of a language

a) Regular expressions provide a more concise and easier to understand notation for tokens than grammars.
b) The lexical rules of a language are simple and to describe them, we donot need notation as powerful as grammars.
c) Efficient lexical analyzers can be constructed automatically from RE than from grammars.
d) Separating the syntactic structure of a language into lexical and nonlexical parts provides a convenient way of modularizing the front end into two manageable-sized components.

2. Eliminating ambiguous grammar.

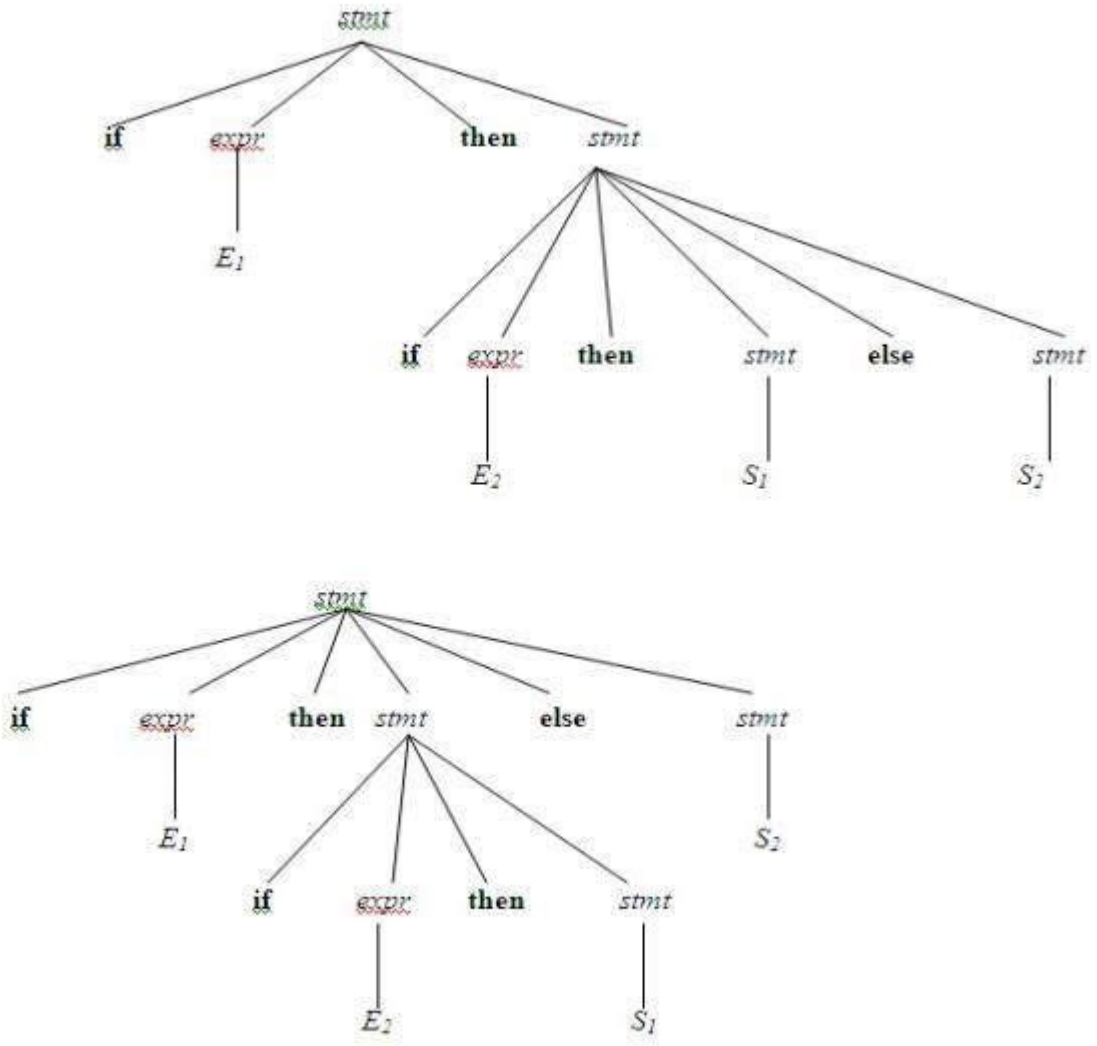
Ambiguity of the grammar that produces more than one parse tree for leftmost or rightmost derivation can be eliminated by re-writing the grammar.

Consider this example,

G: *stmt*→**if** *expr* **then** *stmt*
|**if** *expr* **then** *stmt* **else** *stmt*
|**other**

This grammar is ambiguous since the string **if E1 then if E2 then S1 else S2 has the** following two parse trees for leftmost derivation

.



Two parse trees for an ambiguous sentence

The general rule is "Match each else with the closest unmatched then.This disambiguating rule can be used directly in the grammar,

To eliminate ambiguity, the following grammar may be used:

stmt→matched | unmatchedstmt

matched→**if** *expr stmt* **then** *matched* **else** *matchedstmt* | **other**

unmatched→ **if** *expr* **then** *stmt* | **if** *expr* **then** *matched* **else** *unmatchedstmt*

3. Eliminating left-recursion

.

Because we try to generate a leftmost derivation by scanning the input from left to right, grammars of the form $A \rightarrow A\ x$ may cause endless recursion.Such grammars are called left-recursive and they must be transformed if we want to use a top-down parser.

- A grammar is left recursive if for a non-terminal A, there is a derivation $A \Rightarrow^{+} A\alpha$

- **To eliminate direct left recursion replace**

1) $A \rightarrow A\alpha \mid \beta$ **with** $A' \rightarrow \alpha A' \mid \varepsilon$

2) $A \rightarrow A\alpha_1 \mid A\alpha_2 \mid ... \mid A\alpha_m \mid \beta_1 \mid \beta_2 \mid ... \mid \beta_n$

with

$A \rightarrow \beta_1 B \mid \beta_2 B \mid ... \mid \beta_n B$

$B \rightarrow \alpha_1 B \mid \alpha_2 B \mid ... \mid \alpha_m B \mid \varepsilon$

4. Left-factoring

Left factoring is a grammar transformation that is useful for producing a grammar suitable for predictive parsing. When it is not clear which of two alternative productions to use to expand a non-terminal A, we can rewrite the A-productions to defer the decision until we have seen enough of the input to make the right choice.

- Consider $S \rightarrow$ if E then S else S | if E then S
  - Which of the two productions should we use to expand non-terminal S when the next token is if?

We can solve this problem by factoring out the common part in these rules.

$A \rightarrow \alpha\beta_1 \mid \alpha\beta_2 \mid ... \mid \alpha\beta_n \mid \gamma$

becomes

$A \rightarrow \alpha B \mid \gamma$

$B \rightarrow \beta_1 \mid \beta_2 \mid ... \mid \beta_n$

Consider the grammar , G : $S \rightarrow$ iEtS | iEtSeS | a

$E \rightarrow b$

Left factored, this grammar becomes

$S \rightarrow$ iEtSS' | a

$S' \rightarrow eS \mid \varepsilon$

$E \rightarrow b$

PARSING

It is the process of analyzing a continuous stream of input in order to determine its grammatical structure with respect to a given formal grammar.

Types of parsing:

1. Top down parsing
2. Bottom up parsing

Top-down parsing : A parser can start with the start symbol and try to transform it to the input string. **Example :** LL Parsers.

.

Bottom-up parsing : A parser can start with input and attempt to rewrite it into the start symbol.
Example : LR Parsers.

TOP-DOWN PARSING

It can be viewed as an attempt to find a left-most derivation for an input string or an attempt to construct a parse tree for the input starting from the root to the leaves.

Types of TOP-DOWN PARSING

1. Recursive descent parsing
2. Predictive parsing

RECURSIVE DESCENT PARSING

Recursive descent is a top-down parsing technique that constructs the parse tree from the top and the input is read from left to right. It uses procedures for every non-terminal entity. This parsing technique recursively parses the input to make a parse tree, which may or may not require back-tracking. But the grammar associated with it (if not left factored) cannot avoid back-tracking. A form of recursive-descent parsing that does not require any back-tracking is known as **predictive parsing**.

This parsing technique is regarded recursive as it uses context-free grammar which is recursive in nature.

This parsing method may involve backtracking.

Example for :backtracking

Consider the grammar G : S → cAd

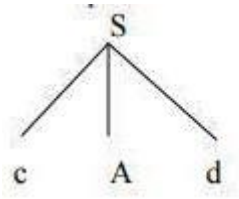
A→ab|a

and the input string w=cad.

The parse tree can be constructed using the following top-down approach :

Step1:

Initially create a tree with single node labeled S. An input pointer points to 'c', the first symbol of w. Expand the tree with the production of S.



Step2:

The leftmost leaf 'c' matches the first symbol of w, so advance the input pointer to the second symbol of w 'a' and consider the next leaf 'A'. Expand A using the first alternative.

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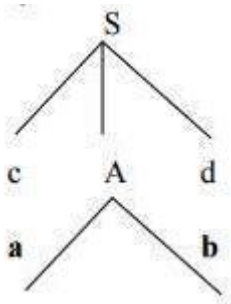
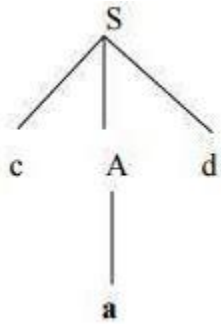


Step3:

The second symbol 'a' of w also matches with second leaf of tree. So advance the input pointer to third symbol of w 'd'**.** But the third leaf of tree is b which does not match with the input symbol **d.** Hence discard the chosen production and reset the pointer to second **backtracking.**

Step4: Now try the second alternative for A.



Now we can halt and announce the successful completion of parsing.

Predictive parsing

It is possible to build a nonrecursive predictive parser by maintaining a stack explicitly, rather than implicitly via recursive calls. The key problem during predictive parsing is that of determining the production to be applied for a nonterminal . The nonrecursive parser in figure looks up the production to be applied in parsing table. In what follows, we shall see how the table can be constructed directly from certain grammars.

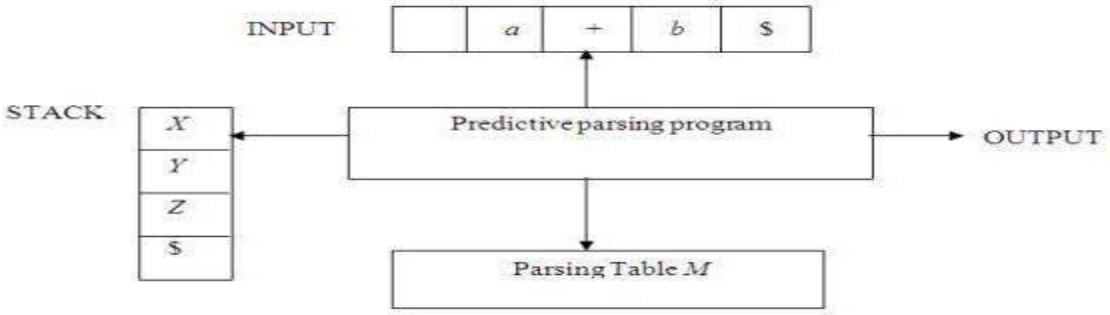


Fig. 2.4 Model of a nonrecursive predictive parser

A table-driven predictive parser has an input buffer, a stack, a parsing table, and an output stream. The input buffer contains the string to be parsed, followed by $, a symbol used as a right endmarker to indicate the end of the input string. The stack contains a sequence of grammar symbols

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with $ on the bottom, indicating the bottom of the stack. Initially, the stack contains the start symbol of the grammar on top of $. The parsing table is a two dimensional array M[A,a] where A is a nonterminal, and a is a terminal or the symbol $. The parser is controlled by a program that behaves as follows. The program considers X, the symbol on the top of the stack, and a, the current input symbol. These two symbols determine the action of the parser. There are three possibilities.

1 If X= a=$, the parser halts and announces successful completion of parsing.
2 If X=a!=$, the parser pops X off the stack and advances the input pointer to the next input symbol.
3 If X is a nonterminal, the program consults entry M[X,a] of the parsing table M. This entry will be either an X-production of the grammar or an error entry. If, for example, M[X,a]={X- >UVW}, the parser replaces X on top of the stack by WVU(with U on top). As output, we shall assume that the parser just prints the production used; any other code could be executed here. If M[X,a]=error, the parser calls an error recovery routine

Implementation of predictive parser:

1. Elimination of left recursion, left factoring and ambiguous grammar.
2. Construct FIRST() and FOLLOW() for all non-terminals.
3. Construct predictive parsing table.
4. Parse the given input string using stack and parsing table

Algorithm for Nonrecursive predictive parsing.

Input. A string w and a parsing table M for grammar G.
Output. If w is in L(G), a leftmost derivation of w; otherwise, an error indication.

Method. Initially, the parser is in a configuration in which it has $S on the stack with S, the start symbol of G on top, and w$ in the input buffer. The program that utilizes the predictive parsing table M to produce a parse for the input is shown in Fig.

```
set ip to point to the first symbol of w$. repeat

    let X be the top stack symbol and a the symbol pointed to by ip. if X is a terminal of $ then
      if X=a then
         pop X from the stack and advance ip else error()
      else
      if M[X,a]=X->Y1Y2...Yk then begin pop X from the stack;
         push Yk,Yk-1...Y1 onto the stack, with Y1 on top; output the production X-> Y1Y2...Yk
      end
      else error()
  until X=$ /* stack is empty */
```

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FIRST AND FOLLOW

The construction of a predictive parsing table is aided by two functions associated with a grammar:

1. FIRST
2. FOLLOW

To compute FIRST(X) for all grammar symbols X, apply the following rules until no more terminals or e can be added to any FIRST set.

Rules for FIRST **( ):**

1. If X is terminal, then FIRST(X) is {X}.
2. If X → ε is a production, then add ε to FIRST(X).
3. If X is non-terminal and X → aα is a production then add a to FIRST(X).
4. If X is non-terminal and X → Y1 Y2…Yk is a production, then place a in FIRST(X) if for some i, a is in FIRST(Yi), and ε is in all of FIRST(Y1),…,FIRST(Yi-1);that is, Y1,….Yi-1=> ε. If ε is in FIRST(Yj) for all j=1,2,..,k, then add ε to FIRST(X).

Rules for FOLLOW ( ):

1. If S is a start symbol, then FOLLOW(S) contains $.
2. If there is a production A → αBβ, then everything in FIRST(β) except ε is placed in follow(B).
3. If there is a production A → αB, or a production A → αBβ where FIRST(β) contains ε, then everything in FOLLOW(A) is in FOLLOW(B).

Algorithm for construction of predictive parsing table:

Input : Grammar G
Output : Parsing table M
Method :

1. For each production A → α of the grammar, do steps 2 and 3.
2. For each terminal a in FIRST(α), add A → α to M[A, a].
3. If ε is in FIRST(α), add A → α to M[A, b] for each terminal b in FOLLOW(A). If ε is in FIRST(α) and $ is in FOLLOW(A) , add A → α to M[A, $].
4. Make each undefined entry of M be error.

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Example:

Consider the following grammar :

E→E+T|T

T→T*F|F

F→(E)|id

After eliminating left-recursion the grammar is

E →TE'

E' → +TE' | ε

T →FT'

T' → *FT' | ε

F → (E)|id

First( ) :

FIRST(E) = { (, id}

FIRST(E') ={+ , ε }

FIRST(T) = { (, id}

FIRST(T') = {*, ε }

FIRST(F) = { (, id }

Follow( ):

FOLLOW(E) = { $,) }

FOLLOW(E') = { $,) }

FOLLOW(T) = { +, $,) }

FOLLOW(T') = { +, $,) }

FOLLOW(F) = {+, * , $,) }

Predictive parsing Table

NON-TERMINAL	id	+	*	(	)	$
E	E→TE'			E→TE'		
E'		E'→+TE'			E'→ε	E'→ε
T	T→FT'			T→FT'		
T'		T'→ε	T'→*FT'		T'→ε	T'→ε
F	F→id			F→(E)		

.

Stack Implementation

stack	Input	Output
$E	id+id*id $	
$E'T	id+id*id $	E→TE'
$E'T'F	id+id*id $	T→FT'
$E'T'id	id+id*id $	F→id
$E'T'	+id*id $	
$E'	+id*id $	T' → ε
$E'T+	+id*id $	E' → +TE'
$E'T	id*id $	
$E'T'F	id*id $	T→FT'
$E'T'id	id*id $	F→id
$E'T'	*id $	
$E'T'F*	*id $	T' → *FT'
$E'T'F	id $	
$E'T'id	id $	F→id
$E'T'	$	
$E'	$	T' → ε
$	$	E' → ε

.

LL(1) GRAMMAR

The parsing table algorithm can be applied to any grammar G to produce a parsing table M. For some Grammars, for example if G is left recursive or ambiguous, then M will have at least one multiply-defined entry. A grammar whose parsing table has no multiply defined entries is said to be LL(1). It can be shown that the above algorithm can be used to produce for every LL(1) grammar G, a parsing table M that parses all and only the sentences of G. LL(1) grammars have several distinctive properties. No ambiguous or left recursive grammar can be LL(1). There remains a question of what should be done in case of multiply defined entries. One easy solution is to eliminate all left recursion and left factoring, hoping to produce a grammar which will produce no multiply defined entries in the parse tables. Unfortunately there are some grammars which will give an LL(1) grammar after any kind of alteration. In general, there are no universal rules to convert multiply defined entries into single valued entries without affecting the language recognized by the parser.

The main difficulty in using predictive parsing is in writing a grammar for the source language such that a predictive parser can be constructed from the grammar. Although left recursion elimination and left factoring are easy to do, they make the resulting grammar hard to read and difficult to use the translation purposes. To alleviate some of this difficulty, a common organization for a parser in a compiler is to use a predictive parser for control constructs and to use operator precedence for expressions.however, if an lr parser generator is available, one can get all the benefits of predictive parsing and operator precedence automatically.

ERROR RECOVERY IN PREDICTIVE PARSING

The stack of a nonrecursive predictive parser makes explicit the terminals and nonterminals that the parser hopes to match with the remainder of the input. We shall therefore refer to symbols on the parser stack in the following discussion. An error is detected during predictive parsing when the terminal on top of the stack does not match the next input symbol or when nonterminal A is on top of the stack, a is the next input symbol, and the parsing table entry M[A,a] is empty.

Consider error recovery predictive parsing using the following two methods

Panic-Mode recovery
Phrase Level recovery

Panic-mode error recovery is based on the idea of skipping symbols on the input until a token in a selected set of synchronizing tokens appears. Its effectiveness depends on the choice of synchronizing set. The sets should be chosen so that the parser recovers quickly from errors that are likely to occur in practice.

As a starting point, we can place all symbols in FOLLOW(A) into the synchronizing set for nonterminal A. If we skip tokens until an element of FOLLOW(A) is seen and pop A from the stack, it is likely that parsing can continue.

It is not enough to use FOLLOW(A) as the synchronizingset for A. Fo example , if semicolons terminate statements, as in C, then keywords that begin statements may not appear in the FOLLOW set of the nonterminal generating expressions. A missing semicolon after an assignment may therefore result in the keyword beginning the next statement being skipped. Often, there is a hierarchica structure on constructs in a language; e.g., expressions appear within statement, which appear within bblocks,and so on. We can add to the

.

synchronizing set of a lower construct the symbols that begin higher constructs. For example, we might add keywords that begin statements to the synchronizing sets for the nonterminals generaitn expressions.

If we add symbols in FIRST(A) to the synchronizing set for nonterminal A, then it may be possible to resume parsing according to A if a symbol in FIRST(A) appears in the input.

If a nonterminal can generate the empty string, then the production deriving e can be used as a default. Doing so may postpone some error detection, but cannot cause an error to be missed. This approach reduces the number of nonterminals that have to be considered during error recovery.

If a terminal on top of the stack cannot be matched, a simple idea is to pop the terminal, issue a message saying that the terminal was inserted, and continue parsing. In effect, this approach takes the synchronizing set of a token to consist of all other tokens.

Phrase Level recovery

This involves, defining the blank entries in the table with pointers to some error routines which may

Change, delete or insert symbols in the input or
May also pop symbols from the stack

BOTTOM-UP PARSING

Constructing a parse tree for an input string beginning at the leaves and going towards the root is called bottom-up parsing. A general type of bottom-up parser is a shift-reduce parser.

2.6.1 SHIFT-REDUCE PARSING

Shift-reduce parsing is a type of bottom-up parsing that attempts to construct a parse tree for an input string beginning at the leaves (the bottom) and working up towards the root (the top).

Example:

Consider the grammar: $S \rightarrow aABe$
$A \rightarrow Abc \mid b$
$B \rightarrow d$

The sentence to be recognized is abbcde.

REDUCTION (LEFTMOST)	**RIGHTMOST DERIVATION**
abbcde ($A \rightarrow b$)	$S \rightarrow aABe$
aAbcde($A \rightarrow Abc$)	$\rightarrow aAde$
aAde ($B \rightarrow d$)	$\rightarrow aAbcde$
aABe ($S \rightarrow aABe$)	$\rightarrow abbcde$
S	

The reductions trace out the right-most derivation in reverse.

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Handles: A handle of a string is a substring that matches the right side of a production, and whose reduction to the non-terminal on the left side of the production represents one step along the reverse of a rightmost derivation.

Example:

Consider the grammar:

E→E+E
E→E*E
E→(E)
E→id

And the input string id1+id2*id3

The rightmost derivation is :

E→E+E
→ E+E*E
→ E+E*id3
→ E+id2*id3
→ id1+id2*

In the above derivation the underlined substrings are called handles.

Handle pruning:

A rightmost derivation in reverse can be obtained by "handle pruning". (i.e.) if w is a sentence or string of the grammar at hand, then w = γn, where γn is the nth right sentential form of some rightmost derivation.

Actions in shift-reduce parser:

- shift - The next input symbol is shifted onto the top of the stack.
- reduce - The parser replaces the handle within a stack with a non-terminal.
- accept - The parser announces successful completion of parsing.
- error - The parser discovers that a syntax error has occurred and calls an error recovery routine.

Conflicts in shift-reduce parsing:

There are two conflicts that occur in shift-reduce parsing:

1. Shift-reduce conflict: The parser cannot decide whether to shift or to reduce.
2. Reduce-reduce conflict: The parser cannot decide which of several reductions to make.

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Stack implementation of shift-reduce parsing :

Stack	Input	Action
$	$id_1+id_2*id_3$ $	shift
$ id_1	$+id_2*id_3$ $	reduce by E→id
$E	$+id_2*id_3$ $	shift
$E+	id_2*id_3 $	shift
$E+$id_2$	$*id_3$ $	reduce by E→id
$E+E	$*id_3$ $	shift
$E+E*	id3 $	shift
$E+E*id3	$	reduce by E→id
$E+E*E	$	reduce by E→ E *E
$E+E	$	reduce by E→ E+E
$E	$	accept

1. Shift-reduce conflict:

Example:

Consider the grammar:

E→E+E | E*E | id and input id+id*id

Stack	Input	Action	Stack	Input	
$E+E	*id $	Reduce by E→E+E	$E+E	*id $	Shift
$E	*id $	Shift	$E+E*	id $	Shift
$E*	id $	Shift	$E+E*id	$	Reduce by E→id
$E*id	$	Reduce by E→id	$E+E*E	$	Reduce by E→E*E
$E*E	$	Reduce by E→E*E	$E+E	$	Reduce by E→E*E
$E			$E		

.

2. Reduce-reduce conflict:

Consider the grammar: M→R+R|R+c|R

R→c

input c+c

Stack	Input	Action	Stack	Input	Action
$	c+c $	Shift	$	c+c $	Shift
$c	+c $	Reduce by R→c	$c	+c $	Reduce by R→c
$R	+c $	Shift	$R	+c $	Shift
$R+	c $	Shift	$R+	c $	Shift
$R+c	$	Reduce by R→c	$R+c	$	Reduce by M→R+c
$R+R	$	Reduce by M→R+R	$M	$	
$M	$				

INTRODUCTION TO LR PARSERS

An efficient bottom-up syntax analysis technique that can be used CFG is called LR(k) parsing. The 'L' is for left-to-right scanning of the input, the 'R' for constructing a rightmost derivation in reverse, and the 'k' for the number of input symbols. When 'k' is omitted, it is assumed to be 1.

Advantages of LR parsing:

1. It recognizes virtually all programming language constructs for which CFG can be written.
2. It is an efficient non-backtracking shift-reduce parsing method.
3. A grammar that can be parsed using LR method is a proper superset of a grammar that can be parsed with predictive parser
4. It detects a syntactic error as soon as possible.

Drawbacks of LR method:

It is too much of work to construct a LR parser by hand for a programming language grammar. A specialized tool, called a LR parser generator, is needed. Example: YACC.

Types of LR parsing method:

1. SLR- Simple LR

   Easiest to implement, least powerful.

2. CLR- Canonical LR

   Most powerful, most expensive.

3. LALR- Look-Ahead LR

   Intermediate in size and cost between the other two methods.

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The LR parsing algorithm:

The schematic form of an LR parser is as follows:

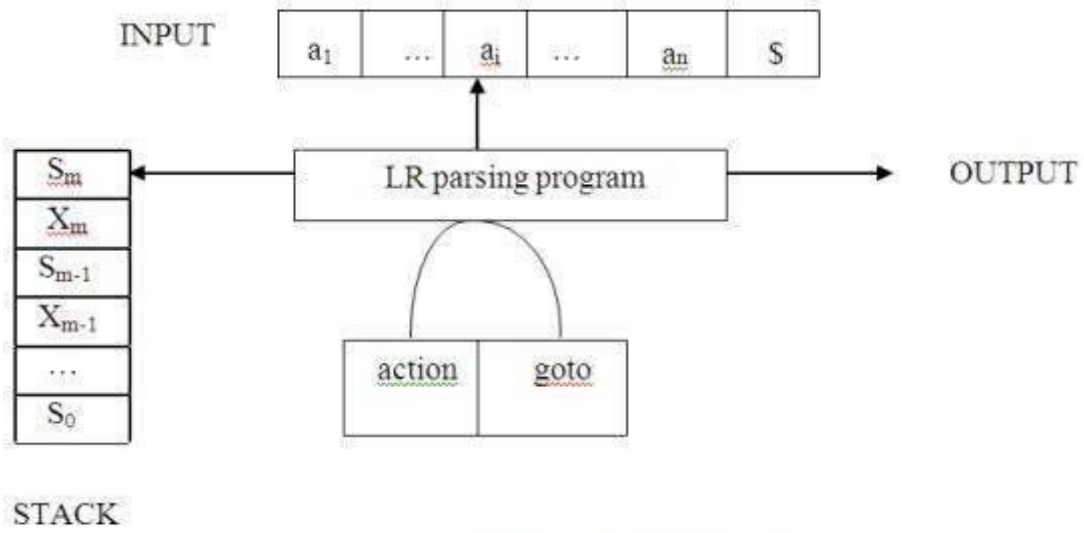

Fig. 2.5 Model of an LR parser

It consists of an input, an output, a stack, a driver program, and a pa parts (action and goto).

a. The driver program is the same for all LR parser.
b. The parsing program reads characters from an input buffer one at a time.
c. The program uses a stack to store a string of the form s0X1s1X2s2…Xmsm, where sm is on top. Each Xi is a grammar symbol and each si is a state.
d. The parsing table consists of two parts : action and goto functions.

Action : The parsing program determines sm, the state currently on top of stack, and ai, the current input symbol. It then consults action[sm,ai] in the action table which can have one of four values:

1. shift s, where s is a state,
2. reduce by a grammar production A → β,
3. accept,
4. Error.

Goto : The function goto takes a state and grammar symbol as arguments and produces a state.

LR Parsing algorithm:

Input: An input string w and an LR parsing table with functions action and goto for grammar G.
Output: If w is in L(G), a bottom-up-parse for w; otherwise, an error indication.
Method: Initially, the parser has s0 on its stack, where s0 is the initial state, and w$ in the input buffer.

The parser then executes the following program:

.

```
set ip to point to the first input symbol of w$; repeat forever begin
        let s be the state on top of the stack and a the symbol pointed to by ip;
if action[s, a] = shift s' then begin
        push a then s' on top of the stack; advance ip to the next input symbol end
        else if action[s, a] = reduce A→β then begin pop 2* | β | symbols off the stack;
        let s' be the state now on top of the stack; push A then goto[s', A] on top of the stack; output the
        production A→ β
        end
        else if action[s, a] = accept then
        return
        else error( )
        end
```

CONSTRUCTING SLR(1) PARSING TABLE

To perform SLR parsing, take grammar as input and do the following:

1. Find LR(0) items.
2. Completing the closure.
3. Compute goto(I,X), where, I is set of items and X is grammar symbol.

LR(0) items:

An LR(0) item of a grammar G is a production of G with a dot at some position of the right side. For example, production A → XYZ yields the four items :

A→.XYZ
A → X . YZ
A → XY . Z
A → XYZ .

Closure operation:

If I is a set of items for a grammar G, then closure(I) is the set of items constructed from I by the two rules:

1. Initially, every item in I is added to closure(I).
2. If A → α . Bβ is in closure(I) and B → γ is a production, then add the item B → . γ to I , if it is not already there. We apply this rule until no more new items can be added to closure(I).

Goto operation:

Goto(I, X) is defined to be the closure of the set of all items [A→ αX . β] such that [A→ α . Xβ] is in I.

Steps to construct SLR parsing table for grammar G are:

1. Augment G and produce G'
2. Construct the canonical collection of set of items C for G'
3. Construct the parsing action function action and goto using the following algorithm that requires FOLLOW(A) for each non-terminal of grammar.

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Algorithm for construction of SLR parsing table:

Input : An augmented grammar G'

Output : The SLR parsing table functions action and goto for G'

Method :

1. Construct C = {I0, I1, …. In}, the collection of sets of LR(0) items for G'.
2. State i is constructed from Ii.. The parsing functions for state i are determined as follows:
   (a) If [A→α·aβ] is in Ii and goto(Ii,a) = Ij, then set action[i,a] to "shift j". Here a must be terminal.
   (b) If [A→α·] is in Ii , then set action[i,a] to "reduce A→α" for all a in FOLLOW(A).
   (c) If [S'→S.] is in Ii, then set action[i,$] to "accept".
   If any conflicting actions are generated by the above rules, we say grammar is not SLR(1).
3. The goto transitions for state i are constructed for all non-term
   If goto(Ii,A) = Ij, then goto[i,A] = j.
4. All entries not defined by rules (2) and (3) are made "error"
5. The initial state of the parser is the one constructed from the [S'→.S].

Example on SLR (1) Grammar

S → E
E → E + T | T
T → T * F | F
F → id

Add Augment Production and insert '•' symbol at the first position for every production in G

S` → •E
E → •E + T
E → •T
T → •T * F
T → •F
F → •id

I0 State:

Add Augment production to the I0 State and Compute the Closure

I0 = Closure (S` → •E)

Add all productions starting with E in to I0 State because "." is followed by the non-terminal. So, the I0 State becomes

I0 = S` → •E
E → •E + T
E → •T

Add all productions starting with T and F in modified I0 State because "." is followed by the non-terminal. So, the I0 State becomes.

I0= S` → •E
E → •E + T
E → •T
T → •T * F
T → •F
F → •id

I1= Go to (I0, E) = closure (S` → E•, E → E• + T)
I2= Go to (I0, T) = closure (E → T•T, T• → * F)
I3= Go to (I0, F) = Closure (T → F•) = T → F•
I4= Go to (I0, id) = closure (F → id•) = F → id•
I5= Go to (I1, +) = Closure (E → E +•T)

Add all productions starting with T and F in I5 State because "." is followed by the non-terminal. So, the I5 State becomes

I5 = E → E +•T
T → •T * F
T → •F
F → •id

Go to (I5, F) = Closure (T → F•) = (same as I3)
Go to (I5, id) = Closure (F → id•) = (same as I4)

I6= Go to (I2, *) = Closure (T → T * •F)

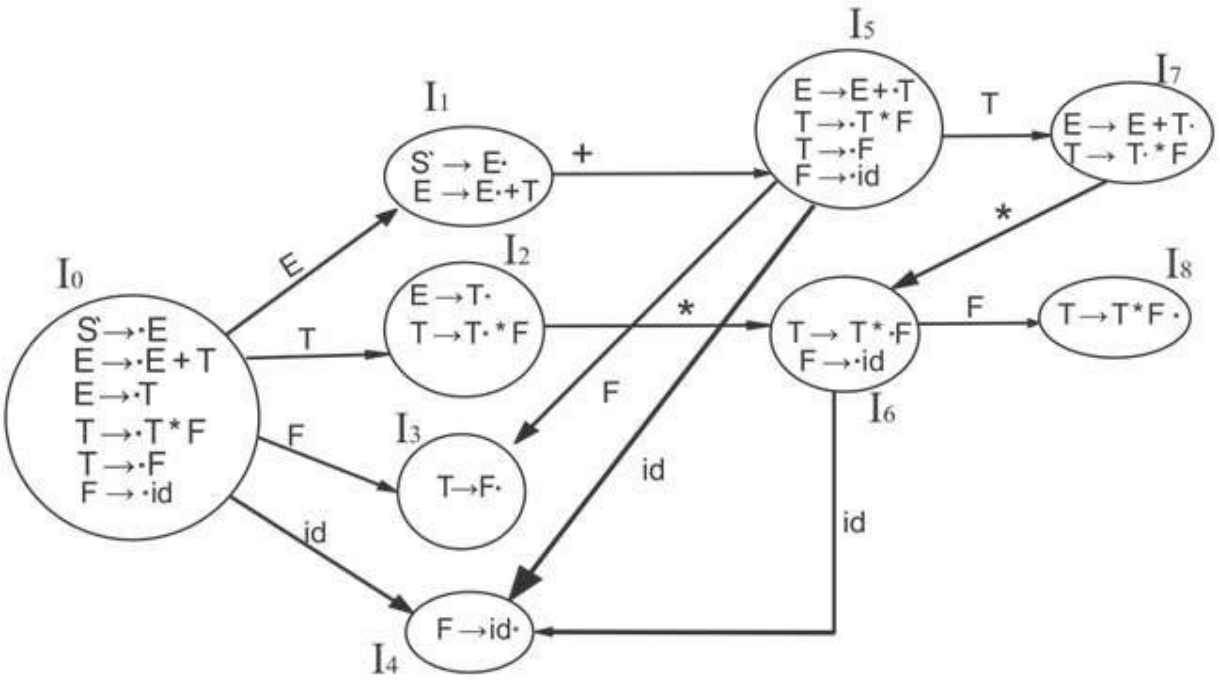
Add all productions starting with F in I6 State because "." is followed by the non-terminal. So, the I6 State becomes

I6 = T → T * •F
F → •id

Go to (I6, id) = Closure (F → id•) = (same as I4)

I7= Go to (I5, T) = Closure (E → E + T•) = E → E + T•
I8= Go to (I6, F) = Closure (T → T * F•) = T → T * F•

Drawing DFA



SLR (1) Table

States	Action				Go to		
	id	+	*	**$**	**E**	**T**	**F**
I_0	S_4				1	2	3
I_1		S_5		Accept			
I_2		R_2	S_6	R2			
I_3		R_4	R_4	R4			
I_4		R_5	R5	R5			
I_5	S4					7	3
I_6	S4						8
I_7		R1	S6	R1			
I_8		R3	R3	R3			

Explanation:

First (E) = First (E + T) U First (T)
First (T) = First (T * F) U First (F)
First (F) = {id}
First (T) = {id}
First (E) = {id}
Follow (E) = First (+T) U {$} = {+, $}
Follow (T) = First (*F) U First (F)
= {*, +, $}
Follow (F) = {*, +, $}

1) I1 contains the final item which drives S → E• and follow (S) = {$}, so action {I1, $} = Accept
2) I2 contains the final item which drives E → T• and follow (E) = {+, $}, so action {I2, +} = R2, action {I2, $} = R2
3) I3 contains the final item which drives T → F• and follow (T) = {+, *, $}, so action {I3, +} = R4, action {I3, *} = R4, action {I3, $} = R4
4) I4 contains the final item which drives F → id• and follow (F) = {+, *, $}, so action {I4, +} = R5, action {I4, *} = R5, action {I4, $} = R5
5) I7 contains the final item which drives E → E + T• and follow (E) = {+, $}, so action {I7, +} = R1, action {I7, $} = R1
6) I8 contains the final item which drives T → T * F• and follow (T) = {+, *, $}, so action {I8, +} = R3, action {I8, *} = R3, action {I8, $} = R3.

UNIT –III

More Powerful LR parser (LR1,LALR) Using Armigers Grammars Equal Recovery in Lr parser Syntax Directed Transactions Definition, Evolution order of SDTS Application of SDTS. Syntax Directed Translation Schemes.

UNIT -3

CANONICAL LR PARSING

CLR refers to canonical lookahead. CLR parsing use the canonical collection of LR (1) items to build the CLR (1) parsing table. CLR (1) parsing table produces the more number of states as compare to the SLR (1) parsing.

In the CLR (1), we place the reduce node only in the lookahead symbols.

Various steps involved in the CLR (1) Parsing:

1) For the given input string write a context free grammar
2) Check the ambiguity of the grammar
3) Add Augment production in the given grammar
4) Create Canonical collection of LR (0) items
5) Draw a data flow diagram (DFA)
6) Construct a CLR (1) parsing table

In the SLR method we were working with LR(0)) items. In CLR parsing we will be using LR(1) items. LR(k) item is defined to be an item using lookaheads of length k. So ,the LR(1) item is comprised of two parts : the LR(0) item and the lookahead associated with the item. The look ahead is used to determine that where we place the final item. The look ahead always add $ symbol for the argument production.

LR(1) parsers are more powerful parser.
for LR(1) items we modify the Closure and GOTO function.

Closure Operation

Closure(I)

repeat

for (each item [A -> ?.B?, a] in I)

for (each production B -> ? in G')

for (each terminal b in FIRST(?a))

add [B -> .? , b] to set I;

until no more items are added to I;

return I;

Goto Operation

```
Goto(I, X)
Initialise J to be the empty set;
for ( each item A -> ?.X?, a ] in I )
   Add item A -> ?X.?, a ] to se J;  /* move the dot one step */
return Closure(J);    /* apply closure to the set */
```

LR(1) items

```
Void items(G')
Initialise C to { closure ({[S' -> .S, $]})};
Repeat
   For (each set of items I in C)
      For (each grammar symbol X)
         if( GOTO(I, X) is not empty and not in C)
            Add GOTO(I, X) to C;
Until no new set of items are added to C;
```

ALGORITHM FOR CONSTRUCTION OF THE CANONICAL LR PARSING TABLE

Input: grammar G'
Output: canonical LR parsing table functions action and goto

1. Construct C = {I0, I1 , ..., In} the collection of sets of LR(1) items for G'.State i is constructed from Ii.
2. if [*A* -> a.ab, b>] is in Ii and goto(Ii, a) = Ij, then set action[i, a] to "shift j". Here a must be a terminal.
3. if [*A* -> a., a] is in Ii, then set action[i, a] to "reduce *A* -> a" for all a in FOLLOW(*A*). Here *A* may *not* be *S*'.
4. if [*S*' -> *S*.] is in Ii, then set action[i, $] to "accept"
5. If any conflicting actions are generated by these rules, the grammar is not LR(1) and the algorithm fails to produce a parser.
6. The goto transitions for state i are constructed for all *nonterminal*s *A* using the rule: If goto(Ii, A)= Ij, then goto[i, *A*] = j.
7. All entries not defined by rules 2 and 3 are made "error".
8. The inital state of the parser is the one constructed from the set of items containing [*S*' -> .S, $].

Example,

Consider the following grammar,

S‟->S
S->CC
C->cC
C->d

Sets of LR(1) items

I0: S‟->.S,$
S->.CC,$
C->.Cc,c/d
C->.d,c/d

I1: S‟->S.,$

I2: S->C.C,$
C->.Cc,$
C->.d,$

I3: C->c.C,c/d C->.Cc,c/dC->.d,c/d

I4: C->d.,c/d

I5: S->CC.,$

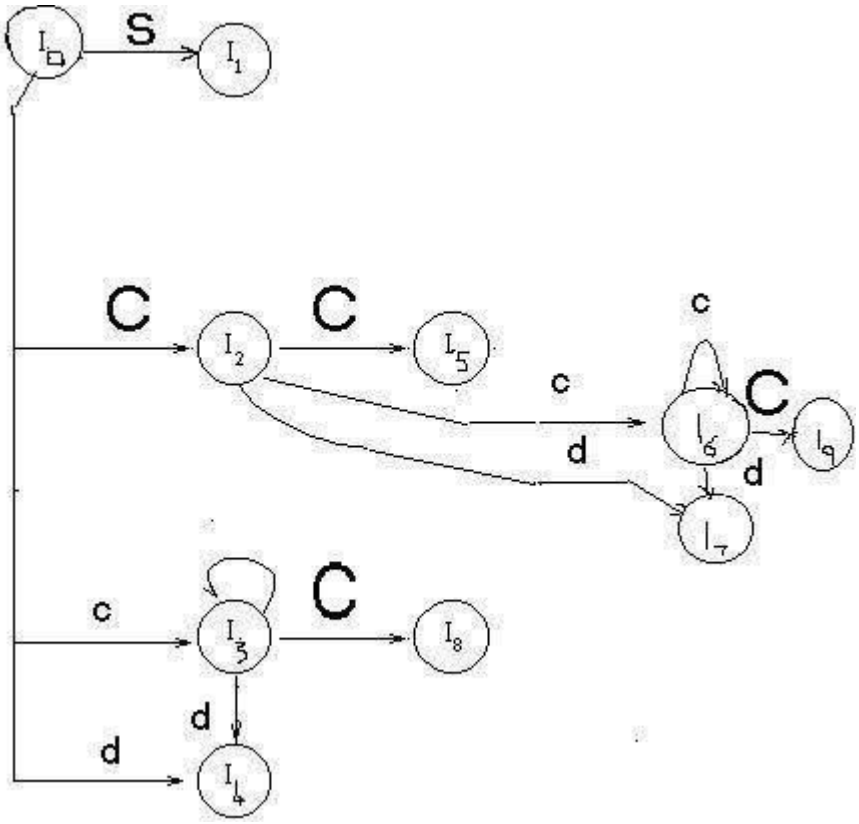
I6: C->c.C,$
C->.cC,$
C->.d,$

I7: C->d.,$

I8: C->cC.,c/d

I9: C->cC.,$

Here is what the corresponding DFA looks like



Parsing Table:state	c	d	$	S	C
0	S3	S4		1	2
1			acc		
2	S6	S7			5
3	S3	S4			8
4	R3	R3			
5			R1		
6	S6	S7			9
7			R3		
8	R2	R2			
9			R2		

3.3.LALR PARSER:

We begin with two observations. First, some of the states generated for LR(1) parsing have the same set of core (or first) components and differ only in their second component, the lookahead symbol. Our intuition is that we should be able to merge these states and reduce the number of states we have, getting close to the number of states that would be generated for LR(0) parsing. This observation suggests a hybrid approach: We can construct the canonical LR(1) sets of items and then look for sets of items having the same core. We merge these sets with common cores into one set of items. The merging of states with common cores can never produce a shift/reduce conflict that was not present in one of the original states because shift actions depend only on the core, not the lookahead. But it is possible for the merger to produce a reduce/reduce conflict.

Our second observation is that we are really only interested in the lookahead symbol in places where there is a problem. So our next thought is to take the LR(0) set of items and add lookaheads only where they are needed. This leads to a more efficient, but much more complicated method.

ALGORITHM FOR EASY CONSTRUCTION OF AN LALR TABLE

Input: G'
Output: LALR parsing table functions with action and goto for G'.

Method:

1. Construct C = {I0, I1 , ..., In} the collection of sets of LR(1) items for G'.
2. For each core present among the set of LR(1) items, find all sets having that core and replace these sets by the union.
3. Let C' = {J0, J1 , ..., Jm} be the resulting sets of LR(1) items. The parsing actions for state i are constructed from Ji in the same manner as in the construction of the canonical LR parsing table.
4. If there is a conflict, the grammar is not LALR(1) and the algorithm fails.
5. The goto table is constructed as follows: If J is the union of one or more sets of LR(1) items, that is, J = I0U I1 U ... U Ik, then the cores of goto(I0, X), goto(I1, X), ..., goto(Ik, X) are the same, since I0, I1 , ..., Ik all have the same core. Let K be the union of all sets of items having the same core asgoto(I1, X).

6. Then goto(J, X) = K.

Consider the above example,

I3 & I6 can be replaced by their union I36:C->c.C,c/d/$

C->.Cc,C/D/$

C->.d,c/d/$

I47:C->d.,c/d/$

I89:C->Cc.,c/d/$

Parsing Table

state	c	d	$	S	C
0	S36	S47		1	2
1			Accept		
2	S36	S47			5
36	S36	S47			89
47	R3	R3			
5			R1		
89	R2	R2	R2		

HANDLING ERRORS

The LALR parser may continue to do reductions after the LR parser would have spotted an error, but the LALR parser will never do a shift after the point the LR parser would have discovered the error and will eventually find the error.

DANGLING ELSE

The dangling else is a problem in computer programming in which an optional else clause in an If–then(–else) statement results in nested conditionals being ambiguous. Formally, the context-free grammar of the language is ambiguous, meaning there is more than one correct parse tree.

In many programming languages one may write conditionally executed code in two forms: the if-then form, and the if-then-else form – the else clause is optional:

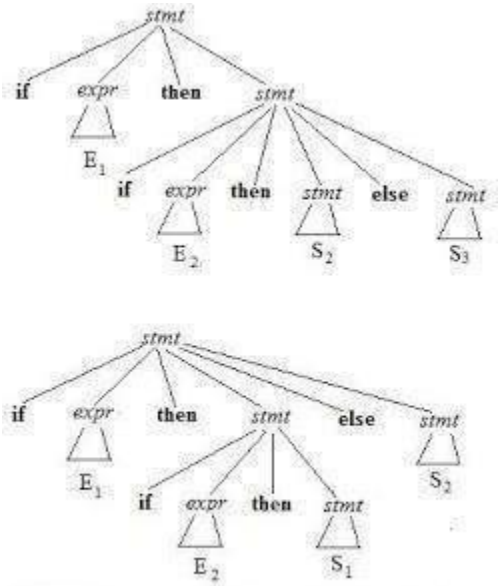


Fig 2.4 Two parse trees for an ambiguous sentence

Consider the grammar:

S ::= E $

E ::= E + E

| E * E

| (E)

| id

| num

and four of its LALR(1) states:

I0: S ::= . E $?

E ::= . E + E +*$

E ::= . E * E +*$

E ::= . (E) +*$

E ::= . id +*$

E ::= . num +*$

I1: S ::= E . $?

E ::= E . + E +*$

E ::= E . * E +*$

I2: E ::= E * . E +*$

E ::= . E + E +*$

E ::= . E * E +*$

E ::= . (E) +*$

E ::= . id +*$

E ::= . num +*$

I3: E ::= E * E . +*$

E ::= E . + E +*$

E ::= E . * E +*$

-

E ::= E . * E +*$

Here we have a shift-reduce error. Consider the first two items in I3. If we have a*b+c and we parsed a*b, do we reduce using E ::= E * E or do we shift more symbols? In the former case we get a parse tree (a*b)+c; in the latter case we get a*(b+c). To resolve this conflict, we can specify that * has higher precedence than +. The precedence of a grammar production is equal to the precedence of the rightmost token at the rhs of the production. For example, the precedence of the production E ::= E * E is equal to the precedence of the operator *, the precedence of the production E ::= (E) is equal to the precedence of the token), and the precedence of the production E ::= if E then E else E is equal to the precedence of the token else. The idea is that if the look ahead has higher precedence than the production currently used, we shift. For example, if we are parsing E + E using the production rule E ::= E + E and the look ahead is *, we shift *. If the look ahead has the same precedence as that of the current production and is left associative, we reduce, otherwise we shift. The above grammar is valid if we define the precedence and associativity of all the operators. Thus, it is very important when you write a parser using CUP or any other LALR(1) parser generator to specify associativities and precedence‟s for most tokens (especially for those used as operators). Note: you can explicitly define the precedence of a rule in CUP using the %prec directive:

E ::= MINUS E %prec UMINUS

where UMINUS is a pseudo-token that has higher precedence than TIMES, MINUS etc, so that -1*2 is equal to (-1)*2, not to -(1*2).

Another thing we can do when specifying an LALR(1) grammar for a parser generator is error recovery. All the entries in the ACTION and GOTO tables that have no content correspond to syntax errors. The simplest thing to do in case of error is to report it and stop the parsing. But we would like to continue parsing finding more errors. This is called *error recovery*. Consider the grammar:

```
S ::= L = E ;
   | { SL }
   ; | error ;
SL ::= S ; |
    SL S ;
```

The special token error indicates to the parser what to do in case of invalid syntax for S (an invalid statement). In this case, it reads all the tokens from the input stream until it finds the first semicolon. The way the parser handles this is to first push an error state in the stack. In case of an error, the parser pops out elements from the stack until it finds an error state where it can proceed. Then it discards tokens from the input until a restart is possible. Inserting error handling productions in the proper places in a grammar to do good error recovery is considered very hard.

LR ERROR RECOVERY

An LR parser will detect an error when it consults the parsing action table and find a blank or error entry. Errors are never detected by consulting the goto table. An LR parser will detect an error as soon as there is no valid continuation for the portion of the input thus far

scanned. A canonical LR parser will not make even a single reduction before announcing the error. SLR and LALR parsers may make several reductions before detecting an error, but they will never shift an erroneous input symbol onto the stack.

PANIC-MODE ERROR RECOVERY

We can implement panic-mode error recovery by scanning down the stack until a state s with a goto on a particular nonterminal A is found. Zero or more input symbols are then discarded until a symbol a is found that can legitimately follow A. The parser then stacks the state GOTO(s, A) and resumes normal parsing. The situation might exist where there is more than one choice for the nonterminal A. Normally these would be nonterminals representing major program pieces, e.g. an expression, a statement, or a block. For example, if A is the nonterminal stmt, a might be semicolon or }, which marks the end of a statement sequence. This method of error recovery attempts to eliminate the phrase containing the syntactic error. The parser determines that a string derivable from A contains an error. Part of that string has already been processed, and the result of this processing is a sequence of states on top of the stack. The remainder of the string is still in the input, and the parser attempts to skip over the remainder of this string by looking for a symbol on the input that can legitimately follow A. By removing states from the stack, skipping over the input, and pushing GOTO(s, A) on the stack, the parser pretends that if has found an instance of A and resumes normal parsing.

PHRASE-LEVEL RECOVERY

Phrase-level recovery is implemented by examining each error entry in the LR action table and deciding on the basis of language usage the most likely programmer error that would give rise to that error. An appropriate recovery procedure can then be constructed; presumably the top of the stack and/or first input symbol would be modified in a way deemed appropriate for each error entry. In designing specific error-handling routines for an LR parser, we can fill in each blank entry in the action field with a pointer to an error routine that will take the appropriate action selected by the compiler designer.

The actions may include insertion or deletion of symbols from the stack or the input or both, or alteration and transposition of input symbols. We must make our choices so that the LR parser will not get into an infinite loop. A safe strategy will assure that at least one input symbol will be removed or shifted eventually, or that the stack will eventually shrink if the end of the input has been reached. Popping a stack state that covers a non terminal should be avoided, because this modification eliminates from the stack a construct that has already been successfully parsed.

Syntax Directed Translations

We associate information with a language construct by attaching attributes to the grammar symbol(s) representing the construct, A syntax-directed definition specifies the values of attributes by associating semantic rules with the grammar productions. For example, an infix-to-postfix translator might have a production and rule

PRODUCTION	SEMANTIC RULE
$E \rightarrow Ei + T$	$E.code = Ei.code \parallel T.code \parallel '+'$

This production has two nonterminals, E and T; the subscript in E1 distinguishes the occurrence of E in the production body from the occurrence of E as the head. Both E and T have a string-valued attribute code. The semantic rule specifies that the string E.code is formed by concatenating Ei.code, T.code, and the character '+'. While the rule makes it explicit that the translation of E is built up from the translations of E1, T, and '+', it may be inefficient to implement the translation directly by manipulating strings.

a syntax-directed translation scheme embeds program fragments called semantic actions within production bodies
There are two notations for attaching semantic rules:

1. **Syntax Directed Definitions.** High-level specification hiding many implementation details (also called **Attribute Grammars**).

2. **Translation Schemes.** More implementation oriented: Indicate the order in which semantic rules are to be evaluated.

Syntax Directed Definitions

Syntax Directed Definitions are a generalization of context-free grammars in which:

1. Grammar symbols have an associated set of **Attributes**;
2. Productions are associated with **Semantic Rules** for computing the values of attributes Such formalism generates **Annotated Parse-Trees** where each node of the tree is a record with a field for each attribute (e.g.,X.a indicates the attribute a of the grammar symbol X).

The value of an attribute of a grammar symbol at a given parse-tree node is defined by a semantic rule associated with the production used at that node.

We distinguish between two kinds of attributes:

1. **Synthesized Attributes.** They are computed from the values of the attributes of the children nodes.

2. **Inherited Attributes.** They are computed from the values of the attributes of both the siblings and the parent nodes

Syntax Directed Definitions: An Example

Let us consider the Grammar for arithmetic expressions. The Syntax Directed Definition associates to each non terminal a synthesized attribute called *val*.

PRODUCTION	SEMANTIC RULE
PRODUCTION	SEMANTIC RULE
$E \rightarrow E_1 + T$	$E.code = E_1.code \parallel T.code \parallel '+'$
$F \rightarrow (E)$	$F.val := E.val$
$F \rightarrow$ digit	$F.val :=$ digit.*lexval*

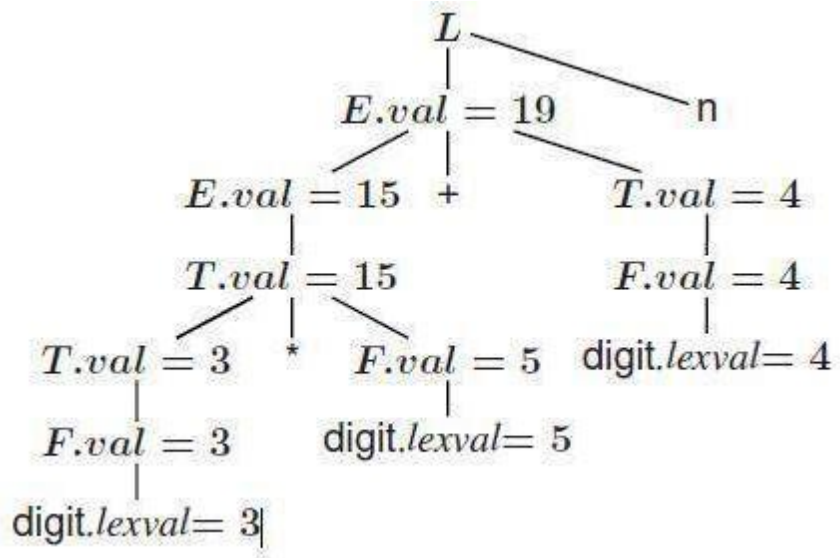
SDD of a simple desk calculator

S-ATTRIBUTED DEFINITIONS

Definition. An **S-Attributed Definition** is a Syntax Directed Definition that uses only synthesized attributes.

• **Evaluation Order.** Semantic rules in a S-Attributed Definition can be evaluated by a bottom-up, or PostOrder, traversal of the parse-tree.

• **Example.** The above arithmetic grammar is an example of an S-Attributed Definition. The annotated parse-tree for the input 3*5+4n is:



L-attributed definition

Definition: A SDD its *L-attributed* if each inherited attribute of Xi in the RHS of A ! X1 :

:Xn depends only on

1. attributes of X1;X2; : : : ;Xi 1 (symbols to the left of Xi in the RHS)
2. inherited attributes of A.

Restrictions for translation schemes:

1. Inherited attribute of Xi must be computed by an action before Xi.
2. An action must not refer to synthesized attribute of any symbol to the right of that action.
3. Synthesized attribute for A can only be computed after all attributes it references have been completed (usually at end of RHS).

Evaluation order of SDTS

1 Dependency Graphs

2 Ordering the Evaluation of Attributes

3 S-Attributed Definitions

4 L-Attributed Definitions

"Dependency graphs" are a useful tool for determining an evaluation order for the attribute instances in a given parse tree. While an annotated parse tree shows the values of attributes, a dependency graph helps us determine how those values can be computed.

1 Dependency Graphs

A *dependency graph* depicts the flow of information among the attribute in-stances in a particular parse tree; an edge from one attribute instance to an-other means that the value of the first is needed to compute the second. Edges express constraints implied by the semantic rules. In more detail:

Suppose that a semantic rule associated with a production *p* defines the value of inherited attribute *B.c* in terms of the value of *X.a.* Then, the dependency graph has an edge from *X.a* to *B.c.* For each node *N* labeled *B* that corresponds to an occurrence of this *B* in the body of production *p,* create an edge to attribute c at *N* from the attribute *a* at the node *M* that corresponds to this occurrence of *X.* Note that *M* could be either the parent or a sibling of *N.*

Since a node ***N*** can have several children labeled ***X,*** we again assume that subscripts distinguish among uses of the same symbol at different places in the production.

Example: Consider the following production and rule:

PRODUCTION	SEMANTIC RULE
$E \rightarrow E_1 + T$	$E.val = E_1.val + T.val$

At every node *N* labeled *E*, with children corresponding to the body of this production, the synthesized attribute *val* at *N* is computed using the values of *val* at the two children, labeled *E* and *T*. Thus, a portion of the dependency graph for every parse tree in which this production is used looks like Fig. 5.6. As a convention, we shall show the parse tree edges as dotted lines, while the edges of the dependency graph are solid.

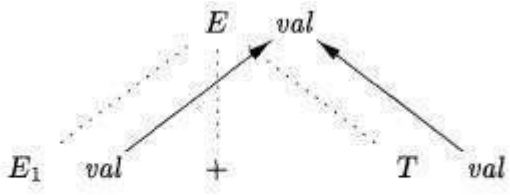


Figure 5.6: $E.val$ is synthesized from $E_1.val$ and $E_2.val$

2. Ordering the Evaluation of Attributes

The dependency graph characterizes the possible orders in which we can evalu-ate the attributes at the various nodes of a parse tree. If the dependency graph has an edge from node M to node N, then the attribute corresponding to M must be evaluated before the attribute of N. Thus, the only allowable orders of evaluation are those sequences of nodes N1, N2,... ,Nk such that if there is an edge of the dependency graph from Ni to Nj, then i < j. Such an ordering embeds a directed graph into a linear order, and is called a topological sort of the graph.

If there is any cycle in the graph, then there are no topological sorts; that is, there is no way to evaluate the SDD on this parse tree. If there are no cycles, however, then there is always at least one topological sort

3. S-Attributed Definitions

An SDD is *S-attributed* if every attribute is synthesized. When an SDD is S-attributed, we can evaluate its attributes in any bottom-up order of the nodes of the parse tree. It is often especially simple to evaluate the attributes by performing a postorder traversal of the parse tree and evaluating the attributes at a node *N* when the traversal leaves *N* for the last time.

S-attributed definitions can be implemented during bottom-up parsing, since a bottom-up parse corresponds to a postorder traversal. Specifically, postorder corresponds exactly to the order in which an LR parser reduces a production body to its head.

4 L-Attributed Definitions

The idea behind this class is that, between the attributes associated with a production body, dependency-graph edges can go from left to right, but not from right to left (hence "L-attributed"). More precisely, each attribute must be either

1. Synthesized, or
2. Inherited, but with the rules limited as follows. Suppose that there is a production A -> X1 X2 ..... Xn, and that there is an inherited attribute Xi.a computed by a rule associated with this production.

Then the rule may use only:
Inherited attributes associated with the head A.
Either inherited or synthesized attributes associated with the occurrences of symbols X1, X2,... , X(i-1) located to the left of Xi.
Inherited or synthesized attributes associated with this occurrence of Xi itself, but only in such a way that there are no cycles in a dependency graph formed by the attributes of this X i

Application of SDTS

1 Construction of Syntax Trees
2 The Structure of a Type

The main application is the construction of syntax trees. Since some compilers use syntax trees as an intermediate representation, a common form of SDD turns its input string into a tree. To complete the translation to intermediate code, the compiler may then walk the syntax tree, using another set of rules that are in effect an SDD on the syntax tree rather than the parse tree.

1 Construction of Syntax Trees

Each node in a syntax tree represents a construct; the children of the node represent the meaningful components of the construct. A syntax-tree node representing an expression $E1 + E_2$ has label + and two children representing the subexpressions *E1* and E_2

implement the nodes of a syntax tree by objects with a suitable number of fields. Each object will have an *op* field that is the label of the node.
The objects will have additional fields as follows:

• If the node is a leaf, an additional field holds the lexical value for the leaf. A constructor function Leaf (op, val) creates a leaf object. Alternatively, if nodes are viewed as records, then Leaf returns a pointer to a new record for a leaf.

• If the node is an interior node, there are as many additional fields as the node has children in the syntax tree. A constructor function Node takes two or more arguments: Node(op,ci,c2,... ,ck) creates an object with first field op and k additional fields for the k children c1,... , .

Example

	PRODUCTION	SEMANTIC RULES
1)	$E \to E_1 + T$	$E.node = \textbf{new}\ Node('+', E_1.node, T.node)$
2)	$E \to E_1 - T$	$E.node = \textbf{new}\ Node('-', E_1.node, T.node)$
3)	$E \to T$	$E.node = T.node$
4)	$T \to (\ E\ )$	$T.node = E.node$
5)	$T \to \textbf{id}$	$T.node = \textbf{new}\ Leaf(\textbf{id}, \textbf{id}.entry)$
6)	$T \to \textbf{num}$	$T.node = \textbf{new}\ Leaf(\textbf{num}, \textbf{num}.val)$

Figure 5.10: Constructing syntax trees for simple expressions

Figure 5.1 1 shows the construction of a syntax tree for the input *a* — 4 + *c*. The nodes of the syntax tree are shown as records, with the *op* field first. Syntax-tree edges are now shown as solid lines. The underlying parse tree, which need not actually be constructed, is shown with dotted edges. The

third type of line, shown dashed, represents the values of *E.node* and *T-node;* each line points to the appropriate syntax-tree node.

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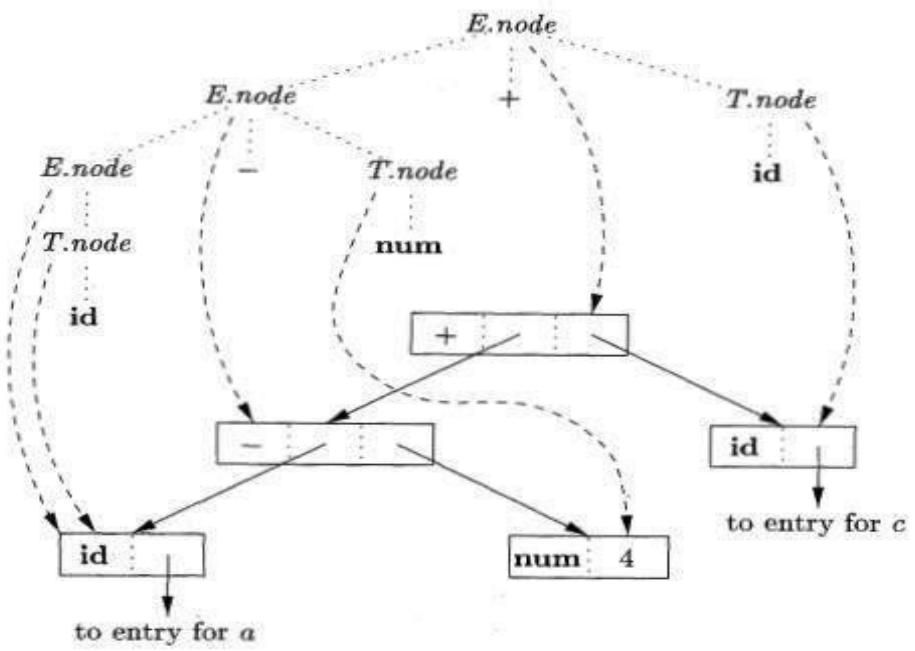


Figure 5.11: Syntax tree for $a - 4 + c$

1) $p_1 = \textbf{new}\ \mathit{Leaf}(\textbf{id}, \mathit{entry\text{-}a});$
2) $p_2 = \textbf{new}\ \mathit{Leaf}(\textbf{num}, 4);$
3) $p_3 = \textbf{new}\ \mathit{Node}('-', p_1, p_2);$
4) $p_4 = \textbf{new}\ \mathit{Leaf}(\textbf{id}, \mathit{entry\text{-}c});$
5) $p_5 = \textbf{new}\ \mathit{Node}('+', p_3, p_4);$

Figure 5.12: Steps in the construction of the syntax tree for $a - 4 + c$

2 The Structure of a Type

The type int [2][3] can be read as, "array of 2 arrays of 3 integers." The corresponding type expression array(2, array(3, integer)) is represented by the tree in Fig. 5.15. The operator array takes two parameters, a number and a type. If types are represented by trees, then this operator returns a tree node labeled array with two children for a number and a type.

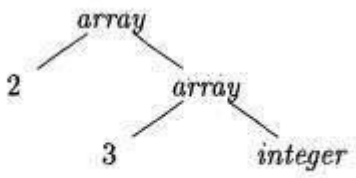


Figure 5.15: Type expression for **int**[2][3]

Nonterminal *B* generates one of the basic types **int** and **float.** *T* generates a basic type when *T* derives *B C* and *C* derives e. Otherwise, *C* generates array components consisting of a sequence of integers, each integer surrounded by brackets.

PRODUCTION	SEMANTIC RULES
$T \rightarrow B\ C$	$T.t = C.t$
	$C.b = B.t$
$B \rightarrow$ **int**	$B.t = integer$
$B \rightarrow$ **float**	$B.t = float$
$C \rightarrow$ [**num**] C_1	$C.t = array(\mathbf{num}.val,\ C_1.t)$
	$C_1.b = C.b$
$C \rightarrow \epsilon$	$C.t = C.b$

Figure 5.16: T generates either a basic type or an array type

An annotated parse tree for the input string **int [** 2] [3] is shown in Fig. 5.17. The corresponding type expression in Fig. 5.15 is constructed by passing the type *integer* from *B,* down the chain of C's through the inherited attributes *b.* The array type is synthesized up the chain of C's through the attributes *t.*

In more detail, at the root for *T* -» *B C,* nonterminal *C* inherits the type from *B,* using the inherited attribute *C.b.* At the rightmost node for C, the production is C e, so C.t equals C.6. The semantic rules for the production C [num] C1 form C.t by applying the operator array to the operands num.ua/ and C1.t.

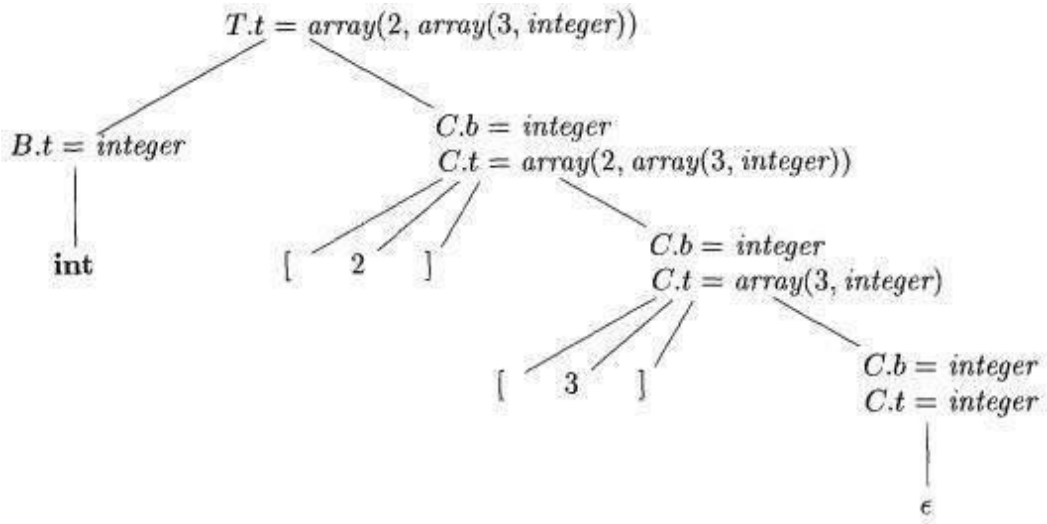


Figure 5.17: Syntax-directed translation of array types

Syntax Directed Translation Schemes.

1 Postfix Translation Schemes
2 Parser-Stack Implementation of Postfix SDT's
3 SDT's With Actions Inside Productions
4 Eliminating Left Recursion From SDT's

syntax-directed translation scheme (SDT) is a context-free grammar with program fragments embedded within production bodies. The program fragments are called *semantic actions* and can appear at any position within a production body. By convention, we place curly braces around actions; if braces are needed as grammar symbols, then we quote them.SDT's are implemented during parsing, without building a parse tree.

Two important classes of SDD's are

1. The underlying grammar is LR-parsable, and the SDD is S-attributed.
2. The underlying grammar is LL-parsable, and the SDD is L-attributed.

1 Postfix Translation Schemes

simplest SDD implementation occurs when we can parse the grammar bottom-up and the SDD is S-attributed. In that case, we can construct an SDT in which each action is placed at the end of the production and is executed along with the reduction of the body to the head of that production. SDT's with all actions at the right ends of the production bodies are called postfix SDT's.

Example 5.14 : The postfix SDT in Fig. 5.18 implements the desk calculator SDD of Fig. 5.1, with one change: the action for the first production prints a value. The remaining actions are exact counterparts of the semantic rules. Since the underlying grammar is LR, and the SDD is S-attributed, these actions can be correctly performed along with the reduction steps of the parser.

$$
\begin{array}{lcll}
L & \rightarrow & E\ \mathbf{n} & \{\ \mathrm{print}(E.val);\ \} \\
E & \rightarrow & E_1 + T & \{\ E.val = E_1.val + T.val;\ \} \\
E & \rightarrow & T & \{\ E.val = T.val;\ \} \\
T & \rightarrow & T_1 * F & \{\ T.val = T_1.val \times F.val;\ \} \\
T & \rightarrow & F & \{\ T.val = F.val;\ \} \\
F & \rightarrow & (\ E\ ) & \{\ F.val = E.val;\ \} \\
F & \rightarrow & \mathbf{digit} & \{\ F.val = \mathbf{digit}.lexval;\ \}
\end{array}
$$

Figure 5.18: Postfix SDT implementing the desk calculator

2 Parser-Stack Implementation of Postfix SDT's

The attribute(s) of each grammar symbol can be put on the stack in a place where they can be found during the reduction. The best plan is to place the attributes along with the grammar symbols (or the LR states that represent these symbols) in records on the stack itself.

In Fig. 5.19, the parser stack contains records with a field for a grammar symbol (or parser state) and, below it, a field for an attribute. The three grammar symbols *X YZ* are on top of the stack; perhaps they

are about to be reduced according to a production like *A —> X YZ.* Here, we show *X.x* as the one attribute of *X*, and so on. In general, we can allow for more attributes, either by making the records large enough or by putting pointers to records on the stack. With small attributes, it may be simpler to make the records large enough, even if some fields go unused some of the time. However, if one or more attributes are of unbounded size — say, they are character strings — then it would be better to put a pointer to the attribute's value in the stack record and store the actual value in some larger, shared storage area that is not part of the stack.

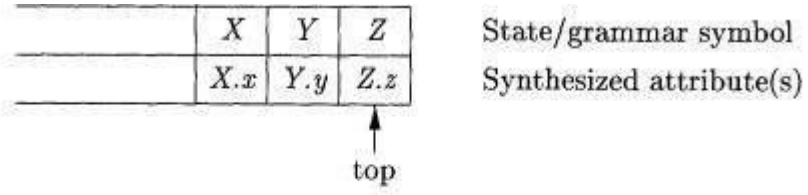


Figure 5.19: Parser stack with a field for synthesized attributes

3 SDT's With Actions Inside Productions

An action may be placed at any position within the body of a production.It is performed immediately after all symbols to its left are processed. Thus,if we have a production B -» X {a} Y, the action a is done after we have recognized X (if X is a terminal) or all the terminals derived from X (if X is a nonterminal).

More precisely,

• If the parse is bottom-up, then we perform action a as soon as this occurrence of X appears on the top of the parsing stack.

• If the parse is top-down, we perform a just before we attempt to expand this occurrence of Y (if Y a nonterminal) or check for Y on the input (if Y is a terminal).

4 Eliminating Left Recursion From SDT's

First, consider the simple case, in which the only thing we care about is the order in which the actions in an SDT are performed. For example, if each action simply prints a string, we care only about the order in which the strings are printed. In this case, the following principle can guide us:

When transforming the grammar, treat the actions as if they were terminal symbols.

This principle is based on the idea that the grammar transformation preserves the order of the terminals in the generated string. The actions are therefore executed in the same order in any left-to-right parse, top-down or bottom-up.

The "trick" for eliminating left recursion is to take two productions

$$A \rightarrow Aa \mid b$$

that generate strings consisting of a *j3* and any number of en's, and replace them by productions that generate the same strings using a new nonterminal *R* (for "remainder") of the first production:

A->bR

R —»• aR | e

If *(3* does not begin with *A,* then *A* no longer has a left-recursive production. In regular-definition terms, with both sets of productions, *A* is defined by ***0(a)****.

Example **5** . 1 7 : Consider the following E-productions from an SDT for translating infix expressions into postfix notation:

E -> E i + T { print('+'); }

E -> T

If we apply the standard transformation to E, the remainder of the left-recursive production is

a = + T { print('-r'); }

and the body of the other production is T. If we introduce R for the remainder of E, we get the set of productions:

E --> T R

R --> + T { printC-h'); } R

R -> e

When the actions of an SDD compute attributes rather than merely printing output, we must be more careful about how we eliminate left recursion from a grammar. However, if the SDD is S-attributed, then we can always construct an SDT by placing attribute-computing actions at appropriate positions in the new productions.

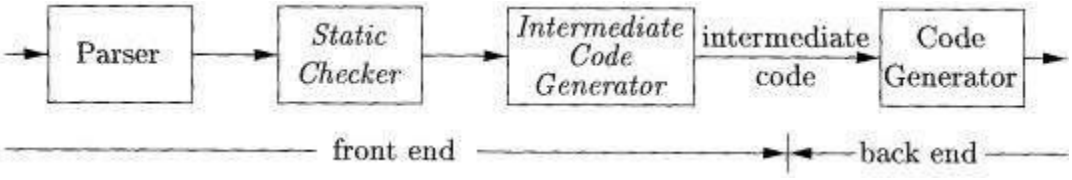
UNIT – IV

Intermediated Code: Generation Variants of Syntax trees 3Addresscode, Types and Deceleration, Translation of Expressions, Type Checking. Canted Flow Back patching?

UNIT 4
INTERMEDIATE CODE

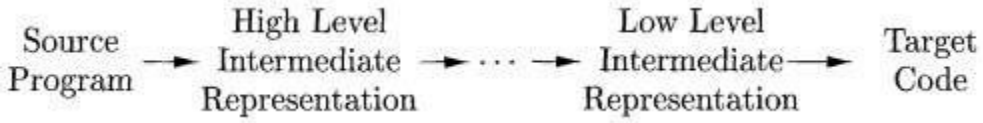
In the analysis-synthesis model of a compiler, the front end analyzes a source program and creates an intermediate representation, from which the back end generates target code. This facilitates *retargeting*: enables attaching a back end for the new machine to an existing front end.

Logical Structure of a Compiler Front End



A compiler front end is organized as in figure above, where parsing, static checking, and intermediate-code generation are done sequentially; sometimes they can be combined and folded into parsing. All schemes can be implemented by creating a syntax tree and traversing the tree.

Static checking includes *type checking,* which ensures that operators are applied to compatible operands. In the process of translating a program in a given source language into code for a given target machine, a compiler construct a sequence of intermediate representations



Sequence of intermediate representations

High-level representations are close to the source language and low-level representations are close to the target machine. A low-level representation is suitable for machine-dependent tasks like register allocation and instruction selection.

Variants of Syntax Trees

1 Directed Acyclic Graphs for Expressions

2 The Value-Number Method for Constructing DAG's

1. Directed Acyclic Graphs for Expressions

Like the syntax tree for an expression, a DAG has leaves corresponding to atomic operands and interior codes corresponding to operators. The difference is that a node *N* in a DAG has more than one parent if *N* represents a com-mon subexpression; in a syntax tree, the tree for the common subexpression would be replicated as many times as the subexpression appears in the original expression.

Example: Consider expression

a + a * (b - c) + (b - c) * d

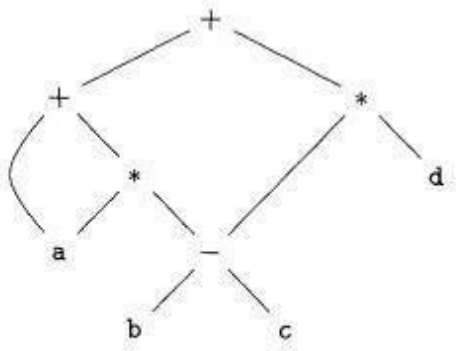
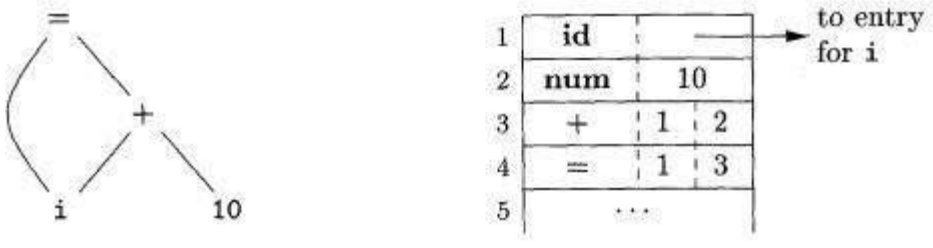


Figure 6.3: Dag for the expression a+a*(b-c)+(b-c)*d

2 The Value-Number Method for Constructing DAG's

The nodes of a syntax tree or DAG are stored in an array of records, as suggested by Fig. 6.6. Each row of the array represents one record, and therefore one node. In each record, the first field is an operation code, indicating the label of the node. In Fig. 6.6(b), leaves have one additional field, which holds the lexical value (either a symbol-table pointer or a constant, in this case), and interior nodes have two additional fields indicating the left and right children.



(a) DAG (b) Array.

Figure 6.6: Nodes of a DAG for $i = i + 10$ allocated in an array

In this array, we refer to nodes by giving the integer index of the record for that node within the array. This integer is called the value number for the node.

Algorithm: The value-number method for constructing the nodes of a DAG.

INPUT : Label op, node /, and node r.

OUTPUT : The value number of a node in the array with signature (op, l,r).

METHOD : Search the array for a node M with label op, left child I, and right child r. If there is such a

node, return the value number of M. If not, create in the array a new node N with label op, left child I, and right child r, and return its value number.

Three-Address Code

1 Addresses and Instructions

2 Quadruples

3 Triples

In three-address code, there is at most one operator on the right side of an instruction; that is, no built-up arithmetic expressions are permitted. Thus a source-language expression like x+y*z might be translated into the sequence of three-address instructions

$$t_1 = y * z$$
$$t_2 = x + t_1$$

where ti and t_2 are compiler-generated temporary names.

1 Addresses and Instructions

An address can be one of the following:

- *A name.* Source-program names to appear as addresses in three-address code. In an implementation, a source name is replaced by a pointer to its symbol-table entry, where all information about the name is kept.
- *A constant.* A compiler must deal with many different types of constants and variables.
- *A compiler-generated temporary.* Useful in optimizing com-pilers, to create a distinct name each time a temporary is needed. These temporaries can be combined, if possible, when registers are allocated to variables.

common three-address instruction

1. Assignment Statement: x = y op z and x = op y

Here,x, y and z are the operands. op represents the operator.

2. Copy Statement: x = y

3. Conditional Jump: If x relop y goto X

If the condition "x relop y" gets **satisfied,** then-

The control is sent directly to the location specified by label X.

All the statements in between are skipped.

If the condition "x relop y" **fails**, then-

The control is not sent to the location specified by label X.

The next statement appearing in the usual sequence is executed.

4. Unconditional Jump- goto X

On executing the statement,The control is sent directly to the location specified by label X.

All the statements in between are skipped.

5. Procedure Call- param x call p return y

Here, p is a function which takes x as a parameter and returns y.

For a procedure call p(x1, …, xn)

param x1

…

param xn

call p, n

6. Indexed copy instructions: x = y[i] and x[i] = y

Left: sets x to the value in the location i memory units beyond y

Right: sets the contents of the location i memory units beyond x to y

7. Address and pointer instructions:

x = &y sets the value of x to be the location (address) of y.

x = *y, presumably y is a pointer or temporary whose value is a location. The value of x is set to the contents of that location.

*x = y sets the value of the object pointed to by x to the value of y.

Data Structure

Three address code is represented as record structure with fields for operator and operands. These records can be stored as array or linked list. Most common implementations of three address code are Quadruples, Triples and Indirect triples.

2. Quadruples

Quadruples consists of four fields in the record structure. One field to store operator op, two fields to store operands or arguments arg1and arg2 and one field to store result res.

res = arg1 op arg2

Example: a = b + c

b is represented as arg1, c is represented as arg2, + as op and a as res.

Unary operators like „-„do not use agr2. Operators like param do not use agr2 nor result. For conditional and unconditional statements res is label. Arg1, arg2 and res are pointers to symbol table or literal table for the names.

Example: a = -b * d + c + (-b) * d

Three address code for the above statement is as follows

t1 = - b

t2 = t1 * d

t3 = t2 + c

t4 = - b

t5 = t4 * d

t6 = t3 + t5

a = t6

Quadruples for the above example is as follows

Op	Arg1	Arg2	Res
-	B		t1
*	t1	d	t2
+	t2	c	t3
-	B		t4
*	t4	d	t5
+	t3	t5	t6
=	t6		a

3 TRIPLES

Triples uses only three fields in the record structure. One field for operator, two fields for operands named as arg1 and arg2. Value of temporary variable can be accessed by the position of the statement the computes it and not by location as in quadruples.

Example: a = -b * d + c + (-b) * d

Triples for the above example is as follows

Stmt no	Op	Arg1	Arg2
(0)	-	b	
(1)	*	d	(0)
(2)	+	c	(1)
(3)	-	b	
(4)	*	d	(3)
(5)	+	(2)	(4)
(6)	=	a	(5)

Arg1 and arg2 may be pointers to symbol table for program variables or literal table for constant or pointers into triple structure for intermediate results.

Example: Triples for statement x[i] = y which generates two records is as follows

Stmt no	Op	Arg1	Arg2
(0)	[]=	x	i
(1)	=	(0)	y

Triples for statement x = y[i] which generates two records is as follows

Stmt no	Op	Arg1	Arg2
(0)	=[]	y	i
(1)	=	x	(0)

Triples are alternative ways for representing syntax tree or Directed acyclic graph for program defined names.

Indirect Triples

Indirect triples are used to achieve indirection in listing of pointers. That is, it uses pointers to triples than listing of triples themselves.

Example: a = -b * d + c + (-b) * d

	Stmt no
(0)	(10)
(1)	(11)
(2)	(12)
(3)	(13)
(4)	(14)
(5)	(15)
(6)	(16)

Stmt no	Op	Arg1	Arg2
(10)	-	b	
(11)	*	d	(0)
(12)	+	c	(1)
(13)	-	b	
(14)	*	d	(3)
(15)	+	(2)	(4)
(16)	=	a	(5)

Types and Declarations

1 Type Expressions
2 Type Equivalence
3 Declarations
4 Storage Layout for Local Names

1 Type Expressions

Types have structure, which we shall represent using *type expressions:* a type expression is either a basic type or is formed by applying an operator called a *type constructor* to a type expression.

Definition

- A basic type is a type expression. Typical basic types for a language include *boolean, char, integer, float,* and *void;* the latter denotes "the absence of a value."
- A type name is a type expression.
- A type expression can be formed by applying the array type constructor to a number and a type
- expression.
- A record is a data structure with named fields. A type expression can be formed by applying the *record* type constructor to the field names and their types.
- If *s* and *t* are type expressions, then their Cartesian product *s* x *t* is a type expression. Products are introduced for completeness; they can be used to represent a list or tuple of types (e.g., for function parameters).
- Type expressions may contain variables whose values are type expressions

2 Type Equivalence

Many type-checking rules have the form, "if two type expressions are equal then return a certain type else error." Potential ambiguities arise when names are given to type expressions. The key issue is whether a name in a type expression stands for itself or whether it is an abbreviation for another type expression.

Since type names denote type expressions, they can set up implicit cycles; see the box on "Type Names and Recursive Types." If edges to type names are redirected to the type expressions denoted by the names, then the resulting graph can have cycles due to recursive types.

When type expressions are represented by graphs, two types are *structurally equivalent* if and only if one of the following conditions is true:

They are the same basic type.
They are formed by applying the same constructor to structurally equivalent types.
One is a type name that denotes the other.

If type names are treated as standing for themselves, then the first two conditions in the above definition lead to *name equivalence* of type expressions.

Name-equivalent expressions are assigned the same value number,. Structural equivalence can be tested using the unification algorithm .

3. Declarations

Understand types and declarations using a simplified grammar that declares just one name at a time; The grammar is

$$
\begin{array}{lcl}
D & \rightarrow & T\ \mathbf{id}\ ;\ D \mid \epsilon \\
T & \rightarrow & B\ C \mid \mathbf{record}\ '\{'\ D\ '\}' \\
B & \rightarrow & \mathbf{int} \mid \mathbf{float} \\
C & \rightarrow & \epsilon \mid [\ \mathbf{num}\ ]\ C
\end{array}
$$

The fragment of the above grammar that deals with basic and array types.Consider storage layout as well as types.Nonterminal *D* generates a sequence of declarations. Nonterminal *T* generates basic, array, or record types. Nonterminal *B* generates one of the basic types int and float. Nonterminal C, for "component," generates strings of zero or more integers, each integer surrounded by brackets. An array type consists of a basic type specified by *B,* followed by array components specified by nonterminal *C. A* record type (the second production for *T)* is a sequence of declarations for the fields of the record, all surrounded by curly braces.

4. Storage Layout for Local Names

From the type of a name, we can determine the amount of storage that will be needed for the name at run time. At compile time, we can use these amounts to assign each name a relative address. The type and relative address are saved in the symbol-table entry for the name. Data of varying length, such as strings, or data whose size cannot be determined until run time, such as dynamic arrays, is handled by reserving a known fixed amount of storage for a pointer to the data.

Address Alignment

The storage layout for data objects is strongly influenced by the address-ing constraints of the target machine. For example, instructions to add integers may expect integers to be *aligned*, that is, placed at certain positions in memory such as an address divisible by 4. Although an array of ten characters needs only enough bytes to hold ten characters, a compiler may therefore allocate 12 bytes — the next multiple of 4 — leaving 2 bytes unused. Space left unused due to alignment considerations is referred to as *padding.* When space is at a premium, a compiler may *pack* data so that no padding is left; additional instructions may then need to be executed at run time to position packed data so that it can be operated on as if it were properly aligned.

Suppose that storage comes in blocks of contiguous bytes, where a byte is the smallest unit of addressable memory. The *width* of a type is the number of storage units needed for objects of that type. A basic type, such as a character, integer, or float, requires an integral number of bytes. For easy access, storage for aggregates such as arrays and classes is allocated in one contiguous block of bytes.

The translation scheme (SDT) computes types and their widths for basic and array types; The SDT uses synthesized attributes *type* and *width* for each nonterminal and two variables *t* and *w* to pass type and width information from a *B* node in a parse tree to the node for the production *C* —> e. In a syntax-directed definition, *t* and *w* would be inherited attributes for *C.*

The body of the T-production consists of nonterminal B, an action, and nonterminal C, which appears on the next line. The action between B and C sets t to B.type and w to B. width. If B —>• int then B.type is set to integer and B.width is set to 4, the width of an integer. Similarly, if B -+ float then B.type is float and B.width is 8, the width of a float.

The productions for C determine whether T generates a basic type or an array type. If C —>• e, then t becomes C.type and w becomes C. width. Otherwise, C specifies an array component. The action for C —> [n u m] C1 forms C.type by applying the type constructor array to the operands num.value and C1.type.

$$
\begin{array}{lll}
T \rightarrow B & & \{\ t = B.type;\ w = B.width;\ \} \\
\quad\ C & & \\
B \rightarrow \textbf{int} & & \{\ B.type = integer;\ B.width = 4;\ \} \\
B \rightarrow \textbf{float} & & \{\ B.type = float;\ B.width = 8;\ \} \\
C \rightarrow \epsilon & & \{\ C.type = t;\ C.width = w;\ \} \\
C \rightarrow [\ \textbf{num}\ ]\ C_1 & & \{\ array(\textbf{num}.value,\ C_1.type); \\
 & & \quad C.width = \textbf{num}.value \times C_1.width;\ \}
\end{array}
$$

Figure 6.15: Computing types and their widths

The width of an array is obtained by multiplying the width of an element by the number of elements in the array. If addresses of consecutive integers differ by 4, then address calculations for an array of integers will include multiplications by 4. Such multiplications provide opportunities for optimization, so it is helpful for the front end to make them explicit.

Example The parse tree for the type i n t [2] [3] is shown by dotted lines in Fig. 6.16. The solid lines show how the type and width are passed from B, down the chain of C's through variables t and w, and then back up the chain as synthesized attributes type and width. The variables t and w are assigned the values of B.type and B.width, respectively, before the subtree with the C nodes is examined. The values of t and w are used at the node for C —> e to start the evaluation of the synthesized attributes up the chain of C nodes.

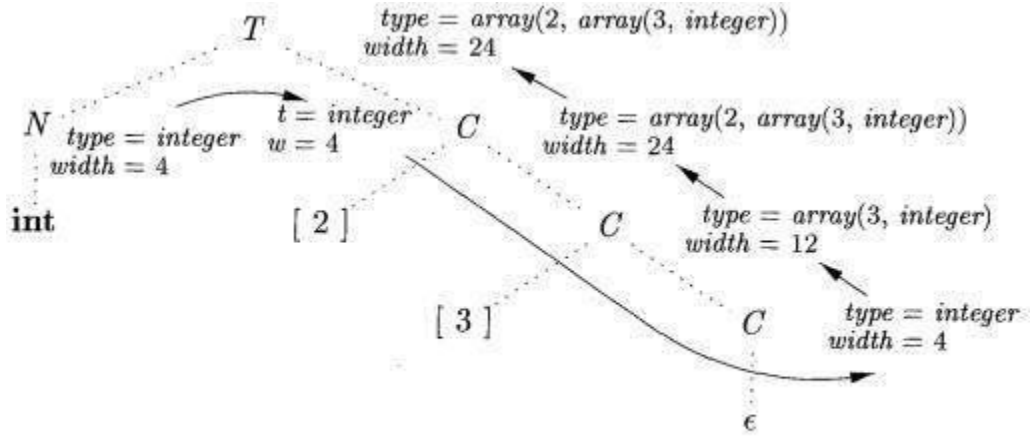


Figure 6.16: Syntax-directed translation of array types

Translations of Expressions

1 Operations Within Expressions
2 Incremental Translation
3 Addressing Array Elements
4 Translation of Array References

1 Operations Within Expressions

The syntax-directed definition builds up the three-address code for an assignment statement *S* using attribute *code* for *S* and attributes *addr* and *code* for anexpression *E.* Attributes *S.code* and *E.code* denote the three-address code for *S* and *E*, respectively. Attribute *E.addr* denotes the address that will hold the value of E.

PRODUCTION	SEMANTIC RULES
$S \rightarrow \mathbf{id} = E\ ;$	$S.code = E.code \mid\mid gen(top.get(\mathbf{id}.lexeme)\ '='\ E.addr)$
$E \rightarrow E_1 + E_2$	$E.addr = \mathbf{new}\ Temp()$ $E.code = E_1.code \mid\mid E_2.code \mid\mid gen(E.addr\ '='\ E_1.addr\ '+'\ E_2.addr)$
$\mid\ - E_1$	$E.addr = \mathbf{new}\ Temp()$ $E.code = E_1.code \mid\mid gen(E.addr\ '='\ '\mathbf{minus}'\ E_1.addr)$
$\mid\ (E_1)$	$E.addr = E_1.addr$ $E.code = E_1.code$
$\mid\ \mathbf{id}$	$E.addr = top.get(\mathbf{id}.lexeme)$ $E.code = ''$

Figure 6.19: Three-address code for expressions

Example The syntax-directed definition in Fig. 6.19 translates the assignment statement a = b + - c; into the TAC

```
t1 = minus c
t2 = b + t1
a = t2
```

2 Incremental Translation

Code attributes can be long strings, so they are generated incrementally In the incremental approach, *gen* not only constructs a three-address instruction, it appends the instruction to the sequence of instructions generated so far. The sequence may either be retained in memory for further processing, or it may be output incrementally.attribute addr represents the address of a node rather than a variable or constant.

$S \rightarrow$	**id** = E ;	{ *gen*(*top.get*(**id**.*lexeme*) '=' E.*addr*); }
$E \rightarrow$	E_1 + E_2	{ E.*addr* = **new** *Temp*(); *gen*(E.*addr* '=' E_1.*addr* '+' E_2.*addr*); }
|	- E_1	{ E.*addr* = **new** *Temp*(); *gen*(E.*addr* '=' '**minus**' E_1.*addr*); }
|	(E_1)	{ E.*addr* = E_1.*addr*; }
|	**id**	{ E.*addr* = *top.get*(**id**.*lexeme*); }

Figure 6.20: Generating three-address code for expressions incrementally

3.Addressing Array Elements

Elements of arrays can be accessed quickly if the elements are stored in a block of consecutive location. Array can be one dimensional or two dimensional.

For one dimensional array:

A: array[low..high] of the ith elements is at:

base + (i-low)*width = i*width + (base - low*width)

Multi-dimensional arrays:

Row major or column major forms

- Row major: a[1,1], a[1,2], a[1,3], a[2,1], a[2,2], a[2,3]
- Column major: a[1,1], a[2,1], a[1, 2], a[2, 2],a[1, 3],a[2,3]
- In row major form, the address of a[i1, i2] is
- Base+((i1-low1)*(high2-low2+1)+i2-low2)*width

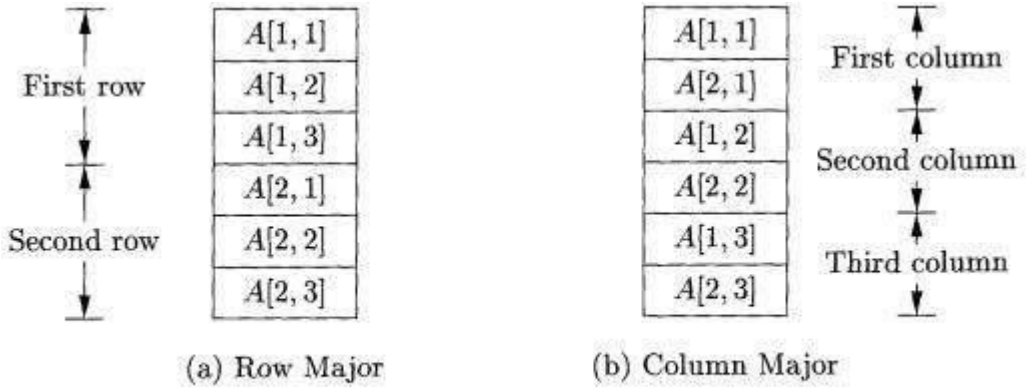


Figure 6.21: Layouts for a two-dimensional array.

4 Translation of Array References

L generate an array name followed by a sequence of index expressions:

$$L \rightarrow L\,[\,E\,] \mid \mathbf{id}\,[\,E\,]$$

Calculate addresses based on widths, using the formula rather than on numbers of elements. The translation scheme generates three-address code for expressions with array references. It consists of the productions and semantic actions together with productions involving nonterminal

$S \rightarrow \mathbf{id} = E\ ;$ { $gen(top.get(\mathbf{id}.lexeme)\ '='\ E.addr);$ }

| $L = E\ ;$ { $gen(L.addr.base\ '['\ L.addr\ ']'\ '='\ E.addr);$ }

$E \rightarrow E_1 + E_2$ { $E.addr = \mathbf{new}\ Temp();$
$gen(E.addr\ '='\ E_1.addr\ '+'\ E_2.addr);$ }

| $\mathbf{id}$ { $E.addr = top.get(\mathbf{id}.lexeme);$ }

| L { $E.addr = \mathbf{new}\ Temp();$
$gen(E.addr\ '='\ L.array.base\ '['\ L.addr\ ']');$ }

$L \rightarrow \mathbf{id}\ [\ E\ ]$ { $L.array = top.get(\mathbf{id}.lexeme);$
$L.type = L.array.type.elem;$
$L.addr = \mathbf{new}\ Temp();$
$gen(L.addr\ '='\ E.addr\ '*'\ L.type.width);$ }

| $L_1\ [\ E\ ]$ { $L.array = L_1.array;$
$L.type = L_1.type.elem;$
$t = \mathbf{new}\ Temp();$
$L.addr = \mathbf{new}\ Temp();$
$gen(t\ '='\ E.addr\ '*'\ L.type.width);$ }
$gen(L.addr\ '='\ L_1.addr\ '+'\ t);$ }

Figure 6.22: Semantic actions for array references

.

Type Checking

1 Rules for Type Checking

2 Type Conversions

3 Overloading of Functions and Operators

4 Type Inference and Polymorphic Functions

5 An Algorithm for Unification

Type checking a compiler needs to assign a type expression to each component of the source program. The compiler must then determine that these type expressions conform to a collection of logical rules that is called the type system for the source language.

1 Rules for Type Checking

Type checking can take on two forms: **synthesis and inference**. *Type synthesis* builds up the type of an expression from the types of its subexpressions. It requires names to be declared before they are used. The type of *E1 + E2* is defined in terms of the types of *E1* and *E2* • A typical rule for type synthesis has the form

$$\begin{array}{l}\textbf{if } f \text{ has type } s \to t \textbf{ and } x \text{ has type } s, \\ \textbf{then} \text{ expression } f(x) \text{ has type } t\end{array} \qquad (6.8)$$

Type inference determines the type of a language construct from the way it is used.Rule for type inference has the form

$$\begin{array}{l}\textbf{if } f(x) \text{ is an expression,} \\ \textbf{then} \text{ for some } \alpha \text{ and } \beta, f \text{ has type } \alpha \to \beta \textbf{ and } x \text{ has type } \alpha\end{array} \qquad (6.9)$$

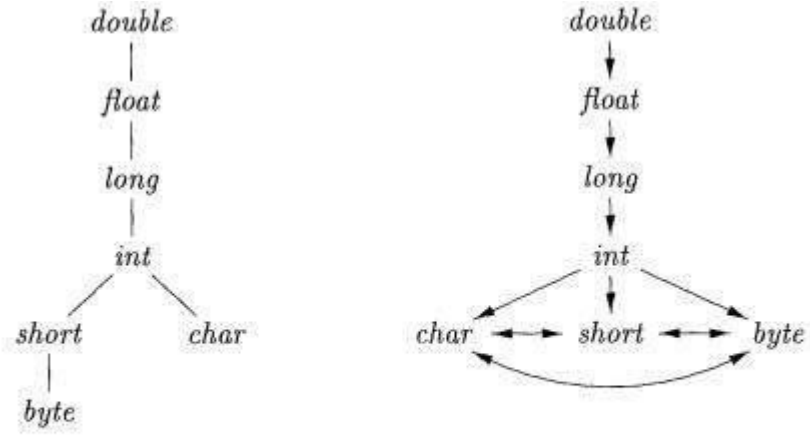
2 Type Conversions

integers are converted to floats when necessary, using a unary operator (f l o a t) . For example, the integer **2** is converted to a float in the code for the expression **2*3.14:**

```
t1 = (float) 2
t2 = t1 * 3.14
```

Type conversion rules vary from language to language. The rules for Java in Fig. 6.25 distinguish between *widening* conversions, which are intended to preserve information, and *narrowing* conversions, which can lose information.

Conversion from one type to another is said to be *implicit* if it is done automatically by the compiler. Implicit type conversions, also called *coercions,*



(a) Widening conversions (b) Narrowing conversions

Figure 6.25: Conversions between primitive types in Java

3 Overloading of Functions and Operators

An *overloaded* symbol has different meanings depending on its context. The + operator in Java denotes either string concatenation or addition

Type-synthesis rule for overloaded functions:

if f can have type $s_i \to t_i$, for $1 \leq i \leq n$, where $s_i \neq s_j$ for $i \neq j$
and x has type s_k, for some $1 \leq k \leq n$ (6.10)
then expression $f(x)$ has type t_k

4 Type Inference and Polymorphic Functions

Type inference is useful for a language like ML, which is strongly typed, but does not require names to be declared before they are used. Type inference ensures that names are used consistently.

The term "polymorphic" refers to any code fragment that can be executed with arguments of different types.

The type of length can be described as, "for any type a, length maps a list of elements of type a to an integer."

```
fun length(x) =
        if null(x) then 0 else length(tl(x)) + 1;
```

Figure 6.28: ML program for the length of a list

The program fragment defines function *length* with one parameter *x*. The body of the function consists of a conditional expression. The predefined function *null* tests whether a list is empty, and the predefined function *tl* (short for "tail") returns the remainder of a list after the first element is removed.

5 An Algorithm for Unification

Unification is the problem of determining whether two expressions s and t can be made identical by substituting expressions for the variables in s and t. Testing equality of expressions is a special case of unification; if s and t have constants but no variables, then s and t unify if and only if they are identical. so it can be used to test structural equivalence of circular types . [7]

Graph-theoretic formulation of unification, where types are represented by graphs. Type variables are represented by leaves and type constructors are represented by interior nodes. Nodes are grouped into equiv-alence classes; if two nodes are in the same equivalence class, then the type expressions they represent must unify. Thus, all interior nodes in the same class must be for the same type constructor, and their corresponding children must be equivalent.

Example 6.18 : Consider the two type expressions

$$((\alpha_1 \to \alpha_2) \times list(\alpha_3)) \to list(\alpha_2)$$
$$((\alpha_3 \to \alpha_4) \times list(\alpha_3)) \to \alpha_5$$

The following substitution S is the most general unifier for these expressions

x	$S(x)$
α_1	α_1
α_2	α_2
α_3	α_1
α_4	α_2
α_5	$list(\alpha_2)$

This substitution maps the two type expressions to the following expression

$$((\alpha_1 \to \alpha_2) \times list(\alpha_1)) \to list(\alpha_2)$$

```
boolean unify(Node m, Node n) {
    s = find(m); t = find(n);
    if ( s = t ) return true;
    else if ( nodes s and t represent the same basic type ) return true;
    else if (s is an op-node with children s1 and s2 and
                t is an op-node with children t1 and t2) {
        union(s, t);
        return unify(s1, t1) and unify(s2, t2);
    }
    else if s or t represents a variable {
        union(s, t);
        return true;
    }
    else return false;
}
```

Figure 6.32: Unification algorithm.

The unification algorithm, uses the following two operations on nodes:

find{n) returns the representative node of the equivalence class currently containing node *n.*

union(m, n) merges the equivalence classes containing nodes *m* and n. If one of the representatives for the equivalence classes of m and n is a non-variable node, *union* makes that nonvariable node be the representative for the merged equivalence class; otherwise, *union* makes one or the other of the original representatives be the new representative. This asymme-try in the specification of *union* is important because a variable cannot be used as the representative for an equivalence class for an expression containing a type constructor or basic type. Otherwise, two inequivalent expressions may be unified through that variable.

Control Flow

1 Boolean Expressions

2 Short-Circuit Code

3 Flow-of-Control Statements

4 Control-Flow Translation of Boolean Expressions

In programming languages, boolean expressions are often used to

1. **Alter the flow of control**. Boolean expressions are used as conditional expressions in statements that alter the flow of control. The value of such boolean expressions is implicit in a position reached in a program. For example, in if *(E)* 5, the expression *E* must be true if statement *S* is reached.

2. **Compute logical values.** A boolean expression can represent *true* Or *false* as values. Such boolean

expressions can be evaluated in analogy to arithmetic expressions using three-address instructions with logical operators.

1 Boolean Expressions

Boolean expressions are composed of the boolean operators (which we denote &&, I I, and !, using the C convention for the operators AND, OR, and NOT, respectively) applied to elements that are boolean variables or relational expressions. Boolean expressions generated by the following grammar:

$$B \rightarrow B \,||\, B \mid B \,\&\&\, B \mid !B \mid (B) \mid E \,\mathbf{rel}\, E \mid \mathbf{true} \mid \mathbf{false}$$

Given the expression Bi I I B2, if we determine that B1 is true, then we can conclude that the entire expression is true without having to evaluate B2.Similarly, given B1 && B2, if Bi is false, then the entire expression is false.

2 Short-Circuit Code

In *short-circuit* (or *jumping)* code, the boolean operators &&, I I, and ! translate into jumps. The operators themselves do not appear in the code; instead, the value of a boolean expression is represented by a position in the code sequence.

Example The statement

i f (x < 100 || x > 200 && x != y) x = 0;

might be translated into the code of Fig. **6.34.** In this translation, the boolean expression is true if control reaches label L_2. If the expression is false, control goes immediately to L_u skipping L_2 and the assignment x = **0.**

```
        if x < 100 goto L2
        ifFalse x > 200 goto L1
        ifFalse x != y goto L1
L2:     x = 0
L1:
```

Figure 6.34: Jumping code

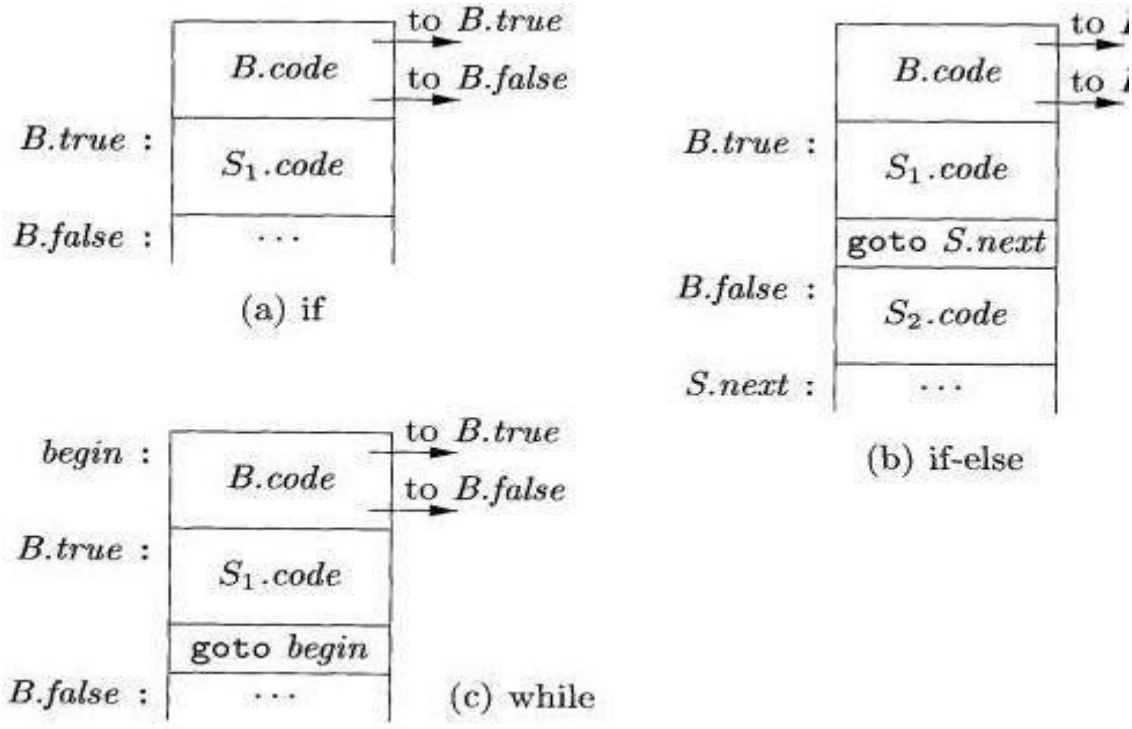
3 Flow-of-Control Statements

$$\begin{aligned} S &\rightarrow \mathbf{if}\ (\ B\ )\ S_1 \\ S &\rightarrow \mathbf{if}\ (\ B\ )\ S_1\ \mathbf{else}\ S_2 \\ S &\rightarrow \mathbf{while}\ (\ B\ )\ S_1 \end{aligned}$$

In these productions, nonterminal *B* represents a boolean expression and non-terminal S represents a statement.

B and S have a synthesized attribute *code*, which gives the translation into three-address instructions. we build up the translations *B.code* and *S.code* as strings, using syntax directed definitions.

The translation of if (B) S1 consists of B.code followed by Si.code, as illustrated in Fig. 6.35(a). Within B.code are jumps based on the value of B. If B is true, control flows to the first instruction of Si.code, and if B is false, control flows to the instruction immediately following S1 . code.



Code for if, if else, while statements

The syntax-directed definition in Fig. 6.36-6.37 produces three-address code for boolean expressions in the context of if-, if-else-, and while-statements.

PRODUCTION	SEMANTIC RULES
$P \rightarrow S$	$S.next = newlabel()$ $P.code = S.code \mid\mid label(S.next)$
$S \rightarrow$ **assign**	$S.code =$ **assign**$.code$
$S \rightarrow$ **if** (B) S_1	$B.true = newlabel()$ $B.false = S_1.next = S.next$ $S.code = B.code \mid\mid label(B.true) \mid\mid S_1.code$
$S \rightarrow$ **if** (B) S_1 **else** S_2	$B.true = newlabel()$ $B.false = newlabel()$ $S_1.next = S_2.next = S.next$ $S.code = B.code$ $\mid\mid label(B.true) \mid\mid S_1.code$ $\mid\mid gen('\texttt{goto}'\ S.next)$ $\mid\mid label(B.false) \mid\mid S_2.code$
$S \rightarrow$ **while** (B) S_1	$begin = newlabel()$ $B.true = newlabel()$ $B.false = S.next$ $S_1.next = begin$ $S.code = label(begin) \mid\mid B.code$ $\mid\mid label(B.true) \mid\mid S_1.code$ $\mid\mid gen('\texttt{goto}'\ begin)$
$S \rightarrow S_1\ S_2$	$S_1.next = newlabel()$ $S_2.next = S.next$ $S.code = S_1.code \mid\mid label(S_1.next) \mid\mid S_2.code$

Figure 6.36: Syntax-directed definition for flow-of-control statements.

4 Control-Flow Translation of Boolean Expressions

Boolean expression *B* is translated into three-address instructions that evaluate *B* using creates labels only when they are needed. Alternatively, unnecessary labels can be eliminated during a subsequent optimization phase.

PRODUCTION	SEMANTIC RULES
$B \rightarrow B_1 \mid\mid B_2$	$B_1.true = B.true$ $B_1.false = newlabel()$ $B_2.true = B.true$ $B_2.false = B.false$ $B.code = B_1.code \mid\mid label(B_1.false) \mid\mid B_2.code$
$B \rightarrow B_1 \,\&\&\, B_2$	$B_1.true = newlabel()$ $B_1.false = B.false$ $B_2.true = B.true$ $B_2.false = B.false$ $B.code = B_1.code \mid\mid label(B_1.true) \mid\mid B_2.code$
$B \rightarrow ! B_1$	$B_1.true = B.false$ $B_1.false = B.true$ $B.code = B_1.code$
$B \rightarrow E_1$ **rel** E_2	$B.code = E_1.code \mid\mid E_2.code$ $\mid\mid gen('\texttt{if}'\ E_1.addr\ \textbf{rel}.op\ E_2.addr\ '\texttt{goto}'\ B.true)$ $\mid\mid gen('\texttt{goto}'\ B.false)$
$B \rightarrow$ **true**	$B.code = gen('\texttt{goto}'\ B.true)$
$B \rightarrow$ **false**	$B.code = gen('\texttt{goto}'\ B.false)$

Figure 6.37: Generating three-address code for booleans

Backpatching

1 One-Pass Code Generation Using Backpatching

2 Backpatching for Boolean Expressions

3 Flow-of-Control Statements

1 One-Pass Code Generation Using Backpatching

The problem in generating three address codes in a single pass is that we may not know the labels that control must go to at the time jump statements are generated. So to get around this

problem a series of branching statements with the targets of the jumps temporarily left unspecified is generated. Back Patching is putting the address instead of labels when the proper label is determined.

To manipulate lists of jumps, Back patching Algorithms perform three types of operations

1.*makelist(i)* creates a new list containing only *i*, an index into the array of instructions; *makelist* returns a pointer to the newly created list.
2. merge(pi,p2) concatenates the lists pointed to by p1 and p2, and returns a pointer to the concatenated list.
3. backpatch(p,i) inserts i as the target label for each of the instructions on the list pointed to by p.

2 Backpatching for Boolean Expressions

Construct a translation scheme suitable for generating code for boolean expressions during bottom-up parsing. A marker nonterminal *M* in the grammar causes a semantic action to pick up, at appropriate times, the index of the next instruction to be generated. The grammar is as follows:

$B \rightarrow B_1 \ || \ M \ B_2 \mid B_1 \ \&\& \ M \ B_2 \mid ! \ B_1 \mid (\ B_1 \ ) \mid E_1 \ \mathbf{rel} \ E_2 \mid \mathbf{true} \mid \mathbf{false}$
$M \rightarrow \epsilon$

The translation scheme is in Fig. 6.43.

1) $B \rightarrow B_1 \ || \ M \ B_2$ { $backpatch(B_1.falselist, M.instr)$;
$B.truelist = merge(B_1.truelist, B_2.truelist)$;
$B.falselist = B_2.falselist$; }

2) $B \rightarrow B_1 \ \&\& \ M \ B_2$ { $backpatch(B_1.truelist, M.instr)$;
$B.truelist = B_2.truelist$;
$B.falselist = merge(B_1.falselist, B_2.falselist)$; }

3) $B \rightarrow \ ! \ B_1$ { $B.truelist = B_1.falselist$;
$B.falselist = B_1.truelist$; }

4) $B \rightarrow (\ B_1 \ )$ { $B.truelist = B_1.truelist$;
$B.falselist \ = \ B_1.falselist$; }

5) $B \rightarrow E_1 \ \mathbf{rel} \ E_2$ { $B.truelist = makelist(nextinstr)$;
$B.falselist = makelist(nextinstr + 1)$;
$emit('\texttt{if}' \ E_1.addr \ \mathbf{rel}.op \ E_2.addr \ '\texttt{goto}\ \_')$;
$emit('\texttt{goto}\ \_')$; }

6) $B \rightarrow \mathbf{true}$ { $B.truelist = makelist(nextinstr)$;
$emit('\texttt{goto}\ \_')$; }

7) $B \rightarrow \mathbf{false}$ { $B.falselist = makelist(nextinstr)$;
$emit('\texttt{goto}\ \_')$; }

8) $M \rightarrow \epsilon$ { $M.instr = nextinstr$; }

Figure 6.43: Translation scheme for boolean expressions

Consider semantic action (1) for the production B -> B1 || M B2. If Bx is true, then B is also true, so the jumps on Bi.truelist become part of B.truelist. If Bi is false, however, we must next test B2, so the target for the jumps B>i .falselist must be the beginning of the code generated for B2 • This target is obtained using the marker nonterminal M. That nonterminal produces, as a synthesized attribute M.instr, the index of the next instruction, just before B2 code starts being generated.

Example

Consider expression

$$x < 100 \;||\; x > 200 \;\&\&\; x \;!= \; y$$

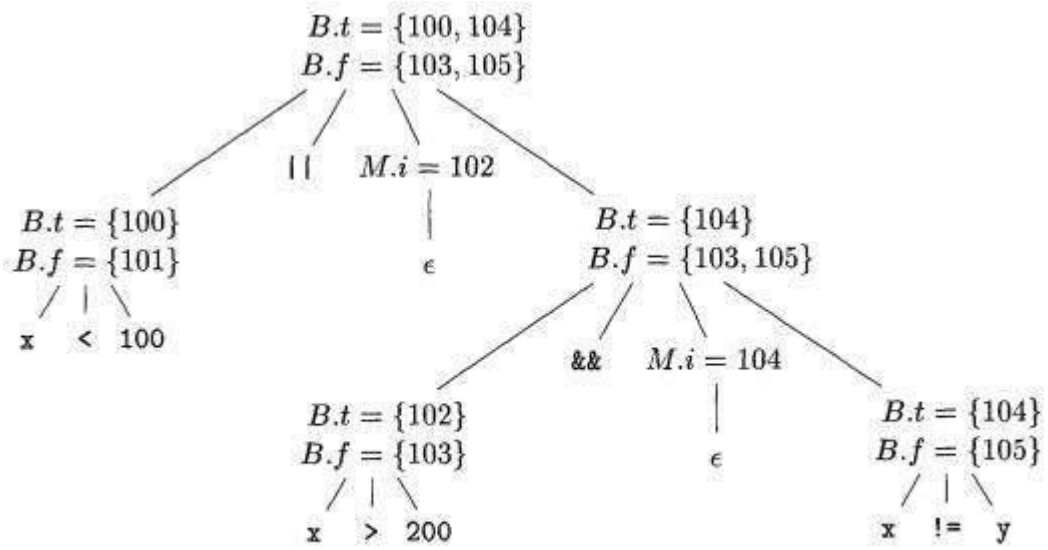


Figure 6.44: Annotated parse tree for $x < 100 \;||\; x > 200 \;\&\&\; x \;!= \; y$

An annotated parse tree is shown in Fig. 6.44; attributes *truelist*, *falselist*, and *instr* are represented by their initial letters. The actions are performed during a depth-first traversal of the tree. Since all actions appear at the ends of right sides, they can be performed in conjunction with reductions during a bottom-up parse. In response to the reduction of *x* < 100 to *B* by production (5), the two instructions

```
100:  if x < 100 goto _
101:  goto _
```

are generated. (start instruction numbers at 100.) The marker nonterminal *M* in the production

$$B \rightarrow B_1 \;||\; M \; B_2$$

records the value of nextinstr, which at this time is 102. The reduction of x > 200 to B by production (5) generates the instructions

```
102:  if x > 200 goto _
103:  goto _
```

The subexpression $x > 200$ corresponds to B_1 in the production

$$B \rightarrow B_1 \;\&\&\; M\; B_2$$

The marker nonterminal *M* records the current value of *nextinstr,* which is now

Reducing *x ! = y* into *B* by production (5) generates

```
104:  if x != y goto _
105:  goto _
```

We now reduce by B —> B1 && M B2. The corresponding semantic action calls backpatch(B1.truelist,M.instr) to bind the true exit of B1 to the first instruction of B2. Since B1.truelist is {102} and M.instr is 104, this call to backpatch fills in 104 in instruction 102. The six instructions generated so far are thus as shown in Fig. 6.45(a).

The semantic action associated with the final reduction by B —> B1 || MB2 calls backpatch({101},102) which leaves the instructions as in Fig

The entire expression is true if and only if the gotos of instructions 100 or 104 are reached, and is false if and only if the gotos of instructions 103 or 105 are reached. These instructions will have their targets filled in later in the compilation, when it is seen what must be done depending on the truth or falsehood

```
100:  if x < 100 goto _
101:  goto _
102:  if x > 200 goto 104
103:  goto _
104:  if x != y goto _
105:  goto _
```

(a) After backpatching 104 into instruction 102.

```
100:  if x < 100 goto _
101:  goto 102
102:  if y > 200 goto 104
103:  goto _
104:  if x != y goto _
105:  goto _
```

(b) After backpatching 102 into instruction 101.

Figure 6.45: Steps in the backpatch process

of the expression. as

3 Flow-of-Control Statements

use backpatching to translate flow-of-control statements in one pass.

$$S \rightarrow \textbf{if}\,(B)\,S \mid \textbf{if}\,(B)\,S\,\textbf{else}\,S \mid \textbf{while}\,(B)\,S \mid \{\,L\,\} \mid A\;;$$
$$L \rightarrow L\,S \mid S$$

The translation scheme in Fig. 6.46 maintains lists of jumps that are filled in when their targets are found.

1) $S \rightarrow \textbf{if}\,(\,B\,)\,M\,S_1$ { *backpatch*($B.truelist$, $M.instr$);
   $S.nextlist = merge(B.falselist, S_1.nextlist)$; }

2) $S \rightarrow \textbf{if}\,(\,B\,)\,M_1\,S_1\,N\,\textbf{else}\,M_2\,S_2$
   { *backpatch*($B.truelist$, $M_1.instr$);
   backpatch($B.falselist$, $M_2.instr$);
   $temp = merge(S_1.nextlist, N.nextlist)$;
   $S.nextlist = merge(temp, S_2.nextlist)$; }

3) $S \rightarrow \textbf{while}\,M_1\,(\,B\,)\,M_2\,S_1$
   { *backpatch*($S_1.nextlist$, $M_1.instr$);
   backpatch($B.truelist$, $M_2.instr$);
   $S.nextlist = B.falselist$;
   emit('goto' $M_1.instr$); }

4) $S \rightarrow \{\,L\,\}$ { $S.nextlist = L.nextlist$; }

5) $S \rightarrow A\;;$ { $S.nextlist =$ **null**; }

6) $M \rightarrow \epsilon$ { $M.instr = nextinstr$; }

7) $N \rightarrow \epsilon$ { $N.nextlist = makelist(nextinstr)$;
   emit('goto _'); }

8) $L \rightarrow L_1\,M\,S$ { *backpatch*($L_1.nextlist$, $M.instr$);
   $L.nextlist = S.nextlist$; }

9) $L \rightarrow S$ { $L.nextlist = S.nextlist$; }

Figure 6.46: Translation of statements

Backpatch the jumps when B is true to the instruction Mi.instr; the latter is the beginning of the code for Si. Similarly, we backpatch jumps when B is false to go to the beginning of the code for S 2 . The list S.nextlist includes all jumps out of Si and S 2 , as well as the jump generated by N. (Variable temp is a temporary that is used only for merging lists.

Semantic actions (8) and (9) handle sequences of statements. In

$$L \rightarrow L_1 \, M \, S$$

the instruction following the code for L± in order of execution is the beginning of S. Thus the L1.nextlist list is backpatched to the beginning of the code for S, which is given by M.instr. In L —> S, L.nextlist is the same as S.nextlist.

Note that no new instructions are generated anywhere in these semantic rules, except for rules (3) and (7). All other code is generated by the semantic actions associated with assignment-statements and expressions. The flow of control causes the proper backpatching so that the assignments and boolean expression evaluations will connect properly.

UNIT – V

Runtime Environments, Stack allocation of space, access to Non Local date on the stack Heap Management code generation–Issues in design of code generation the target Language Address in the target code Basic blocks and Flow graphs. A Simple Code generation.

UNIT 5
RUNTIME ENVIRONMENT

By **runtime**, we mean a program in execution. **Runtime environment** is a state of the target machine, which may include software libraries, **environment** variables, etc., to provide services to the processes running in the system.

Storage Organization

- When the target program executes then it runs in its own logical address space in which the value of each program has a location.
- The logical address space is shared among the compiler, operating system and target machine for management and organization. The operating system is used to map the logical address into physical address which is usually spread throughout the memory.

The run-time representation of an object program in the logical address space consists of data and program areas as shown in Fig. 5.1

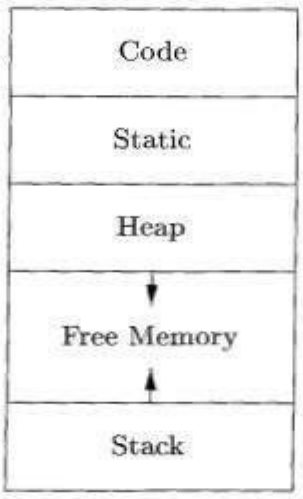


Figure 7.1: Typical subdivision of run-time memory into code and data areas

Storage needed for a name is determined from its type.

- Runtime storage comes into blocks, where a byte is used to show the smallest unit of addressable memory. Using the four bytes a machine word can form. Object of multibyte is stored in consecutive bytes and gives the first byte address.
- Run-time storage can be subdivide to hold the different components of an executing program:

1. Generated executable code
2. Static data objects
3. Dynamic data-object- heap
4. Automatic data objects- stack

Two areas, *Stack* and *Heap,* are at the opposite ends of the remainder of the address space. These areas are dynamic; their size can change as the program executes. Stack to support call/return policy for procedures.Heap to store data that can outlive a call to a procedure. . The heap is used to manage allocate and deallocate data.

Static Versus Dynamic Storage Allocation

The layout and allocation of data to memory locations in the run-time environment are key issues in storage management. The two terms *static* and *dynamic* distinguish between compile time and run time, respectively. We say that a storage-allocation decision is

Static:- if it can be made by the compiler looking only at the text of the program, not at what the program does when it executes.
Dynamic:- if it can be decided only while the program is running.

Compilers use following two strategies for dynamic storage allocation:

Stack storage. Names local to a procedure are allocated space on a stack. stack supports the normal call/return policy for procedures.

Heap storage. Data that may outlive the call to the procedure that created it is usually allocated on a "heap" of reusable storage.The heap is an area of virtual memory that allows objects or other data elements to obtain storage when they are created and to return that storage when they are invalidated.

Stack allocation of space

1 Activation Trees

2 Activation Records

3 Calling Sequences

4 Variable-Length Data on the Stack

Each time a procedure is called, space for its local variables is pushed onto a stack, and when the procedure terminates, that space is popped off the stack.

1 Activation Trees

Stack allocation is a valid allocation for procedures since procedure calls are nested

Example: quicksort algorithm

The main function has three tasks. It calls *readArray,* sets the sentinels, and then calls *quicksort* on the entire data array.

Procedure activations are nested in time. If an activation of procedure *p* calls procedure *q,* then that activation of *q* must end before the activation of *p* can end.

```
int a[11];
void readArray() { /* Reads 9 integers into a[1], ..., a[9]. */
    int i;
    ...
}
int partition(int m, int n) {
    /* Picks a separator value v, and partitions a[m..n] so that
       a[m..p-1] are less than v, a[p] = v, and a[p+1..n] are
       equal to or greater than v. Returns p. */
    ...
}
void quicksort(int m, int n) {
    int i;
    if (n > m) {
        i = partition(m, n);
        quicksort(m, i-1);
        quicksort(i+1, n);
    }
}
main() {
    readArray();
    a[0] = -9999;
    a[10] = 9999;
    quicksort(1,9);
}
```

Figure 7.2: Sketch of a quicksort program

Represent the activations of procedures during the running of an entire program by a tree, called an activation tree. Each node corresponds to one activation, and the root is the activation of the "main" procedure that initiates execution of the program. At a node for an activation of procedure p, the children correspond to activations of the procedures called by this activation of p.

```
enter main()
    enter readArray()
    leave readArray()
    enter quicksort(1,9)
        enter partition(1,9)
        leave partition(1,9)
        enter quicksort(1,3)
            ...
        leave quicksort(1,3)
        enter quicksort(5,9)
            ...
        leave quicksort(5,9)
    leave quicksort(1,9)
leave main()
```

Figure 7.3: Possible activations for the program of Fig. 7.2

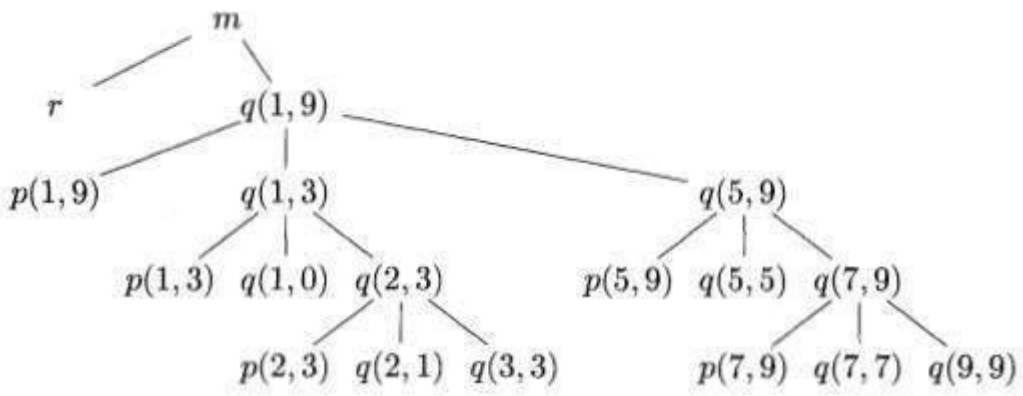


Figure 7.4: Activation tree representing calls during an execution of *quicksort*

2 Activation Records

a. Procedure calls and returns are usually managed by a run-time stack called the control stack.
b. Each live activation has an activation record (sometimes called a frame)
c. The root of activation tree is at the bottom of the stack
d. The current execution path specifies the content of the stack with the last
e. Activation has record in the top of the stack.

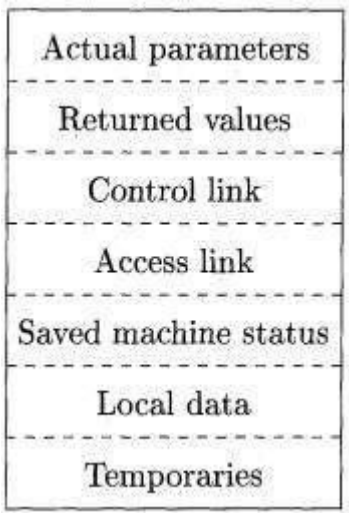


Figure 7.5: A general activation record

An activation record is used to store information about the status of the machine, such as the value of the program counter and machine registers, when a procedure call occurs. When control returns from the call, the activation of the calling procedure can be restarted after restoring the values of relevant registers and setting the program counter to the point immediately after the call. Data objects whose lifetimes are contained in that of an activation can be allocated on the stack along with other information associated with the activation.

An activation record contains all the necessary information required to call a procedure. An activation record may contain the following units (depending upon the source language used).

Temporaries	Stores temporary and intermediate values of an expression.
Local Data	Stores local data of the called procedure.
Machine Status	Stores machine status such as Registers, Program Counter etc., before the procedure is called.
Control Link	Stores the address of activation record of the caller procedure.
Access Link	Stores the information of data which is outside the local scope.
Actual Parameters	Stores actual parameters, i.e., parameters which are used to send input to the called procedure.

Return Value	Stores return values.

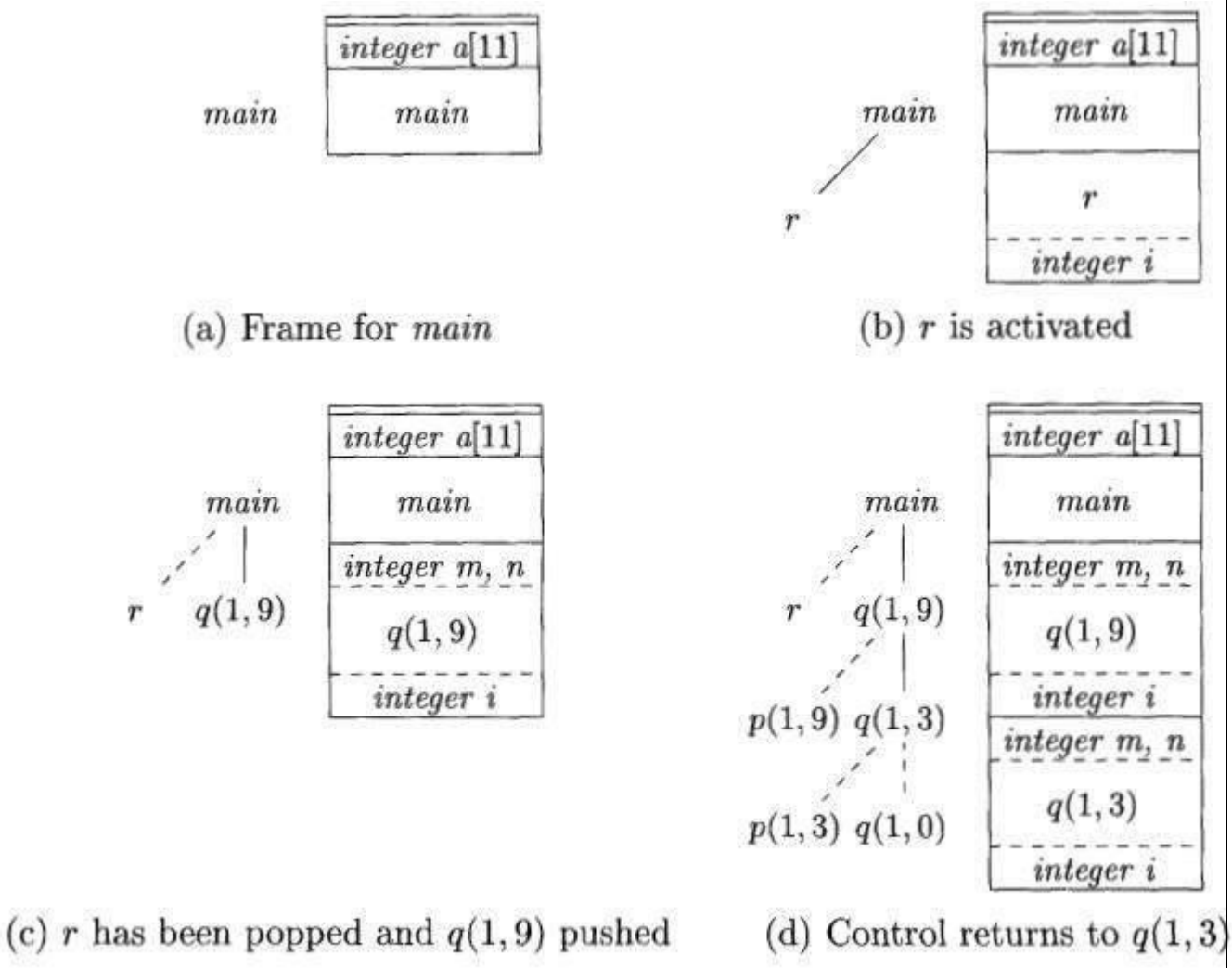


Figure 7.6: Downward-growing stack of activation records

3 Calling Sequences

Designing calling sequences and the layout of activation records, the following

1. Values communicated between caller and callee are generally placed at the beginning of callee's activation record
2. Fixed-length items: are generally placed at the middle. such items typically include the control link, the access link, and the machine status fields.
3. Items whose size may not be known early enough: are placed at the end of activation record
4. We must locate the top-of-stack pointer judiciously: a common approach is to have it point to the end of fixed length fields in the activation record.

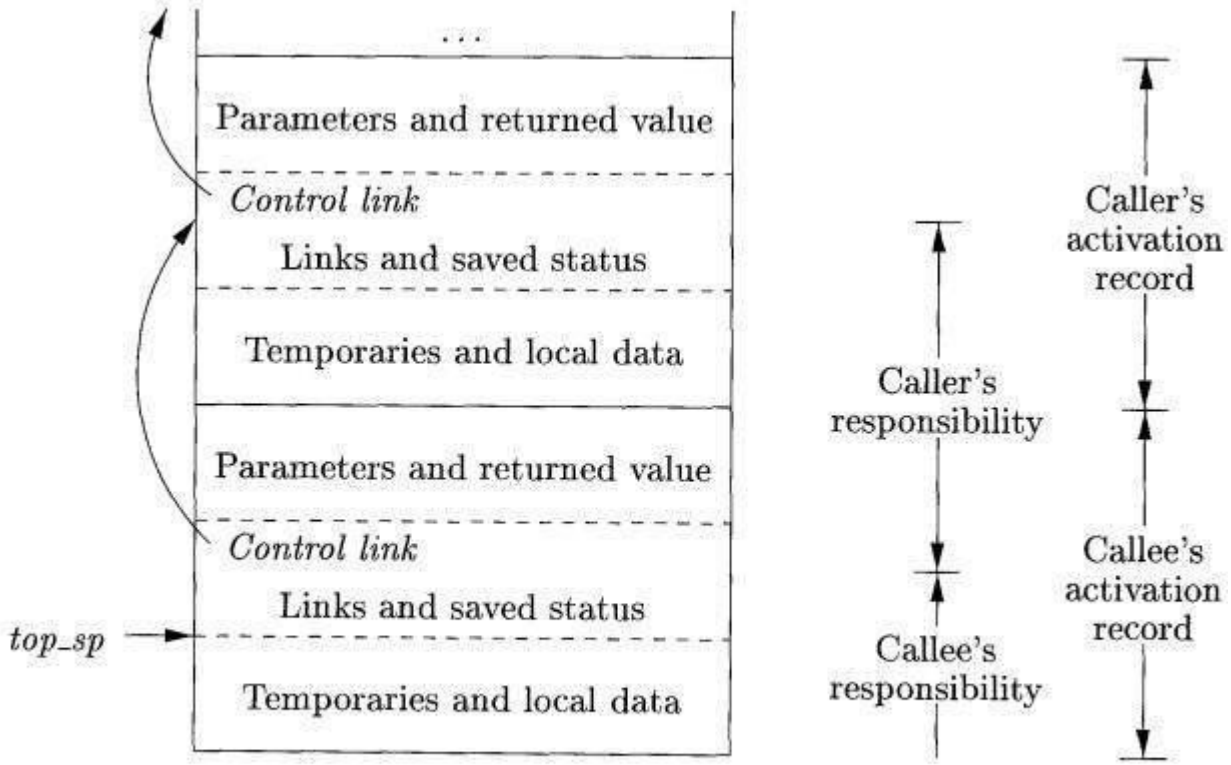


Figure 7.7: Division of tasks between caller and callee

A register topsp points to the end of the machine-status field in the current top activation record. This position within the callee's activation record is known to the caller, so the caller can be made responsible for setting topsp before control is passed to the callee. The calling sequence and its division between caller and callee is as follows:

1. The caller evaluates the actual parameters.

The caller stores a return address and the old value of *topsp* into the callee's activation record. The caller then increments *topsp* to the position shown in Fig. 7.7. That is, *topsp* is moved past the caller's local data and temporaries and the callee's parameters and status fields.

The callee saves the register values and other status information.

The callee initializes its local data and begins execution.

A suitable, corresponding **return sequence** is:

1. The callee places the return value next to the parameters, as in Fig. 7.5.

2. Using information in the machine-status field, the callee restores *topsp* and other registers, and then branches to the return address that the caller placed in the status field.

3. Although *topsp* has been decremented, the caller knows where the return value is, relative to

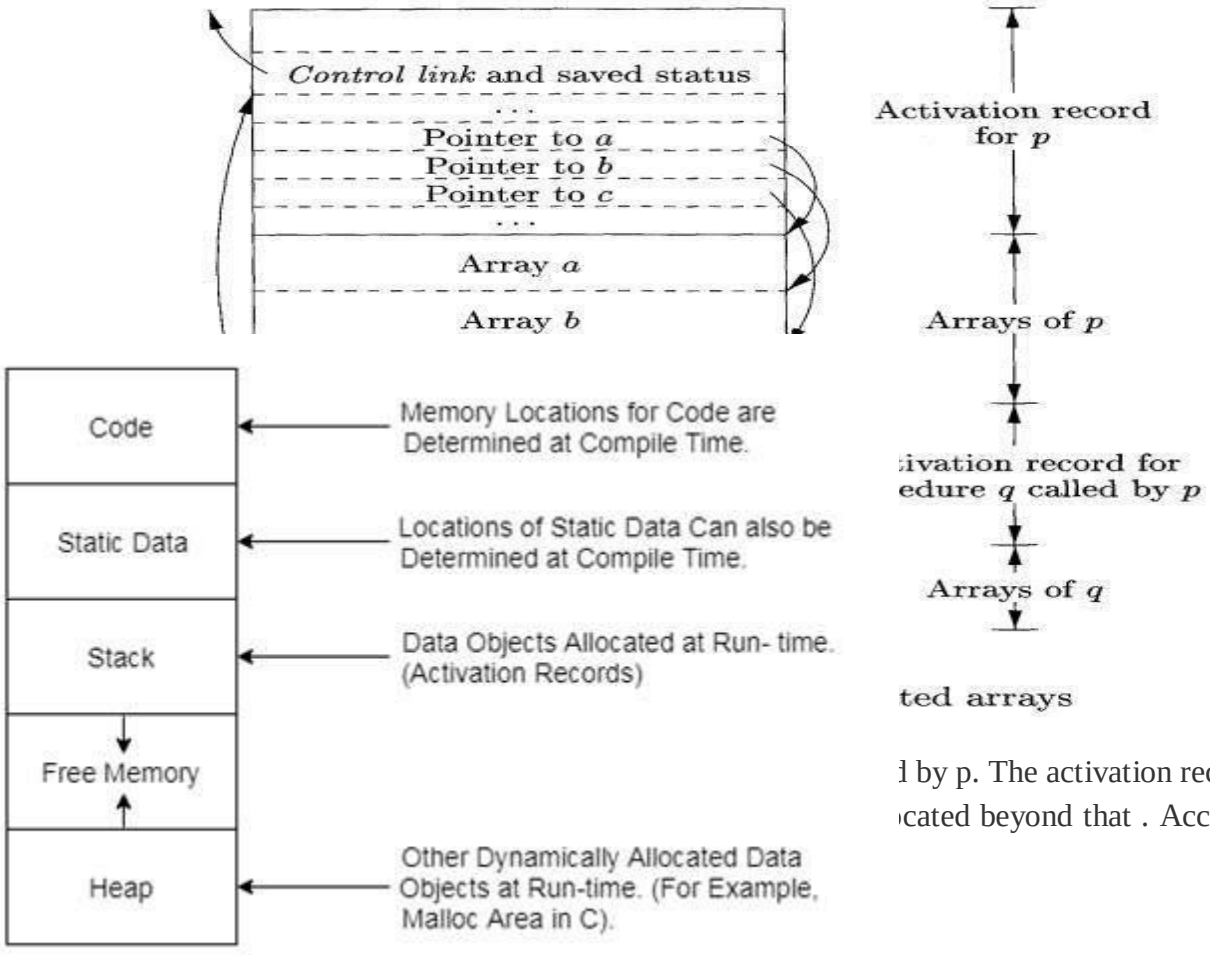
the current value of *topsp;* the caller therefore may use that value.

4. Variable-Length Data on the Stack

The run-time memory-management system must deal frequently with the allocation of space for objects the sizes of which are not known at compile time, but which are local to a procedure and thus may be allocated on the stack.

it is possible to allocate objects, arrays, or other structures of unknown size on the stack. The reason to prefer placing objects on the stack if possible is that we avoid the expense of garbage collecting their space. Note that the stack can be used only for an object if it is local to a procedure and becomes inaccessible when the procedure returns.

A common strategy for allocating variable-length arrays (i.e., arrays whose size depends on the value of one or more parameters of the called procedure) is shown in Fig. 7.8. The same scheme works for objects of any type if they are local to the procedure called and have a size that depends on the parameters of the call.



l by p. The activation record for q
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Access to Non Local data on the stack

1 Data Access Without Nested Procedures

2 Issues With Nested Procedures

3 A Language With Nested Procedure Declarations

4 Nesting Depth

5 Access Links

6 Manipulating Access Links

7 Access Links for Procedure Parameters

8 Displays

Consider how procedures access their data. Especially im-portant is the mechanism for finding data used within a procedure *p* but that does not belong to *p*

1 Data Access Without Nested Procedures

Names are either local to the procedure in question or are declared globally.

1. For global names the address is known statically at compile time providing there is only one source file. If multiple source files, the linker knows. In either case no reference to the activation record is needed; the addresses are know prior to execution.
2. For names local to the current procedure, the address needed is in the AR at a known-at-compile-time constant offset from the sp. In the case of variable size arrays, the constant offset refers to a pointer to the actual storage.

2 Issues With Nested Procedures

Access becomes far more complicated when a language allows procedure dec-larations to be nested .The reason is that knowing at compile time that the declaration of *p* is immediately nested within *q* does not tell us the relative positions of their activation records at run time. In fact, since either *p* or *q* or both may be recursive, there may be several activation records of *p* and/or *q* on the stack.

Finding the declaration that applies to a nonlocal name *x in a* nested pro-cedure *p* is a static decision; it can be done by an extension of the static-scope rule for blocks. Suppose *x* is declared in the enclosing procedure *q*. Finding the relevant activation of *q* from an activation of *p* is a dynamic decision; it re-quires additional run-time information about activations. One possible solution is to use access links.

3. A Language With Nested Procedure Declarations

In various languages with nested procedures, one of the most influential is ML.

ML is a *functional language*, meaning that variables, once declared and initialized, are not changed. There are only a few exceptions, such as the array, whose elements can be changed by special function calls.

• Variables are defined, and have their unchangeable values initialized,

v a l (name) = (expression)

• Functions are defined using the syntax:

fun (name) ((arguments)) = (body)

• For function bodies, use let-statements of the form:

let (list of definitions) in (statements) end The definitions are normally v a l or fun statements. The scope of each such definition consists of all following definitions, up to the in, and all the statements up to the end. Most importantly, function definitions can be nested. For example, the body of a function p can contain a let-statement that includes the definition of another (nested) function q. Similarly, q can have function definitions within its own body, leading to arbitrarily deep nesting of function

4. Nesting Depth

Nesting depth is 1 to procedures that are not nested within any other procedure. For example, all C functions are at nesting depth 1. However, if a procedure *p* is defined immediately within a procedure at nesting depth *i*, then give *p* the nesting depth *i*

```
1) fun sort(inputFile, outputFile) =
       let
2)         val a = array(11,0);
3)         fun readArray(inputFile) = ··· ;
4)               ··· a ··· ;
5)         fun exchange(i,j) =
6)               ··· a ··· ;
7)         fun quicksort(m,n) =
               let
8)                 val v = ··· ;
9)                 fun partition(y,z) =
10)                      ··· a ··· v ··· exchange ···
               in
11)                ··· a ··· v ··· partition ··· quicksort
               end
       in
12)        ··· a ··· readArray ··· quicksort ···
       end;
```

Figure 7.10: A version of quicksort, in ML style, using nested functions

+1

5. Access Links

A direct implementation of the normal static scope rule for nested functions is obtained by adding a pointer called the *access link* to each activation record. If procedure *p* is nested immediately within procedure *q* in the source code, then the access link in any activation of *p* points to the most recent activation of *q.* Note that the nesting depth of *q* must be exactly one less than the nesting depth of p. Access links form a chain from the activation record at the top of the stack to a sequence of activations at progressively lower nesting depths.

Figure 7.11 shows a sequence of stacks that might result from execution of the function *sort* of Fig. 7.10. In Fig. 7.11(a), we see the situation after *sort* has called *readArray* to load input into the array *a* and then called *quicksort(l,* 9) to sort the array. The access link from *quicksort(l,* 9) points to the activation record for *sort,* not because *sort* called *quicksort* but because *sort* is the most closely nested function surrounding *quicksort* in the program.

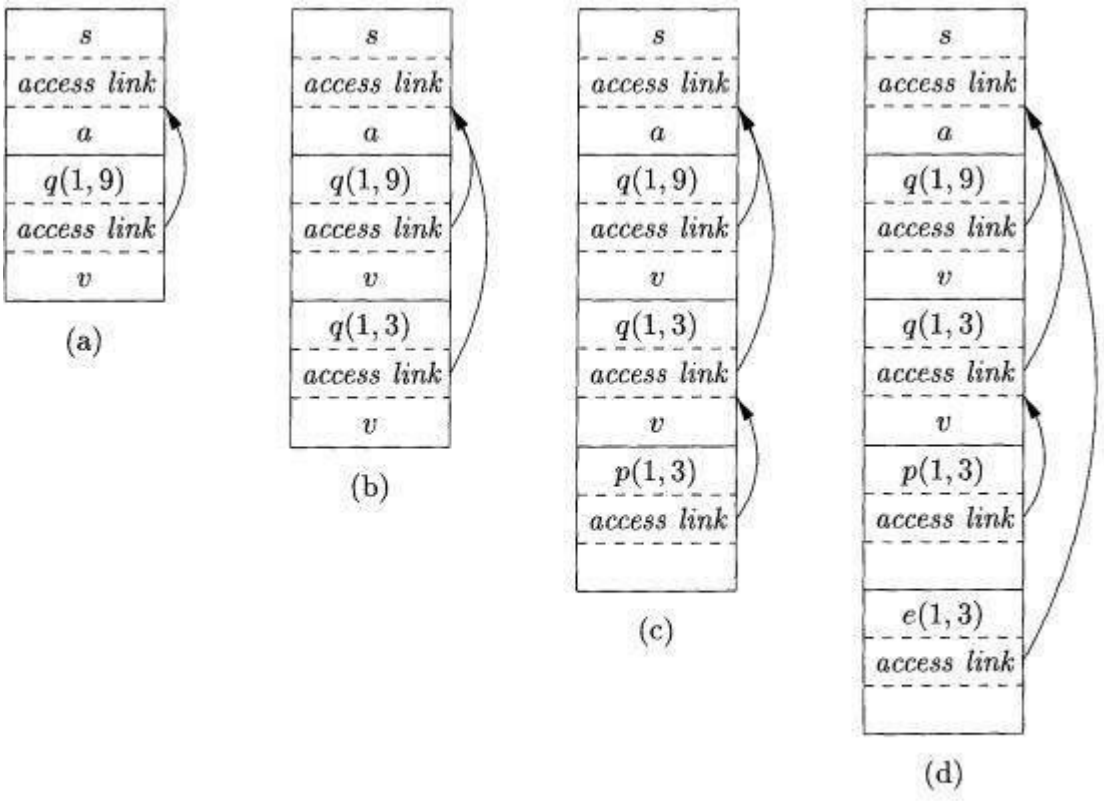


Figure 7.11: Access links for finding nonlocal data

see a recursive call to *quicksort(l,* 3), followed by a call to *partition,* which calls *exchange.* Notice that *quicksort(l,* 3)'s access link points to *sort,* for the same reason that *quicksort(l,* 9)'s does.

6. Manipulating Access Links

The harder case is when the call is to a procedure-parameter; in that case, the particular procedure being called is not known until run time, and the nesting depth of the called procedure may

differ in different executions of the call. consider situation when a procedure *q* calls procedure *p,* explicitly. There are three cases:

1. Procedure *p* is at a higher nesting depth than *q.* Then p must be defined immediately within *q,* or the call by *q* would not be at a position that is within the scope of the procedure name *p.* Thus, the nesting depth of *p* is exactly one greater than that of *q,* and the access link from *p* must lead to *q.* It is a simple matter for the calling sequence to include a step that places in the access link for *p* a pointer to the activation record of *q.*

2. The call is recursive, that is, *p* = *q.*Then the access link for the new activation record is the same as that of the activation record below it.

3. The nesting depth np of p is less than the nesting depth nq of q. In order for the call within q to be in the scope of name p, procedure q must be nested within some procedure r, while p is a procedure defined immediately within r. The top activation record for r can therefore be found by following the chain of access links, starting in the activation record for q, for nq — np + 1 hops. Then, the access link for p must go to this activation of r.

7. Access Links for Procedure Parameters

When a procedure *p* is passed to another procedure *q* as a parameter, and *q* then calls its parameter (and therefore calls *p* in this activation of *q),* it is possible that *q* does not know the context in which *p* appears in the program. If so, it is impossible for *q* to know how to set the access link for *p.* The solution to this is, when procedures are used as parameters, the caller needs to pass, along with the name of the procedure-parameter, the proper access link for that parameter. The caller always knows the link, since if *p* is passed by procedure *r* as an actual parameter, then *p* must be a name accessible to r, and therefore, r can determine the access link for *p* exactly as if *p* were being called by *r* directly.

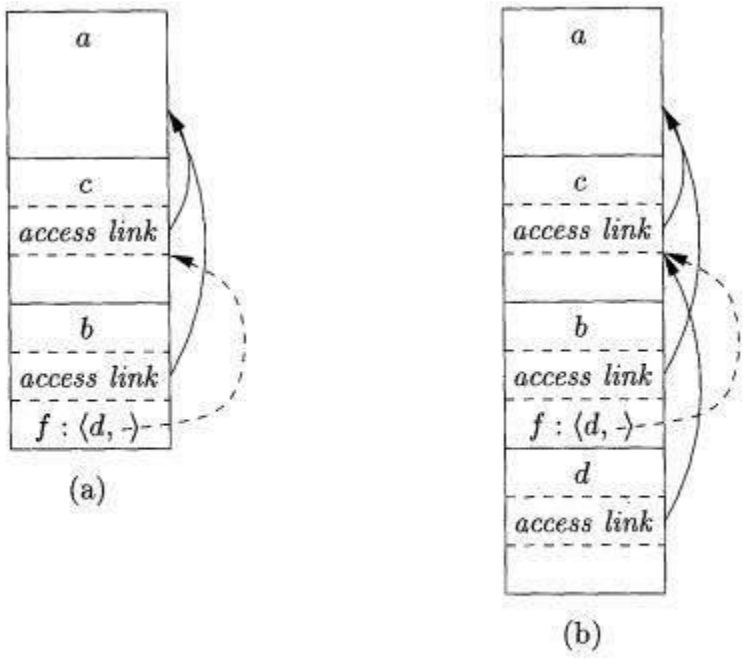


Figure 7.13: Actual parameters carry their access link with them

8. Displays

One problem with the access-link approach to nonlocal data is that if the nesting depth gets large, we may have to follow long chains of links to reach the data we need. A more efficient implementation uses an auxiliary array *d*, called the *display*, which consists of one pointer for each nesting depth. We arrange that, at all times, *d[i]* is a pointer to the highest activation record on the stack for any procedure at nesting depth *i*. Examples of a display are shown in Fig. 7.14.

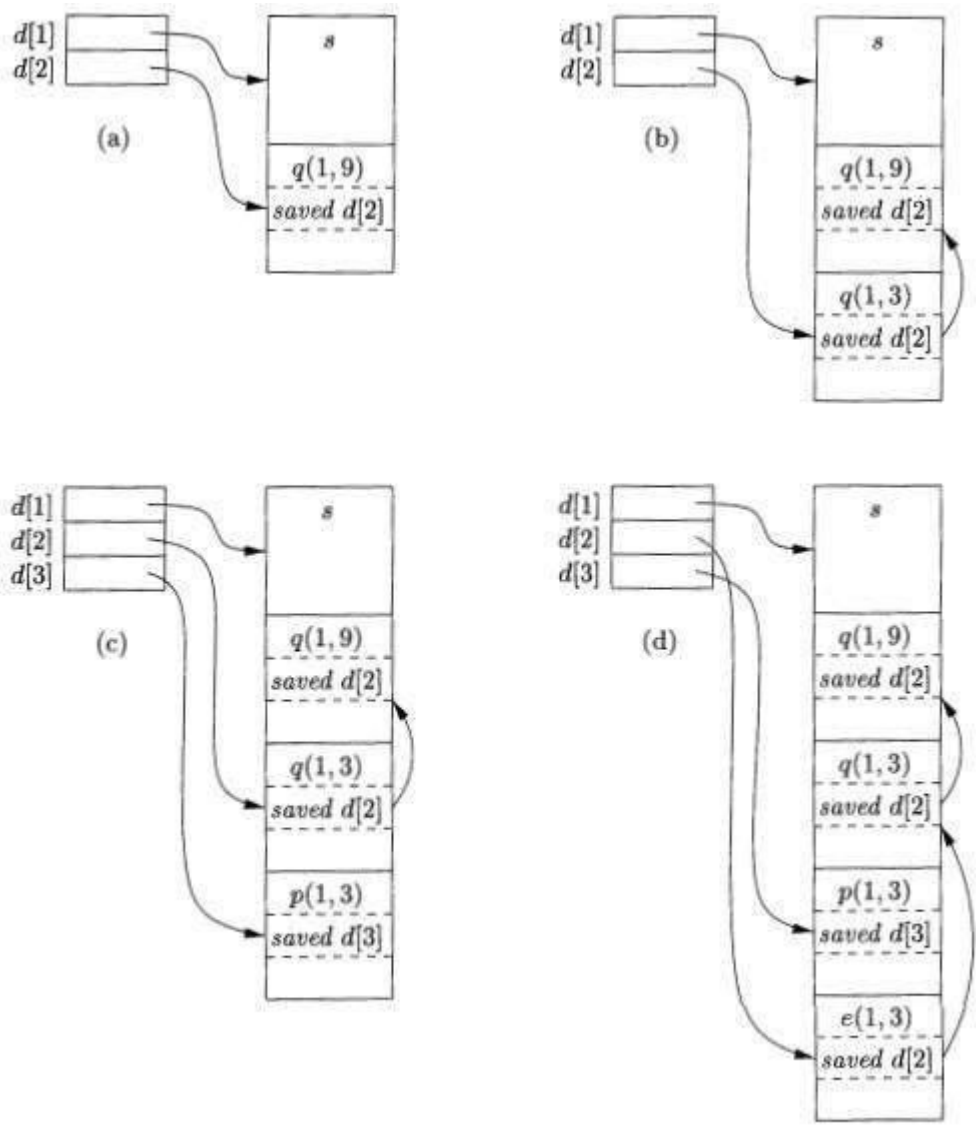


Figure 7.14: Maintaining the display

In order to maintain the display correctly, we need to save previous values of display entries in new activation records.

Heap Management

The heap is the portion of the store that is used for data that lives indefinitely, or until the program explicitly deletes it.

1 The Memory Manager

2 The Memory Hierarchy of a Computer

3 Locality in Programs

4 Reducing Fragmentation

5 Manual Deallocation Requests

1 The Memory Manager

It performs two basic functions:

• **Allocation**. When a program requests memory for a variable or object,[3] the memory manager produces a chunk of contiguous heap memory of the requested size. If possible, it satisfies an allocation request using free space in the heap; if no chunk of the needed size is available, it seeks to increase the heap storage space by getting consecutive bytes of virtual memory from the operating system. If space is exhausted, the memory manager passes that information back to the application program.

• **Deallocation**. The memory manager returns deallocated space to the pool of free space, so it can reuse the space to satisfy other allocation requests. Memory managers typically do not return memory to the operating sys-tem, even if the program's heap usage drops.

Thus, the memory manager must be prepared to service, in any order, allo-cation and deallocation requests of any size, ranging from one byte to as large as the program's entire address space.

Here are the properties we desire of memory managers:

• **Space Efficiency**. A memory manager should minimize the total heap space needed by a program. Larger programs to run in a fixed virtual address space..

• **Program Efficiency.** A memory manager should make good use of the memory subsystem to allow programs to run faster.

• **Low Overhead**. Because memory allocations and deallocations are fre-quent operations in many programs, it is important that these operations be as efficient as possible. That is, we wish to minimize the *overhead*

2. The Memory Hierarchy of a Computer

The efficiency of a program is determined not just by the number of instructions executed, but also by how long it takes to execute each of these instructions. The time taken to execute an instruction can vary significantly, since the time taken to access different parts of memory can vary from

nanoseconds to milliseconds. Data-intensive programs can therefore benefit significantly from optimizations that make good use of the memory subsystem.

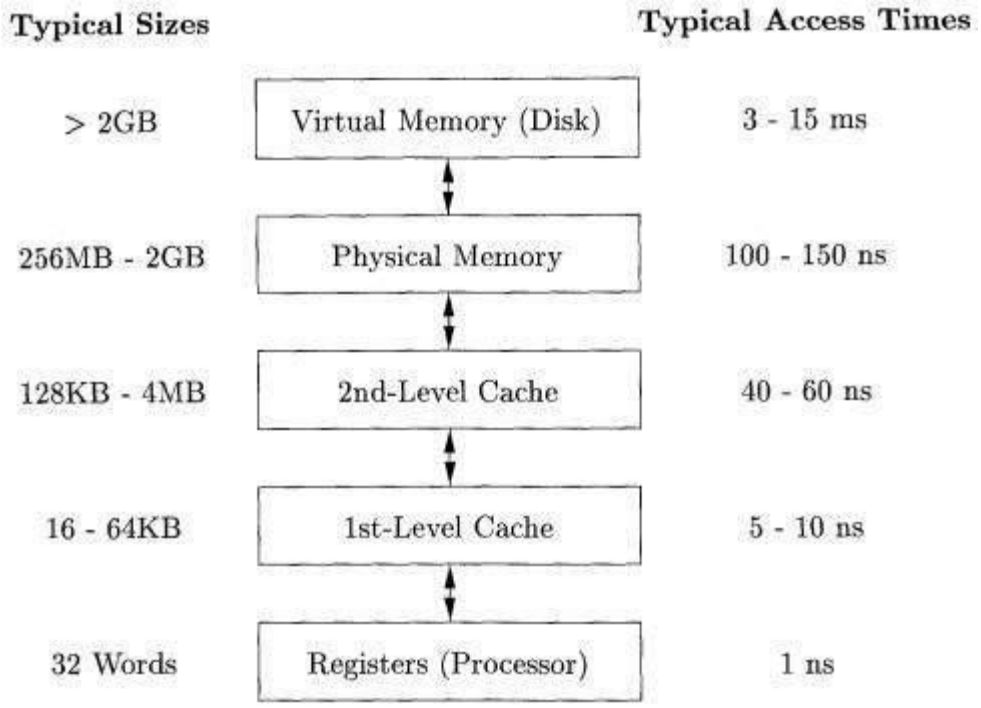


Figure 7.16: Typical Memory Hierarchy Configurations

3. Locality in Programs

Most programs exhibit a high degree of locality; that is, they spend most of their time executing a relatively small fraction of the code and touching only a small fraction of the data. We say that a program has temporal locality if the memory locations it accesses are likely to be accessed again within a short period of time. We say that a program has *spatial locality* if memory locations close to the location accessed are likely also to be accessed within a short period of time.

Programs spend 90% of their time executing 10% of the code. Programs often contain many instructions that are never executed. Programs built with components and libraries use only a small fraction of the provided functionality.

The typical program spends most of its time executing innermost loops and tight recursive cycles in a program. By placing the most common instructions and data in the fast-but-small storage, while leaving the rest in the slow-but-large storage. Average memory-access time of a program can be lowered significantly.

4. Reducing Fragmentation

To begin with the whole heap is a single chunk of size 500K bytes

After a few allocations and deallocations, there are holes

In the above picture, it is not possible to allocate 100K or 150K even though total free memory is 150K

With each deallocation request, the freed chunks of memory are added back to the pool of free space. We coalesce contiguous holes into larger holes, as the holes can only get smaller otherwise. If we are not careful, the memory may end up getting fragmented, consisting of large numbers of small, noncontiguous holes. It is then possible that no hole is large enough to satisfy a future request, even though there may be sufficient aggregate free space.

Best - Fit and Next - Fit Object Placement

We reduce fragmentation by controlling how the memory manager places new objects in the heap. It has been found empirically that a good strategy for minimizing fragmentation for real life programs is to allocate the requested memory in the smallest available hole that is large enough. This *best-fit* algorithm tends to spare the large holes to satisfy subsequent, larger requests. An alternative, called *first-fit*, where an object is placed in the first (lowest-address) hole in which it fits, takes less time to place objects, but has been found inferior to best-fit in overall performance.

To implement best-fit placement more efficiently, we can separate free space into *bins*, according to their sizes.Binning makes it easy to find the best-fit chunk.

M a n a g i n g and Coalescing Free Space

When an object is deallocated manually, the memory manager must make its chunk free, so it can be allocated again. In some circumstances, it may also be possible to combine *(coalesce)* that chunk with adjacent chunks of the heap, to form a larger chunk. There is an advantage to doing so, since we can always use a large chunk to do the work of small chunks of equal total size, but many small chunks cannot hold one large object, as the combined chunk could.

Automatic garbage collection can eliminate fragmentation altogether if it moves all the allocated objects to contiguous storage.

5. Manual Deallocation Requests

In manual memory management, where the programmer must explicitly arrange for the deallocation of data, as in C and C + + . Ideally, any storage that will no longer be accessed should be deleted.

Problems with Manual Deallocation

1. Memory leaks ‰

   Failing to delete data that cannot be referenced ‰

   Important in long running or nonstop programs „

2. Dangling pointer dereferencing ‰

   Referencing deleted data „

Both are serious and hard to debug

Garbage Collection „

1. Reclamation of chunks of storage holding objects that can no longer be accessed by a program „
2. GC should be able to determine types of objects ‰

   Then, size and pointer fields of objects can be determined by the GC ‰

   Languages in which types of objects can be determined at compile time or run-time are type safe „

   Java is type safe „

   C and C++ are not type safe because they permit type casting, which creates new pointers

„

   Thus, any memory location can be (theoretically) accessed at any time and hence cannot be considered inaccessible

Code Generation

It takes as input the intermediate representation (IR) produced by the front end of the compiler, along with relevant symbol table information, and produces as output a semantically equivalent target program

The most important criterion for a code generator is that it produce correct code.

The following issue arises during the code generation phase:

1 Input to the Code Generator
2 The Target Program
3 Instruction Selection
4 Register Allocation
5.Evaluation Order

Input to code generator

The input to code generator is the intermediate code generated by the front end, along with information in the symbol table that determines the run-time addresses of the data-objects denoted by the names in the intermediate representation. Intermediate codes may be represented mostly in quadruples, triples, indirect triples, Postfix notation, syntax trees, DAG's, etc. The code generation phase just proceeds on an assumption that the input are free from all of syntactic and state semantic errors, the necessary type checking has taken place and the type-conversion operators have been inserted wherever necessary

Target program
The target program is the output of the code generator. The output may be absolute machine language, relocatable machine language, assembly language.

- **Absolute** machine language as output has advantages that it can be placed in a fixed memory location and can be immediately executed.
- **Relocatable** machine language as an output allows subprograms and subroutines to be compiled separately. Relocatable object modules can be linked together and loaded by linking loader. But there is added expense of linking and loading.
- **Assembly** language as output makes the code generation easier. We can generate symbolic instructions and use macro-facilities of assembler in generating code. And we need an additional assembly step after code generation.

.

Instruction selection

Selecting the best instructions will improve the efficiency of the program. It includes the instructions that should be complete and uniform. Instruction speeds and machine idioms also plays a major role when efficiency is considered. But if we do not care about the efficiency of the target program then instruction selection is straight-forward.

For example, three-address statements would be translated into the latter code sequence as shown below:

```
P:=Q+R

S:=P+T

MOV Q, R0

ADD R, R0

MOV R0, P

MOV P, R0

ADD T, R0

MOV R0, S
```

Here the fourth statement is redundant as the value of the P is loaded again in that statement that just has been stored in the previous statement. It leads to an inefficient code sequence. A given intermediate representation can be translated into many code sequences, with significant cost differences between the different implementations. A prior knowledge of instruction cost is needed in order to design good sequences, but accurate cost information is difficult to predict.

Register allocation issues

Use of registers make the computations faster in comparison to that of memory, so efficient utilization of registers is important. The use of registers are subdivided into two subproblems:

1. During **Register allocation –** we select only those set of variables that will reside in the registers at each point in the program.
2. During a subsequent **Register assignment** phase, the specific register is picked to access the variable.

As the number of variables increases, the optimal assignment of registers to variables becomes difficult. Mathematically, this problem becomes NP-complete. Certain machine requires register pairs consist of an even and next odd-numbered register. For example

```
M a, b
```

These types of multiplicative instruction involve register pairs where the multiplicand is an even register and b, the multiplier is the odd register of the even/odd register pair.

Evaluation order –

The code generator decides the order in which the instruction will be executed. The order of computations affects the efficiency of the target code. Among many computational orders, some will require only fewer registers to hold the intermediate results. However, picking the best order in the general case is a difficult NP-complete problem.

Approaches to code generation issues:

Code generator must always generate the correct code. It is essential because of the number of special cases that a code generator might face. Some of the design goals of code generator are:

- Correct
- Easily maintainable
- Testable
- Efficient

The target Language

1 A Simple Target Machine Model

2 Program and Instruction Costs

A Simple Target Machine Model

op source, destination

Where, op is used as an op-code and source and destination are used as a data field.

- It has the following op-codes:
  ADD (add source to destination)
  SUB (subtract source from destination)
  MOV (move source to destination)
- The source and destination of an instruction can be specified by the combination of registers and memory location with address modes.

MODE	FORM	ADDRESS	EXAMPLE	ADDED COST
Absolute	M	M	Add R0, R1	1
Register	R	R	Add temp, R1	0
indexed	c(R)	C+ contents(R)	ADD 100 (R2), R1	1
indirect register	*R	contents(R)	ADD * 100	0
indirect indexed	*c(R)	contents(c+ contents(R))	(R2), R1	1
literal	#c	c	ADD #3, R1	1

- Here, cost 1 means that it occupies only one word of memory.
- Each instruction has a cost of 1 plus added costs for the source and destination.
- Instruction cost = 1 + cost is used for source and destination mode.

2 Program and Instruction Costs

Cost of an instruction to be one plus the costs associated with the addressing modes of the operands . This cost corresponds to the length in words of the instruction. Addressing modes involving registers have zero additional cost, while those involving a memory location or constant in them have an additional cost of one, because such operands have to be stored in the words following the instruction**.**

Examples:

• The instruction LD RO, Rl copies the contents of register Rl into register RO. This instruction has a cost of one because no additional memory words are required.

• The instruction LD RO, M loads the contents of memory location M into register RO. T h e cost is two since the address of memory location M is in the word following the instruction.

• The instruction LD R l , *100(R2) loads into register Rl the value given by contents(contents(100 + contents(K2))). The cost is three because the constant 100 is stored in the word following the instruction.

Example:

1. Move register to memory R0 → M

   MOV R0, M

   cost = 1+1+1 (since address of memory location M is in word following the instruction)

2. Indirect indexed mode: MOV * 4(R0), M

   cost = 1+1+1 (since one word **for** memory location M, one word result of *4(R0) and one **for** instruction)

3. Literal Mode:

   MOV #1, R0

   cost = 1+1+1 = 3 (one word **for** constant 1 and one **for** instruction)

Address in the target code

The information which required during an execution of a procedure is kept in a block of storage called an activation record. The activation record includes storage for names local to the procedure. We can describe address in the target code using the following ways:

1. Static allocation
2. Stack allocation

In static allocation, the position of an activation record is fixed in memory at compile time.

.

In the stack allocation, for each execution of a procedure a new activation record is pushed onto the stack. When the activation ends then the record is popped.

For the run-time allocation and deallocation of activation records the following three-address statements are associated:

1. Call
2. Return
3. Halt
4. Action, a placeholder for other statements

Assume that the run-time memory is divided into areas for:

1. Code
2. Static data
3. Stack

Static allocation:

1. Implementation of call statement:

The following code is needed to implement static allocation:

MOV #here + 20, callee.static_area /*it saves return address*/</p>

GOTO callee.code_area /* It transfers control to the target code for the called procedure*/

Where,

callee.static_area shows the address of the activation record.

callee.code_area shows the address of the first instruction for called procedure.

#here + 20 literal are used to return address of the instruction following GOTO.

2. Implementation of return statement:

The following code is needed to implement return from procedure callee:

GOTO * callee.static_area

It is used to transfer the control to the address that is saved at the beginning of the activation record.

3. Implementation of action statement:

The ACTION instruction is used to implement action statement.

4. Implementation of halt statement:

The HALT statement is the final instruction that is used to return the control to the operating system.

Stack allocation

Using the relative address, static allocation can become stack allocation for storage in activation records.

In stack allocation, register is used to store the position of activation record so words in activation records can be accessed as offsets from the value in this register.

The following code is needed to implement stack allocation:

1. Initialization of stack:

```
MOV #stackstart , SP    /*initializes stack*/
HALT                    /*terminate execution*/
```

2. Implementation of Call statement:

```
ADD #caller.recordsize, SP/* increment stack pointer */
MOV #here + 16, *SP             /*Save return address */
GOTO callee.code_area
```

Where,

caller.recordsize is the size of the activation record

#here + 16 is the address of the instruction following the **GOTO**

3. Implementation of Return statement:

```
GOTO *0 ( SP ) /*return to the caller */
SUB #caller.recordsize, SP      /*decrement SP and restore to previous value */
```

Basic blocks and Flow graphs

A graph representation of three-address statements, called a flow graph, is useful for understanding code-generation algorithms, even if the graph is not explicitly constructed by a code-generation algorithm. Nodes in the flow graph represent computations, and the edges represent the flow of control. Flow graph of a program can be used as a vehicle to collect information about the intermediate program. Some register-assignment algorithms use flow graphs to find the inner loops where a program is expected to spend most of its time.

Basic block contains a sequence of statement. The flow of control enters at the beginning of the statement and leave at the end without any halt (except may be the last instruction of the block).

The following sequence of three address statements forms a basic block:

1. t1:= x * x
2. t2:= x * y
3. t3:= 2 * t2
4. t4:= t1 + t3
5. t5:= y * y
6. t6:= t4 + t5

Basic block construction:

Algorithm: Partition into basic blocks

Input: It contains the sequence of three address statements

Output: it contains a list of basic blocks with each three address statement in exactly one block

Method: First identify the leader in the code. The rules for finding leaders are as follows:

- The first statement is a leader.
- Statement L is a leader if there is an conditional or unconditional goto statement like: if. .. goto L or goto L
- Instruction L is a leader if it immediately follows a goto or conditional goto statement like: if goto B or goto B

For each leader, its basic block consists of the leader and all statement up to. It doesn't include the next leader or end of the program.

Consider the following source code for dot product of two vectors a and b of length 10:

```
 begin
    prod :=0;
     i:=1;
do begin
    prod :=prod+ a[i] * b[i];
    i :=i+1;
     end
 while i <= 10
end
```

The three address code for the above source program is given below:

B1

```
(1) prod := 0
 (2) i := 1
```

B2

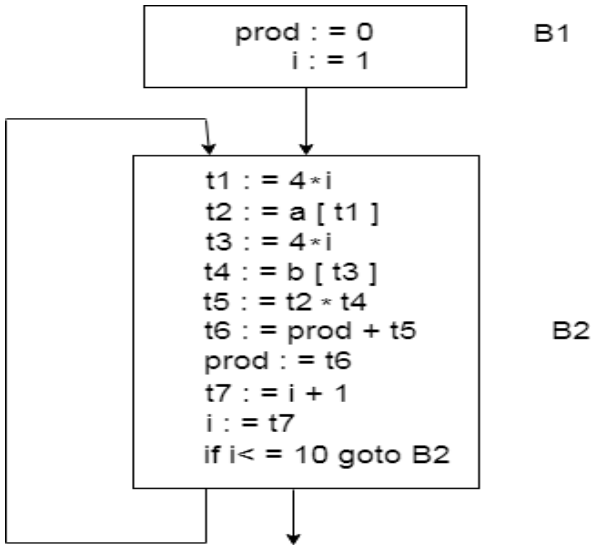
(3) t1 := 4* i
(4) t2 := a[t1]
(5) t3 := 4* i
(6) t4 := b[t3]
(7) t5 := t2*t4
(8) t6 := prod+t5
(9) prod := t6
(10) t7 := i+1
(11) i := t7
(12) if i<=10 goto (3)

Basic block B1 contains the statement (1) to (2)
Basic block B2 contains the statement (3) to (12)

Flow Graph

Flow graph is a directed graph. It contains the flow of control information for the set of basic block.

A control flow graph is used to depict that how the program control is being parsed among the blocks. It is useful in the loop optimization.Flow graph for the vector dot product is given as follows:



1.Block B1 is the initial node. Block B2 immediately follows B1, so from B2 to B1 there is an edge.

2.The target of jump from last statement of B1 is the first statement B2, so from B1 to B2 there is an edge.

A Simple Code generation.

Code generator is used to produce the target code for three-address statements. It uses registers to store the operands of the three address statement.

Consider the three address statement x:= y + z. It can have the following sequence of codes:

MOV x, R_0

ADD y, R_0

Register and Address Descriptors:

- A register descriptor contains the track of what is currently in each register. The register descriptors show that all the registers are initially empty.
- An address descriptor is used to store the location where current value of the name can be found at run time.

A code-generation algorithm:

The algorithm takes a sequence of three-address statements as input. For each three address statement of the form a:= b op c perform the various actions. These are as follows:

1. Invoke a function getreg to find out the location L where the result of computation b op c should be stored.
2. Consult the address description for y to determine y'. If the value of y currently in memory and register both then prefer the register y' . If the value of y is not already in L then generate the instruction **MOV y' , L** to place a copy of y in L.
3. Generate the instruction **OP z' , L** where z' is used to show the current location of z. if z is in both then prefer a register to a memory location. Update the address descriptor of x to indicate that x is in location L. If x is in L then update its descriptor and remove x from all other descriptor.
4. If the current value of y or z have no next uses or not live on exit from the block or in register then alter the register descriptor to indicate that after execution of x : = y op z those register will no longer contain y or z.

Generating Code for Assignment Statements:

The assignment statement d:= (a-b) + (a-c) + (a-c) can be translated into the following sequence of three address code:

t:= a-b

u:= a-c

v:= t +u

d:= v+u

Code sequence for the example is as follows:

Statement	Code Generated	Register descriptor Register empty	Address descriptor
t:= a - b	MOV a, R0 SUB b, R0	R0 contains t	t in R0
u:= a - c	MOV a, R1 SUB c, R1	R0 contains t R1 contains u	t in R0 u in R1
v:= t + u	ADD R1, R0	R0 contains v R1 contains u	u in R1 v in R1
d:= v + u	ADD R1, R0 MOV R0, d	R0 contains d	d in R0 d in R0 and memory

Machine Independent Optimization. The principle sources of Optimization, peep hole Optimization, Introduction to Date flow Analysis.

MACHINE INDEPENDENT OPTIMIZATION

Elimination of unnecessary instructions in object code, or the replacement of one sequence of instructions by a faster sequence of instructions that does the same thing is usually called "code improvement" or "code optimization."

Optimizations are classified into two categories.

1. Machine independent optimizations:

   Machine independent optimizations are program transformations that improve the target code without taking into consideration any properties of the target machine

2. Machine dependant optimizations:

   Machine dependant optimizations are based on register allocation and utilization of special machine-instruction sequences.

The Principal Sources of Optimization

A transformation of a program is called local if it can be performed by looking only at the statements in a basic block; otherwise, it is called global. Many transformations can be performed at both the local and global levels.

Function-Preserving Transformations: There are a number of ways in which a compiler can improve a program without changing the function it computes.

: Common sub expression elimination

Copy propagation,

Dead-code elimination

Constant folding

Common Sub expressions elimination:

An occurrence of an expression E is called a common sub-expression if E was previously computed, and the values of variables in E have not changed since the previous computation. We can avoid recomputing the expression if we can use the previously computed value.

- For example

t1: = 4*i

t2: = a [t1]

t3: = 4*j

t4: = 4*i

t5: = n

t6: = b [t4] +t5

The above code can be optimized using the common sub-expression elimination as

t1: = 4*i
t2: = a [t1]
t3: = 4*j
t5: = n
t6: = b [t1] +t5

The common sub expression t4: =4*i is eliminated as its computation is already in t1 and the value of i is not been changed from definition to use.

Copy Propagation:

Assignments of the form f : = g called copy statements, or copies for short. The idea behind the copy-propagation transformation is to use g for f, whenever possible after the copy statement f: = g. Copy propagation means use of one variable instead of another.

• For example:

x=Pi;
A=x*r*r;

The optimization using copy propagation can be done as follows: A=Pi*r*r;
Here the variable x is eliminated

Dead-Code Eliminations:

A variable is live at a point in a program if its value can be used subsequently; otherwise, it is dead at that point.

Example:

```
i=0;
if(i==1)
{
a=b+5;
}
```

Here, 'if' statement is dead code because this condition will never get satisfied.

Constant folding:

Deducing at compile time that the value of an expression is a constant and using the constant instead is known as constant folding. One advantage of copy propagation is that it often turns the copy statement into dead code.

For example,

a=3.14157/2 can be replaced by
a=1.570

Loop Optimizations:

In loops, especially in the inner loops, programs tend to spend the bulk of their time. The running time of a program may be improved if the number of instructions in an inner loop is decreased, even if we increase the amount of code outside that loop.

Three techniques are important for loop optimization:

1. Code motion, which moves code outside a loop;
2. Induction-variable elimination, which we apply to replace variables from inner loop.
3. Reduction in strength, which replaces expensive operation by a cheaper one, such as a multiplication by an addition.

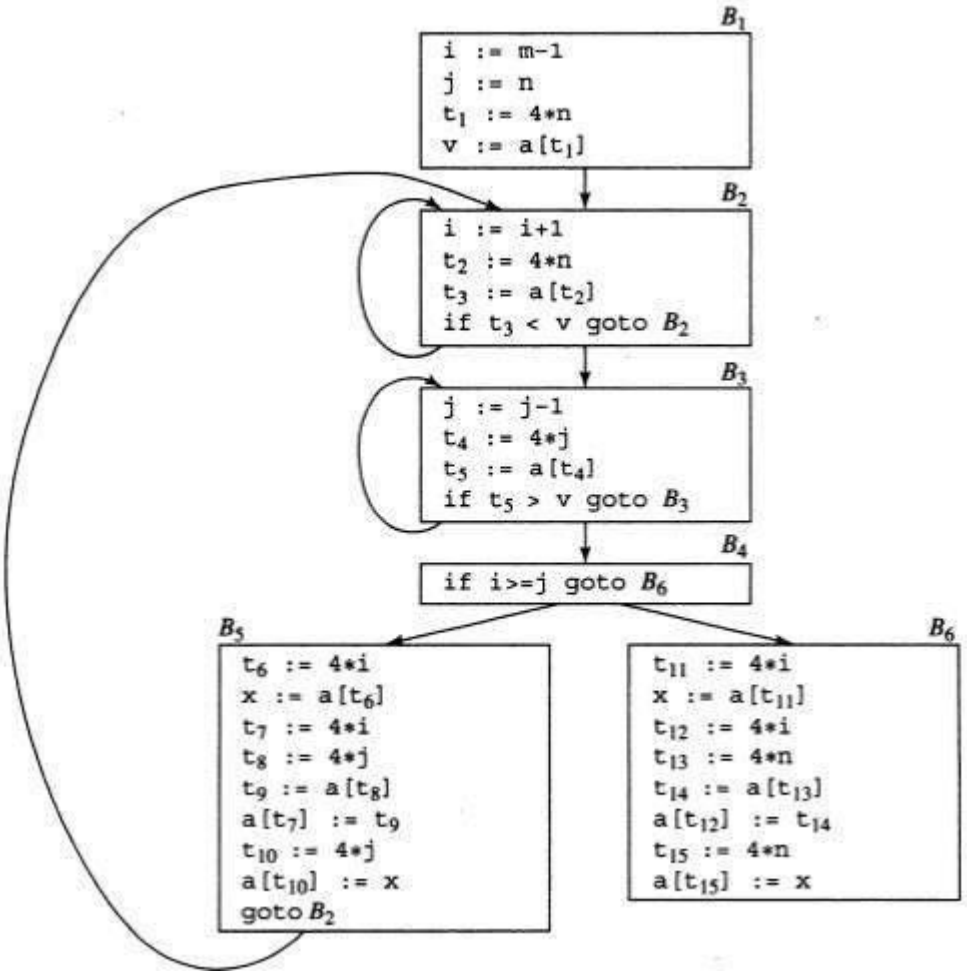


Fig. 5.2 Flow graph

Code Motion:

This transformation takes an expression that yields the same result independent of the number of times a loop is executed (a loop-invariant computation) and places the expression before the loop. Note that the notion "before the loop" assumes the existence of an entry for the loop. For example, evaluation of limit-2 is a loop-invariant computation in the following while-statement:

```
while (i <= limit-2)
```

Code motion will result in the equivalent of

```
t= limit-2;
while (i<=t) /* statement does not change limit or t */
```

Induction Variables :

Loops are usually processed inside out. For example consider the loop around B3. Note that the values of j and t4 remain in lock-step; every time the value of j decreases by 1, that of t4 decreases by 4 because 4*j is assigned to t4. Such identifiers are called induction variables.

When there are two or more induction variables in a loop, it may be possible to get rid of all but one, by the process of induction-variable elimination. For the inner loop around B3 in Fig.5.3 we cannot get rid of either j or t4 completely; t4 is used in B3 and j in B4.

However, we can illustrate reduction in strength and illustrate a part of the process of induction-variable elimination. Eventually j will be eliminated when the outer loop of B2- B5 is considered.

Example:

As the relationship t4:=4*j surely holds after such an assignment to t4 in Fig. and t4 is not changed elsewhere in the inner loop around B3, it follows that just after the statement j:=j-1 the relationship t4:= 4*j-4 must hold. We may therefore replace the assignment t4:= 4*j by t4:= t4-4. The only problem is that t4 does not have a value when we enter block B3 for the first time. Since we must maintain the relationship t4=4*j on entry to the block B3, we place an initializations of t4 at the end of the block where j itself is initialized, shown by the dashed addition to block B1 in Fig.5.3.

The replacement of a multiplication by a subtraction will speed up the object code if multiplication takes more time than addition or subtraction

Reduction In Strength:

Reduction in strength replaces expensive operations by equivalent cheaper ones on the target machine. Certain machine instructions are considerably cheaper than others and can often be used as special cases of more expensive operators. For example, x^2 is invariably cheaper to implement as x*x than as a call to an exponentiation routine. Fixed-point multiplication or division by a power of two is cheaper to implement as a shift. Floating-point division by a constant can be implemented as multiplication by a constant, which may be cheaper.

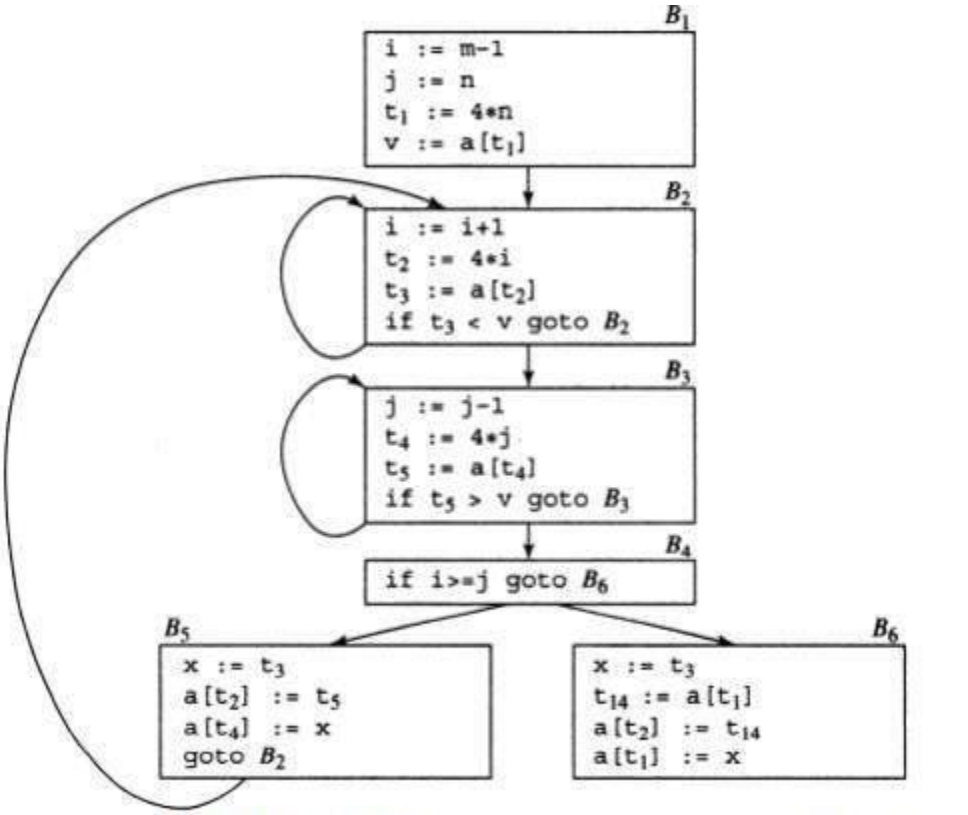


Fig. 5.3 B5 and B6 after common subexpression elimination

Fig. 5.3 B5 and B6 after common subexpression elimination

PEEPHOLE OPTIMIZATION

A statement-by-statement code-generations strategy often produces target code that contains redundant instructions and suboptimal constructs. The quality of such target code can be improved by applying "optimizing" transformations to the target program.

A simple but effective technique for improving the target code is peephole optimization, A method for trying to improving the performance of the target program by examining a short sequence of target instructions (called the peephole) and replacing these instructions by a shorter or faster sequence, whenever possible.

The peephole is a small, moving window on the target program.

Characteristics of peephole optimizations:

Redundant-instructions elimination

Flow-of-control optimizations

Algebraic simplifications

Use of machine idioms

Unreachable code

Redundant-instructions elimination

see the instructions sequence

(1) MOV R0,a

(2) MOV a,R0

we can delete instructions (2) because whenever (2) is executed. (1) will ensure that the value of a is already in register R0.If (2) had a label we could not be sure that (1) was always executed immediately before (2) and so we could not remove (2).

Unreachable Code:

Another opportunity for peephole optimizations is the removal of unreachable instructions. An unlabeled instruction immediately following an unconditional jump may be removed. This operation can be repeated to eliminate a sequence of instructions. For example, for debugging purposes, a large program may have within it certain segments that are executed only if a variable debug is 1. In C, the source code might look like:

```
#define debug 0
….

If ( debug ) {
Print debugging information

}
```

In the intermediate representations the if-statement may be translated as:

If debug =1 goto L1 goto L2

L1: print debugging information L2: ..... (a)

One obvious peephole optimization is to eliminate jumps over jumps .Thus no matter what the value of debug; (a) can be replaced by:

```
        If debug ≠1 goto L2
        Print debugging information
L2:.....(b)

        If debug ≠0 goto L2
        Print debugging information
L2:.....(c)
```

As the argument of the statement of (c) evaluates to a constant true it can be replaced

By goto L2. Then all the statement that print debugging aids are manifestly unreachable and can be eliminated one at a time.

Flows-Of-Control Optimizations:

The unnecessary jumps can be eliminated in either the intermediate code or the target code by the following types of peephole optimizations. We can replace the jump sequence

goto L1
….

L1: gotoL2 (d)

by the sequence

goto L2
….

L1: goto L2

If there are now no jumps to L1, then it may be possible to eliminate the statement L1:goto L2 provided it is preceded by an unconditional jump .Similarly, the sequence

if a < b goto L1
….

L1: goto L2 (e)

can be replaced by

If a < b goto L2

….

L1: goto L2

Ø Finally, suppose there is only one jump to L1 and L1 is preceded by an unconditional goto. Then the sequence

goto L1

L1: if a < b goto L2 (f) L3:

may be replaced by

If a < b goto L2
goto L3

…….

L3:

While the number of instructions in(e) and (f) is the same, we sometimes skip the unconditional jump in (f), but never in (e).Thus (f) is superior to (e) in execution time

Algebraic Simplification:

There is no end to the amount of algebraic simplification that can be attempted through peephole optimization. Only a few algebraic identities occur frequently enough that it is worth considering implementing them. For example, statements such as

x := x+0 or
x := x * 1

are often produced by straightforward intermediate code-generation algorithms, and they can be eliminated easily through peephole optimization.

Reduction in Strength:

Reduction in strength replaces expensive operations by equivalent cheaper ones on the target machine. Certain machine instructions are considerably cheaper than others and can often be used as special cases of more expensive operators.

For example, x^2 is invariably cheaper to implement as x*x than as a call to an exponentiation routine. Fixed-point multiplication or division by a power of two is cheaper to implement as a shift. Floating-point division by a constant can be implemented as multiplication by a constant, which may be cheaper.

$X^2 \rightarrow X*X$

Use of Machine Idioms:

The target machine may have hardware instructions to implement certain specific operations efficiently. For example, some machines have auto-increment and auto-decrement addressing modes. These add or subtract one from an operand before or after using its value. The use of these modes greatly improves the quality of code when pushing or popping a stack, as in parameter passing. These modes can also be used in code for statements like i : =i+1.

i:=i+1 → i++

i:=i-1 → i- -

Introduction to Date flow Analysis.

1 The Data-Flow Abstraction

2 The Data-Flow Analysis Schema

3 Data-Flow Schemas on Basic Blocks

4 Reaching Definitions

5 Live-Variable Analysis

6 Available Expressions

"Data-flow analysis" refers to a body of techniques that derive information about the flow of data along program execution paths.

1. The Data-Flow Abstraction

The execution of a program can be viewed as a series of transformations of the program state, which consists of the values of all the variables in the program. Each execution of an intermediate-code statement transforms an input state to a new output state. The input state is associated with the *program point before* the statement and the output state is associated with the *program point after* the statement.

When we analyze the behavior of a program, we must consider all the possible sequences of program points ("paths") through a flow graph that the program execution can take. We then extract, from the possible program states at each point, the information we need for the particular data-flow analysis problem we want to solve. In more complex analyses, we must consider paths that jump among the flow graphs for various procedures, as calls and returns are executed.

Within one basic block, the program point after a statement is the same as the program point before the next statement.

If there is an edge from block *B1* to block $B2_2$, then the program point after the last statement of *B1* may be followed immediately by the program point before the first statement of B_2.

Thus, we may define an execution path (or just path) from point pi to point pn to be a sequence of points pi,p2,... ,pn such that for each i = 1,2, ... ,n - 1, either

1. Pi is the point immediately preceding a statement and $pi_{+}i$ is the point immediately following that same statement, or

2. pi is the end of some block and p_{i+1} is the beginning of a successor block.

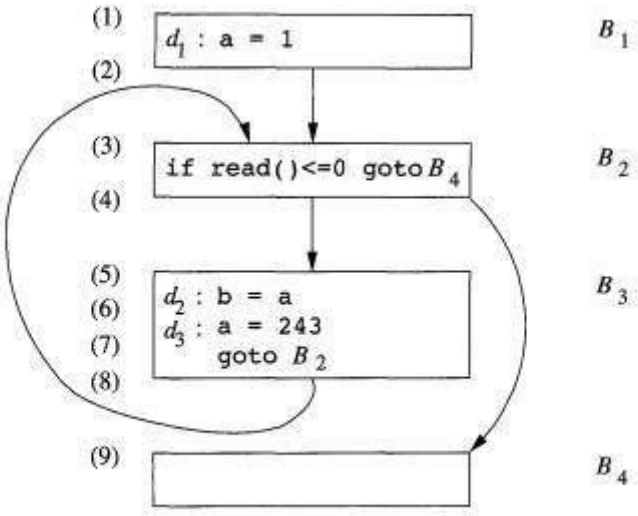


Figure 9.12: Example program illustrating the data-flow abstraction

. In data-flow analysis, we do not distinguish among the paths taken to reach a program point. Moreover, we do not keep track of entire states; rather, we abstract out certain details, keeping only the data we need for the purpose of the analysis. Two examples will illustrate how the same program states may lead to different information abstracted at a point.

1. To help users debug their programs, we may wish to find out what are all the values a variable may have at a program point, and where these values may be defined. For instance, we may summarize all the program states at point (5) by saying that the value of *a* is one of {1,243}, and that it may be defined by one of **{ ^ 1 , ^ 3 } .** The definitions that *may* reach a program point along some path are known as *reaching definitions.*

2. Suppose, instead, we are interested in implementing constant folding. If a use of the variable *x* is reached by only one definition, and that definition assigns a constant to *x,* then we can simply replace *x* by the constant. If, on the other hand, several definitions of *x* may reach a single program point, then we cannot perform constant folding on *x.* Thus, for constant folding we wish to find those definitions that are the unique definition of their variable to reach a given program point, no matter which execution path is taken. For point (5) of Fig. 9.12, there is no definition that *must* be the definition of *a* at that point, so this set is empty for *a* at point (5). Even if a variable has a unique definition at a point, that definition must assign a constant to the variable. Thus, we may simply describe certain variables as "not a constant," instead of collecting all their possible values or all their possible definitions.

2. The Data-Flow Analysis Schema

, we associate with every program point a data-flow value that represents an abstraction of the set of all possible program states that can be observed for that point. The set of possible data-flow values is the domain for this application. For example, the domain of data-flow values for reaching definitions is the set of all subsets of definitions in the program.

A particular data-flow value is a set of definitions, and we want to associate with each point in the program the exact set of definitions that can reach that point. As discussed above, the choice of abstraction depends on the goal of the analysis; to be efficient, we only keep track of information that is relevant.

Denote the data-flow values before and after each statement s by IN[S] and OUT[s], respectively. The data-flow problem is to find a solution to a set of constraints on the IN[S]'S and OUT[s]'s, for all statements s. There are two sets of constraints: those based on the semantics of the statements ("transfer functions") and those based on the flow of control.

Transfer Functions

The data-flow values before and after a statement are constrained by the semantics of the statement. For example, suppose our data-flow analysis involves determining the constant value of variables at points. If variable *a* has value *v* before executing statement **b** = a, then both *a* and *b* will have the value *v* after the statement. This relationship between the data-flow values before and after the assignment statement is known as a transfer function.

Transfer functions come in two flavors: information may propagate forward along execution paths, or it may flow backwards up the execution paths. In a forward-flow problem, the transfer function of a statement s, which we shall usually denote f(a), takes the data-flow value before the statement and produces a new data-flow value after the statement. That is,

$$\text{OUT}[s] = f_s(\text{IN}[s]).$$

Conversely, in a backward-flow problem, the transfer function f(a) for statement 8 converts a data-flow value after the statement to a new data-flow value before the statement. That is,

$$\text{IN}[s] = f_s(\text{OUT}[s]).$$

Control – Flow Constraints

The second set of constraints on data-flow values is derived from the flow of control. Within a basic block, control flow is simple. If a block *B* consists of statements s1, s_2, • • • ,s_n in that order, then the control-flow value out of ***Si*** is the same as the control-flow value into ***Si+i.*** That is,

$$\text{IN}[s_{i+1}] = \text{OUT}[s_i], \text{ for all } i = 1, 2, \ldots, n-1.$$

However, control-flow edges between basic blocks create more complex constraints between the last statement of one basic block and the first statement of the following block. For example, if we are interested in collecting all the definitions that may reach a program point, then the set of definitions reaching the leader statement of a basic block is the union of the definitions after the last statements of each of the predecessor blocks. The next section gives the details of how data flows among the blocks.

3. Data-Flow Schemas on Basic Blocks

While a data-flow schema involves data-flow values at each point in the program, we can save time and space by recognizing that what goes on inside a block is usually quite simple. Control flows from the beginning to the end of the block, without interruption or branching. Thus, we can restate the schema in terms of data-flow values entering and leaving the blocks. We denote the data-flow values immediately before and immediately after each basic block *B* by *m[B]* and 0 U T [S] , respectively. The constraints involving *m[B]* and 0UT[B] can be derived from those involving ***w[s]*** and OUT[s] for the various statements ***s*** in *B* as follows.

Suppose block B consists of statements s 1 , . . . , sn, in that order. If si is the first statement of basic block B, then m[B] = I N [S I] , Similarly, if sn is the last statement of basic block B, then OUT[S] = OUT[s,,] . The transfer function of a basic block B, which we denote fB, can be derived by composing the transfer functions of the statements in the block. That is, let fa. be the transfer function of statement st. Then of statement si. Then fB = f,sn, o . . . o f,s2, o fsl. . The relationship between the beginning and end of the block is

$$\text{OUT}[B] = f_B(\text{IN}[B]).$$

The constraints due to control flow between basic blocks can easily be rewritten by substituting *IN[B]* and *OUT[B]* for **IN[SI**] and **OUT[sn],** respectively. For instance, if data-flow values are information about the sets of constants that *may* be assigned to a variable, then we have a forward-flow problem in which

$$\text{IN}[B] = \bigcup_{P \text{ a predecessor of } B} \text{OUT}[P].$$

When the data-flow is backwards as we shall soon see in live-variable analy-sis, the equations are similar, but with the roles of the IN's and OUT's reversed. That is,

$$\text{IN}[B] = f_B(\text{OUT}[B])$$
$$\text{OUT}[B] = \bigcup_{S \text{ a successor of } B} \text{IN}[S].$$

Unlike linear arithmetic equations, the data-flow equations usually do not have a unique solution. Our goal is to find the most "precise" solution that satisfies the two sets of constraints: control-flow and transfer constraints. That is, we need a solution that encourages valid code improvements, but does not justify unsafe transformations — those that change what the program computes.

4. Reaching Definitions

"Reaching definitions" is one of the most common and useful data-flow schemas. By knowing where in a program each variable *x* may have been defined when control reaches each point p, we can determine many things about *x.* For just two examples, a compiler then knows whether *x* is a constant at point *p,* and a debugger can tell whether it is possible for *x* to be an undefined variable, should *x* be used at *p.*

We say a definition *d reaches* a point *p* if there is a path from the point immediately following *d* to *p,* such that *d* is not "killed" along that path. We *kill* a definition of a variable *x* if there is any other definition of *x* anywhere along the path . if a definition *d* of some variable *x* reaches point p, then *d* might be the place at which the value of *x* used at *p* was last defined.

A definition of a variable *x* is a statement that assigns, or may assign, a value to *x.* Procedure parameters, array accesses, and indirect references all may have aliases, and it is not easy to tell if a statement is referring to a particular variable *x.* Program analysis must be conservative; if we do not note that the path may have loops, so we could come to another occurrence of *d* along the path, which does not "kill" *d.*

know whether a statement *s* is assigning a value to *x,* we must assume that it *may* assign to it; that is, variable *x* after statement *s* may have either its original value before *s* or the new value created by *s.* For the sake of simplicity, the rest of the chapter assumes that we are dealing only with variables that have no aliases. This class of variables includes all local scalar variables in most languages; in the case of C and *C++,* local variables whose addresses have been computed at some point are excluded.

Transfer Equations for Reaching Definitions

Start by examining the details of a single statement. Consider a definition

```
d: u = v+w
```

Here, and frequently in what follows, + is used as a generic binary operator. This statement "generates" a definition *d* of variable *u* and "kills" all the

other definitions in the program that define variable *u,* while leaving the re-maining incoming definitions unaffected. The transfer function of definition *d* thus can be expressed as

$$f_d(x) = gen_d \cup (x - kill_d) \qquad (9.1)$$

where gend = {d}, the set of definitions generated by the statement, and killd is the set of all other definitions of u in the program.

The transfer function of a basic block can be found by composing the transfer functions of the statements contained therein. The composition of functions of the form (9.1), which we shall refer to as "gen-kill form," is also of that form, as we can see as follows. Suppose there are two functions fi(x) = gen1 U (x - kill1) and f2(x) = gen2 U (x — kill2). Then

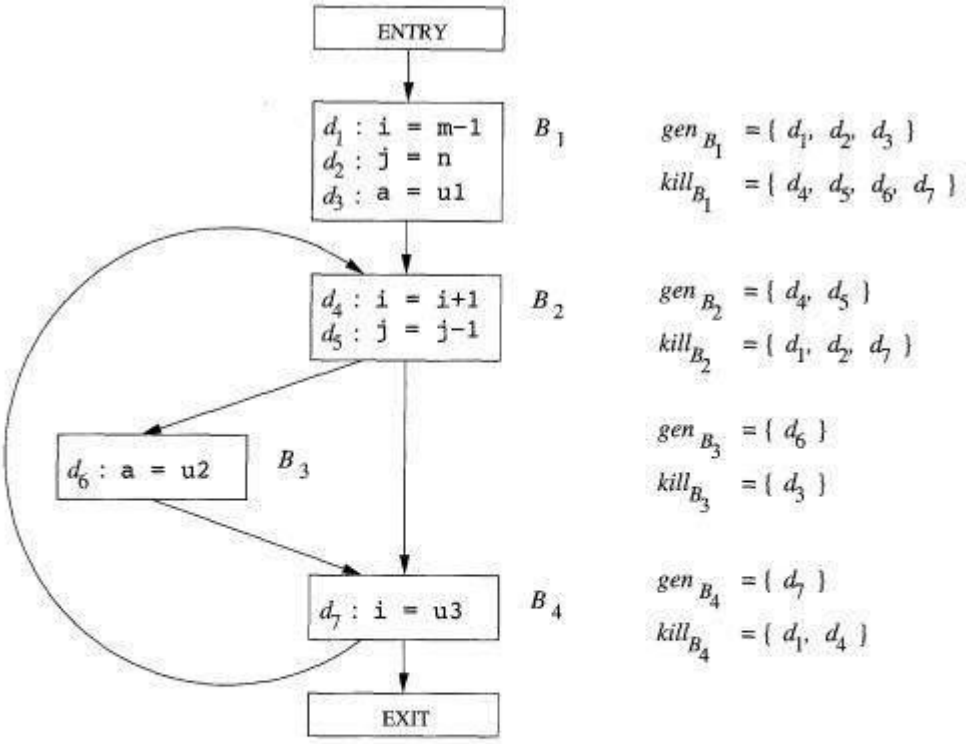


Figure 9.13: Flow graph for illustrating reaching definitions

$$\begin{aligned} f_2(f_1(x)) &= gen_2 \cup (gen_1 \cup (x - kill_1) - kill_2) \\ &= (gen_2 \cup (gen_1 - kill_2)) \cup (x - (kill_1 \cup kill_2)) \end{aligned}$$

This rule extends to a block consisting of any number of statements. Suppose block *B* has *n* statements, with transfer functions *fi(x) = geni* U *(x — kilh)* for i = 1,2, ... , n. Then the transfer function for block *B* may be written as:

$$f_B(x) = gen_B \cup (x - kill_B),$$

where

$$kill_B = kill_1 \cup kill_2 \cup \cdots \cup kill_n$$

and

$$\begin{aligned} gen_B &= gen_n \cup (gen_{n-1} - kill_n) \cup (gen_{n-2} - kill_{n-1} - kill_n) \cup \\ &\quad \cdots \cup (gen_1 - kill_2 - kill_3 - \cdots - kill_n) \end{aligned}$$

Thus, like a statement, a basic block also generates a set of definitions and kills a set of definitions. The gen set contains all the definitions inside the block that are "visible" immediately after the block — we refer to them as downwards exposed. A definition is downwards exposed in a basic block only if it is

not "killed" by a subsequent definition to the same variable inside the same basic block. A basic block's kill set is simply the union of all the definitions killed by the individual statements. Notice that a definition may appear in both the gen and kill set of a basic block. If so, the fact that it is in gen takes precedence, because in gen-kill form, the kill set is applied before the gen set.

Example 9 . 1 0 : The gen set for the basic block

d_1: a = 3
d_2: a = 4

is {d2} since d1 is not downwards exposed. The kill set contains both d1 and d2, since d1 kills d2 and vice versa. Nonetheless, since the subtraction of the kill set precedes the union operation with the gen set, the result of the transfer function for this block always includes definition d2.

Control - Flow Equations

Next, we consider the set of constraints derived from the control flow between basic blocks. Since a definition reaches a program point as long as there exists at least one path along which the definition reaches, **O U T** [P] **C** *m[B]* whenever there is a control-flow edge from *P* to *B.* However, since a definition cannot reach a point unless there is a path along which it reaches, *w[B]* needs to be no larger than the union of the reaching definitions of all the predecessor blocks. That is, it is safe to assume

$$\text{IN}[B] = \bigcup_{P \text{ a predecessor of } B} \text{OUT}[P]$$

We refer to union as the *meet operator* for reaching definitions. In any data-flow schema, the meet operator is the one we use to create a summary of the contributions from different paths at the confluence of those paths.

Iterative Algorithm for Reaching Definitions

We assume that every control-flow graph has two empty basic blocks, an ENTRY node, which represents the starting point of the graph, and an EXIT node to which all exits out of the graph go. Since no definitions reach the beginning of the graph, the transfer function for the ENTRY block is a simple constant function that returns 0 as an answer. That is, O U T [E N T R Y] = 0.

The reaching definitions problem is defined by the following equations:

$$\text{OUT}[\text{ENTRY}] = \emptyset$$

and for all basic blocks B other than ENTRY,

$$\text{OUT}[B] = gen_B \cup (\text{IN}[B] - kill_B)$$

$$\text{IN}[B] = \bigcup_{P \text{ a predecessor of } B} \text{OUT}[P].$$

These equations can be solved using the following algorithm. The result of the algorithm is the *least fixedpoint* of the equations, i.e., the solution whose assigned values to the **IN** ' s and OUT's is contained in the corresponding values for any other solution to the equations. The result of the algorithm below is acceptable, since any definition in one of the sets **IN** or OUT surely must reach the point described. It is a desirable solution, since it does not include any definitions that we can be sure do not reach.

A l g o r i t h m 9 . 1 1 : Reaching definitions.

INPUT: A flow graph for which *kills* and gen_B have been computed for each block *B*.

OUTPUT: I N [B] and O U T **[B] ,** the set of definitions reaching the entry and exit of each block *B* of the flow graph.

METHOD: We use an iterative approach, in which we start with the "estimate" OUT[JB] = 0 for all *B* and converge to the desired values of **IN** and OUT. As we must iterate until the **IN** ' s (and hence the OUT's) converge, we could use a boolean variable *change* to record, on each pass through the blocks, whether any OUT has changed. However, in this and in similar algorithms described later, we assume that the exact mechanism for keeping track of changes is understood, and we elide those details.

The algorithm is sketched in Fig. 9.14. The first two lines initialize certain data-flow values.[4] Line (3) starts the loop in which we iterate until convergence, and the inner loop of lines (4) through (6) applies the data-flow equations to every block other than the entry. •

Algorithm 9.11 propagates definitions as far as they will go with-out being killed, thus simulating all possible executions of the program. Algo-rithm 9.11 will eventually halt, because for every *B*, OUT[B] never shrinks; once a definition is added, it stays there forever. (See Exercise 9.2.6.) Since the set of all definitions is finite, eventually there must be a pass of the while-loop during which nothing is added to any OUT, and the algorithm then terminates. We are safe terminating then because if the OUT's have not changed, the **IN** ' s will

```
1)  OUT[ENTRY] = ∅;
2)  for (each basic block B other than ENTRY) OUT[B] = ∅;
3)  while (changes to any OUT occur)
4)          for (each basic block B other than ENTRY) {
5)                  IN[B] = ⋃_{P a predecessor of B} OUT[P];
6)                  OUT[B] = gen_B ∪ (IN[B] − kill_B);
            }
```

Figure 9.14: Iterative algorithm to compute reaching definitions

not change on the next pass. And, if the IN'S do not change, the OUT's cannot, so on all subsequent passes there can be no changes.

The number of nodes in the flow graph is an upper bound on the number of times around the while-loop. The reason is that if a definition reaches a point, it can do so along a cycle-free path, and the number of nodes in a flow graph is an upper bound on the number of nodes in a cycle-free path. Each

time around the while-loop, each definition progresses by at least one node along the path in question, and it often progresses by more than one node, depending on the order in which the nodes are visited.

In fact, if we properly order the blocks in the for-loop of line (5), there is empirical evidence that the average number of iterations of the while-loop is under 5 (see Section 9.6.7). Since sets of definitions can be represented by bit vectors, and the operations on these sets can be implemented by logical operations on the bit vectors, Algorithm 9.11 is surprisingly efficient in practice.

Example 9 . 1 2 : We shall represent the seven definitions d1, d2, • • • ,d>j in the flow graph of Fig. 9.13 by bit vectors, where bit i from the left represents definition d{. The union of sets is computed by taking the logical OR of the corresponding bit vectors. The difference of two sets S — T is computed by complementing the bit vector of T, and then taking the logical AND of that complement, with the bit vector for S.

Shown in the table of Fig. 9.15 are the values taken on by the IN and OUT sets in Algorithm 9.11. The initial values, indicated by a superscript 0, as in OUTfS]0 , are assigned, by the loop of line (2) of Fig. 9.14. They are each the empty set, represented by bit vector 000 0000. The values of subsequent passes of the algorithm are also indicated by superscripts, and labeled IN [I?]1 and OUTfS]1 for the first pass and m[Bf and OUT[S]2 for the second.

Suppose the for-loop of lines (4) through (6) is executed with B taking on the values

$$B_1, B_2, B_3, B_4, \text{EXIT}$$

in that order. With B = B1, since OUT [ENTRY] = 0, [IN B1]-Pow(1) is the empty set, and OUT[P1]1 is genBl. This value differs from the previous value OUT[Si]0 , so

Block B	$\text{OUT}[B]^0$	$\text{IN}[B]^1$	$\text{OUT}[B]^1$	$\text{IN}[B]^2$	$\text{OUT}[B]^2$
B_1	000 0000	000 0000	111 0000	000 0000	111 0000
B_2	000 0000	111 0000	001 1100	111 0111	001 1110
B_3	000 0000	001 1100	000 1110	001 1110	000 1110
B_4	000 0000	001 1110	001 0111	001 1110	001 0111
EXİT	000 0000	001 0111	001 0111	001 0111	001 0111

Figure 9.15: Computation of IN and OUT

we now know there is a change on the first round (and will proceed to a second round).

Then we consider $B = B_2$ and compute

$$\begin{aligned}\text{IN}[B_2]^1 &= \text{OUT}[B_1]^1 \cup \text{OUT}[B_4]^0 \\ &= 111\ 0000 + 000\ 0000 = 111\ 0000 \\ \text{OUT}[B_2]^1 &= gen[B_2] \cup (\text{IN}[B_2]^1 - kill[B_2]) \\ &= 000\ 1100 + (111\ 0000 - 110\ 0001) = 001\ 1100\end{aligned}$$

This computation is summarized in Fig. 9.15. For instance, at the end of the first pass, OUT [5 2] 1 = 001 1100, reflecting the fact that d4 and d5 are generated in B2, while d3 reaches the beginning of B2 and is not killed in B2.

Notice that after the second round, OUT [B2] has changed to reflect the fact that d& also reaches the beginning of B2 and is not killed by B2. We did not learn that fact on the first pass, because the path from d6 to the end of B2, which is B3 -» B4 -> B2, is not traversed in that order by a single pass. That is, by the time we learn that d$ reaches the end of B4, we have already computed IN[B2] and OUT [B 2] on the first pass.

There are no changes in any of the OUT sets after the second pass. Thus, after a third pass, the algorithm terminates, with the IN's and OUT's as in the final two columns of Fig. 9.15.

5. Live-Variable Analysis

Some code-improving transformations depend on information computed in the direction opposite to the flow of control in a program; we shall examine one such example now. In *live-variable analysis* we wish to know for variable *x* and point *p* whether the value of *x* at *p* could be used along some path in the flow graph starting at *p.* If so, we say *x* is *live* at *p;* otherwise, *x* is *dead* at *p.*

An important use for live-variable information is register allocation for basic blocks. Aspects of this issue were introduced in Sections 8.6 and 8.8. After a value is computed in a register, and presumably used within a block, it is not necessary to store that value if it is dead at the end of the block. Also, if all registers are full and we need another register, we should favor using a register with a dead value, since that value does not have to be stored.

Here, we define the data-flow equations directly in terms of IN [5] and OUTpB], which represent the set of variables live at the points immediately before and after block B, respectively. These equations can also be derived by first defining the transfer functions of individual statements and composing them to create the transfer function of a basic block. Define

1. defB as the set of variables defined (i.e., definitely assigned values) in B prior to any use of that variable in B, and useB as the set of variables whose values may be used in B prior to any definition of the variable.

Example 9 . 1 3 : For instance, block B2 in Fig. 9.13 definitely uses i. It also uses j before any redefinition of j, unless it is possible that i and j are aliases of one another. Assuming there are no aliases among the variables in Fig. 9.13, then uses2 = {i,j}- Also, B2 clearly defines i and j. Assuming there are no aliases, defB2 = as well.

As a consequence of the definitions, any variable in useB must be considered live on entrance to block B, while definitions of variables in defB definitely are dead at the beginning of B. In effect, membership in defB "kills" any opportunity for a variable to be live because of paths that begin at B.

Thus, the equations relating def and use to the unknowns IN and OUT are defined as follows:

$$\text{IN}[\text{EXIT}] = \emptyset$$

and for all basic blocks B other than EXIT,

$$\text{IN}[B] = use_B \cup (\text{OUT}[B] - def_B)$$
$$\text{OUT}[B] = \bigcup_{S \text{ a successor of } B} \text{IN}[S]$$

The first equation specifies the boundary condition, which is that no variables are live on exit from the program. The second equation says that a variable is live coming into a block if either it is used before redefinition in the block or it is live coming out of the block and is not redefined in the block. The third equation says that a variable is live coming out of a block if and only if it is live coming into one of its successors.

The relationship between the equations for liveness and the reaching-defin-itions equations should be noticed:

Both sets of equations have union as the meet operator. The reason is that in each data-flow schema we propagate information along paths, and we care only about whether *any* path with desired properties exist, rather than whether something is true along *all* paths.

• However, information flow for liveness travels "backward," opposite to the direction of control flow, because in this problem we want to make sure that the use of a variable *x* at a point *p* is transmitted to all points prior to *p* in an execution path, so that we may know at the prior point that *x* will have its value used.

To solve a backward problem, instead of initializing O U T [E N T R Y] , we initialize I N [EXIT] . Sets I N and O U T have their roles interchanged, and use and def substitute for gen and kill, respectively. As for reaching definitions, the solution to the liveness equations is not necessarily unique, and we want the so-lution with the smallest sets of live variables. The algorithm used is essentially a backwards version of Algorithm 9.11.

Algorithm 9 . 1 4 : Live-variable analysis.

INPUT: A flow graph with def and use computed for each block.

OUTPUT: m[B] and O U T [£] , the set of variables live on entry and exit of each block B of the flow graph.

```
IN[EXIT] = ∅;
for (each basic block B other than EXIT) IN[B] = ∅;
while (changes to any IN occur)
        for (each basic block B other than EXIT) {
                OUT[B] = ⋃S a successor of B IN[S];
                IN[B] = use_B ∪ (OUT[B] − def_B);
        }
```

Figure 9.16: Iterative algorithm to compute live variables

6. Available Expressions

An expression x + y is available at a point p if every path from the entry node to p evaluates x + y, and after the last such evaluation prior to reaching p, there are no subsequent assignments to x or y.5 For the available-expressions data-flow schema we say that a block kills expression x + y if it assigns (or may 5 N o te that, as usual in this chapter, we use the operator + as a generic operator, not necessarily standing for addition.

assign) x or y and does not subsequently recompute x + y. A block generates expression x + y if it definitely evaluates x + y and does not subsequently define x or y.

Note that the notion of "killing" or "generating" an available expression is not exactly the same as that for reaching definitions. Nevertheless, these notions of "kill" and "generate" behave essentially as they do for reaching definitions.

The primary use of available-expression information is for detecting global common subexpressions. For example, in Fig. 9.17(a), the expression 4 * i in block Bs will be a common subexpression if 4 * i is available at the entry point of block B3. It will be available if i is not assigned a new value in block B2, or if, as in Fig. 9.17(b), 4 * i is recomputed after i is assigned in B2.

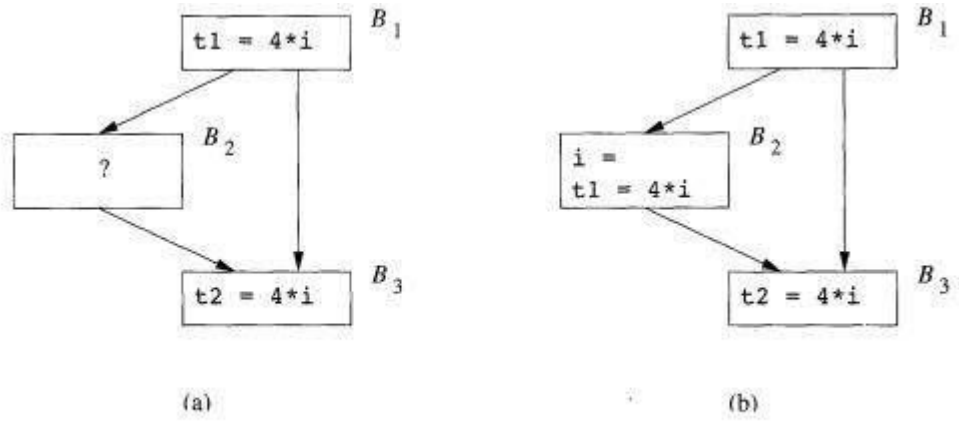


Figure 9.17: Potential common subexpressions across blocks

We can compute the set of generated expressions for each point in a block, working from beginning to end of the block. At the point prior to the block, no expressions are generated. If at point p set *S* of

expressions is available, and *q* is the point after p, with statement x = y+**z** between them, then we form the set of expressions available at *q* by the following two steps.

Add to *S* the expression *y + z.*

Delete from *S* any expression involving variable *x.*

Note the steps must be done in the correct order, as *x* could be the same as *y* or *z.* After we reach the end of the block, *S* is the set of generated expressions for the block. The set of killed expressions is all expressions, say *y + z,* such that either *y* or *z* is defined in the block, and *y + z* is not generated by the block.

E x a m p l e 9.15 : Consider the four statements of Fig. 9.18. After the first, ***b + c*** is available. After the second statement, a — ***d*** becomes available, but ***b + c*** is no longer available, because ***b*** has been redefined. The third statement does not make ***b + c*** available again, because the value of c is immediately changed.

After the last statement, *a — d* is no longer available, because *d* has changed. Thus no expressions are generated, and all expressions involving *a, b,* c, or *d* are killed.

Statement	Available Expressions
	$\emptyset$
`a = b + c`	
	$\{b+c\}$
`b = a - d`	
	$\{a-d\}$
`c = b + c`	
	$\{a-d\}$
`d = a - d`	
	$\emptyset$

Figure 9.18: Computation of available expressions

We can find available expressions in a manner reminiscent of the way reach-ing definitions are computed. Suppose *U* is the "universal" set of all expressions appearing on the right of one or more statements of the program. For each block *B,* let ***IN[B]*** be the set of expressions in *U* that are available at the point just before the beginning of *B.* Let **OUT***[B]* be the same for the point following the end of *B.* Define ***e.genB*** to be the expressions generated by *B* and *eJnills* to be the set of expressions in *U* killed in *B.* Note that **I N , O U T ,** e_#en, and *eJkill* can all be represented by bit vectors. The following equations relate the unknowns

IN and OUT to each other and the known quantities e_gen and e_kill:

$$\text{OUT}[\text{ENTRY}] = \emptyset$$

and for all basic blocks B other than ENTRY,

$$\text{OUT}[B] = e\_gen_B \cup (\text{IN}[B] - e\_kill_B)$$
$$\text{IN}[B] = \bigcap_{P \text{ a predecessor of } B} \text{OUT}[P].$$

T he above equations look almost identical to the equations for reaching definitions. Like reaching definitions, the boundary condition is OUT [ENTRY] = 0, because at the exit of the E N T R Y node, there are no available expressions.

The most important difference is that the meet operator is intersection rather than union. This operator is the proper one because an expression is available at the beginning of a block only if it is available at the end of all its predecessors. In contrast, a definition reaches the beginning of a block whenever it reaches the end of any one or more of its predecessors.

The use of D rather than U makes the available-expression equations behave differently from those of reaching definitions. While neither set has a unique solution, for reaching definitions, it is the solution with the smallest sets that corresponds to the definition of "reaching," and we obtained that solution by starting with the assumption that nothing reached anywhere, and building up to the solution. In that way, we never assumed that a definition *d* could reach a point *p* unless an actual path propagating *d* to *p* could be found. In contrast, for available expression equations we want the solution with the largest sets of available expressions, so we start with an approximation that is too large and work down.

It may not be obvious that by starting with the assumption "everything (i.e., the set *U)* is available everywhere except at the end of the entry block" and eliminating only those expressions for which we can discover a path along which it is not available, we do reach a set of truly available expressions. In the case of available expressions, it is conservative to produce a subset of the exact set of available expressions. The argument for subsets being conservative is that our intended use of the information is to replace the computation of an available expression by a previously computed value. Not knowing an expres-sion is available only inhibits us from improving the code, while believing an expression is available when it is not could cause us to change what the program computes.

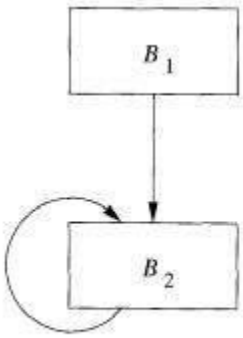


Figure 9.19: Initializing the OUT sets to $\emptyset$ is too restrictive.

Example 9 . 1 6 : We shall concentrate on a single block, B2 in Fig. 9.19, to illustrate the effect of the initial approximation of OUT[B2] on IN [B 2] - Let G and K abbreviate e.genB2 and e-killB2, respectively. The data-flow equations for block B2 are

$$\text{IN}[B_2] = \text{OUT}[B_1] \cap \text{OUT}[B_2]$$

$$\text{OUT}[B_2] = G \cup (\text{IN}[B_2] - K)$$

These equations may be rewritten as recurrences, with I^j and O^j being the jth approximations of $\text{IN}[B_2]$ and $\text{OUT}[B_2]$, respectively:

$$I^{j+1} = \text{OUT}[B_1] \cap O^j$$

$$O^{j+1} = G \cup (I^{j+1} - K)$$

Starting with $O^0 = \emptyset$, we get $I^1 = \text{OUT}[B_1] \cap O^0 = \emptyset$. However, if we start with $O^0 = U$, then we get $I^1 = \text{OUT}[B_1] \cap O^0 = \text{OUT}[B_1]$, as we should. Intuitively, the solution obtained starting with $O^0 = U$ is more desirable, because it correctly reflects the fact that expressions in $\text{OUT}[B_1]$ that are not killed by B_2 are available at the end of B_2. □

Algorithm 9 . 1 7 : Available expressions.

INPUT: A flow graph with e-kills and e.gens computed for each block B. The initial block is B1.

OUTPUT: IN [5] and O U T [5] , the set of expressions available at the entry and exit of each block B of the flow graph.

```
OUT[ENTRY] = ∅;
for (each basic block B other than ENTRY) OUT[B] = U;
while (changes to any OUT occur)
        for (each basic block B other than ENTRY) {
                IN[B] = ⋂_{P a predecessor of B} OUT[P];
                OUT[B] = e_gen_B ∪ (IN[B] − e_kill_B);
        }
```

Figure 9.20: Iterative algorithm to compute available expressions

Figure 9.20: Iterative algorithm to compute available expressions