卐 Compiler Short Revision Notes 卐

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Contents

1 Intro And ML Lex

$(a \odot b)|\epsilon$ represents the language {"", "ab"}. In writing regular expressions, we will sometimes omit the concatenation symbol or the epsilon, and we will assume that Kleene closure "binds tighter" than concatenation, and concatenation binds tighter than alternation; so that $ab|c$ means $(a \odot b)|c$, and $(a|)$ means $(a|\epsilon)$. Let us introduce some more abbreviations: [abed] means $(a|b|c|d)$, [b-g] means [bcdefg], [b-gM-Qkr] means [bcdefgMNOPQkr], M? means $(M|\epsilon)$, and M^+ means $(M \odot M^*)$.

```
if                                  (IF);
[a-z][a-z0-9]*                      (ID);
[0-9]+                              (NUM);
([0-9]+"."[0-9]*)|([0-9]*"."[0-9]+) (REAL);
("--"[a-z]*"\n")|(" "|"\n"|"\t")+   (continue());
.                                   (error(); continue());
```

FIGURE 2.2. Regular expressions for some tokens.

Longest match: The longest initial substring of the input that can match any regular expression is taken as the next token.
Rule priority: For a **particular** longest initial substring, the first regular expression that can match determines its token type. This means that the order of writing down the regular-expression rules has significance.
So according to the rules, if8 match as a single identifier and not as the two tokens if and 8. And "if 89" begin with a reserved word and not by an identifier by rule priority rule.

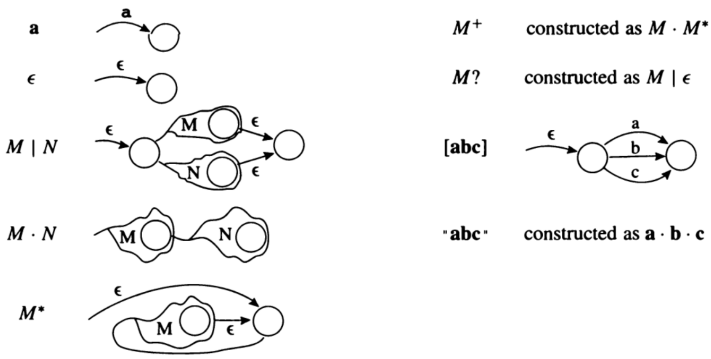


FIGURE 2.6. Translation of regular expressions to NFAs.

$$\textbf{DFAedge}(d, c) = \textbf{closure}(\bigcup_{s \in d} \textbf{edge}(s, c))$$

Using **DFAedge**, we can write the NFA simulation algorithm more formally. If the start state of the NFA is s_1, and the input string is $c_1, \ldots, c_k$, then the algorithm is:

```
d ← closure({s1})
for i ← 1 to k
  d ← DFAedge(d, ci)
```

Abstractly, there is an edge from d_i to d_j labeled with c if d_j = **DFAedge**(d_i, c). We let Σ be the alphabet.

```
states[0] ← {};     states[1] ← closure({s1})
p ← 1;     j ← 0
while j ≤ p
  foreach c ∈ Σ
    e ← DFAedge(states[j], c)
    if e = states[i] for some i ≤ p
      then trans[j, c] ← i
      else p ← p + 1
           states[p] ← e
           trans[j, c] ← p
  j ← j + 1
```

DFA construction is a mechanical task easily performed by computer, so it makes sense to have an automatic lexical analyzer generator to translate regular expressions into a DFA.
The output of ML-Lex is a program in ML - a lexical analyzer that interprets a DFA using the algorithm described in Section 2.3 and executes

the action fragments on each match. The action fragments are just ML statements that return token values.

```
(* ML Declarations: *)
    type lexresult = Tokens.token
    fun eof() = Tokens.EOF(0,0)
    %%
(* Lex Definitions: *)
    digits=[0-9]+
    %%
(* Regular Expressions and Actions: *)
    if                  => (Tokens.IF(yypos,yypos+2));
    [a-z][a-z0-9]*      => (Tokens.ID(yytext,yypos,yypos+size yytext));
    {digits}            => (Tokens.NUM(Int.fromString yytext,
                                yypos,yypos+size yytext));
    ({digits}"."[0-9]*)|([0-9]*"."{digits})
                        => (Tokens.REAL(Real.fromString yytext,
                                    yypos, yypos+size yytext));
    ("--"[a-z]*"\n")|(" "|"\n"|"\t")+
                        =>    (continue());
                        =>    (ErrorMsg.error yypos "illegal character";
                               continue());
```

PROGRAM 2.9. ML-Lex specification of the tokens from Figure 2.2.

The format is: user declarations %% ML-Lex definitions %% rules
The first part of the specification, above the first %% mark, contains functions and types written in ML. These must include the type lexresult, which is the result type of each call to the lexing function; and the function eof, which the lexing engine will call at end of file. This section can also contain utility functions for the use of the semantic actions in the third section. It is called with the same argument as lex (see %arg, below), and must return a value of type lexresult.
The second part of the specification contains regular-expression abbreviations and state declarations. For example, the declaration digits=$[0-9]^+$ in this section allows the name {digits} to stand for a nonempty sequence of digits within regular expressions.
In the definitions section, the user can define named regular expressions, a set of start states, and specify which of the various bells and whistles of ML-Lex are desired.
The third part contains regular expressions and actions. The actions are fragments of ordinary ML code. Each action must return a value of type lexresult. In this specification, lexresult is a token from the Tokens structure.
In the action fragments, several special variables are available. The string matched by the regular expression is yytext. The file position of the beginning of the matched string is yypos. The function continue () calls the lexical analyzer recursively.

In this particular example, each token is a data constructor parameterized by two integers indicating the position – in the input file – of the beginning and end of the token.

```
structure Tokens =
struct
   type pos = int
   datatype token = EOF of pos * pos
                  | IF of pos * pos
                  | ID of string * pos * pos
                  | NUM of int * pos * pos
                  | REAL of real * pos * pos
                    ⋮
end
```

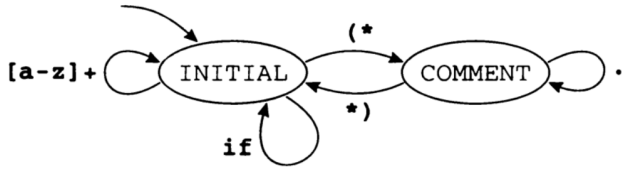
Arguments given to token are called payload.
The tokens are defined by the combined effect of
1. The %term commands used in the ML-Yacc declaration section of your ML-Yacc specification. These may add extra values to the token function's argument and thus extend the payload.
2. The lexresult type declaration in the user declarations of your ML-Lex specification
If a token has been defined by the %term command in the .yacc file with no type, then its payload is usually two integers - its the %pos declaration which says so, see chapter 9.4.3 on page 22. For example, looking at the SML/NJ compiler, we see that the semicolon is defined by the ML-Yacc %term command in file ml.grm as SEMICOLON. There is no type specification. The payload is two integers specifying the character positions in the source file of the start and end of the semicolon:

```
<INITIAL>";" => (Tokens.SEMICOLON(yypos,yypos+1));
```

If a token has been defined in ML-Yacc with a type, then its payload will be a value of that type, followed by two integers - again, its the %pos declaration which calls for those two integers, see chapter 9.4.3 on page 22.. For example, looking at the SML/NJ compiler, we see that a real number is defined by the ML-Yacc %term command in file ml.grm as REAL of string. The payload is therefore a string followed by two integers specifying the character position in the source file of the start and end of the real number:

```
<INITIAL>{real} => (Tokens.REAL(yytext,
yypos,
yypos+size yytext));
```

But sometimes the step-by-step, state-transition model of automata is appropriate. ML-Lex has a mechanism to mix states with regular expressions. One can declare a set of start states; each regular expression can be prefixed by the set of start states in which it is valid. The action fragments can explicitly change the start state. In effect, we have a finite automaton whose edges are labeled, not by single symbols, but by regular expressions. This example shows a language with simple identifiers, if tokens, and comments delimited by (* and *) brackets:



The ML-Lex specification corresponding to this machine is

the usual preamble ...

```
%%
%s COMMENT
%%
<INITIAL>if         => (Tokens.IF(yypos,yypos+2));
<INITIAL>[a-z]+     => (Tokens.ID(yytext,yypos,
                                  yypos+size(yytext)));
<INITIAL>"(*"       => (YYBEGIN COMMENT; continue());
<COMMENT>"*)"       => (YYBEGIN INITIAL; continue());
<COMMENT>.          => (continue();)
```

This example can be easily augmented to handle nested comments, via a global variable that is incremented and decremented in the semantic actions.
Any regular expression not prefixed by a $<$ *state* $>$ operates in all states; this feature is rarely useful.
Certain rules

- An individual character stands for itself, except for the reserved characters

  ```
  ? * + | ( ) ^ $ / ; . = < > [ { " \  $
  ```

- A backslash followed by one of the reserved characters stands for that character.
- Inside the brackets, only the symbols

  ```
  \ - ^
  ```

  are reserved. An initial up-arrow ^ stands for the complement of the characters listed, e.g. [^abc] stands any character except a, b, or c.
- To include ^ literally in a bracketed set, put it anywhere but first; to include - literally in a set, put it first or last.
- The dot . character stands for any character except newline, i.e. the same as

  ```
  [^\n]
  ```

- The following special escape sequences are available, inside or outside of square brackets:

  ```
  \b backspace
  \n newline
  \t horizontal tab
  \ddd where ddd is a 3 digit decimal escape
  ```

- Any regular expression may be enclosed in parentheses ( ) for syntactic (but, as usual, not semantic) effect
- A sequence of characters will stand for itself (reserved characters will be taken literally) if it is enclosed in double quotes " ".
- A postfix repetition range {a, b} where a and b are small integers stands for any number of repetitions between a and b of the preceding expression. The notation {a} stands for exactly a repetitions. Ex: [0-9]{3} Any three-digit decimal number.
- The rules should match all possible input. If some input occurs that does not match any rule, the lexer created by ML-Lex will raise an exception LexError.
- The user may recursively call the lexing function with lex(). (If %arg is used, the lexing function may be re-invoked with the same argument by using continue().) This is convenient for ignoring white space or comments silently:

  ```
  [\ \t\n]+       => ( lex());
  ```

- To switch start states, the user may call YYBEGIN with the name of a start state.

- If the lexer is to be used with the ML-Yacc parser, then additional glue declarations are needed:

```
5 structure T = Tokens
6 type pos = int (* Position in file *)
7 type svalue = T.svalue
8 type ('a,'b) token = ('a,'b) T.token
9 type lexresult = (svalue,pos) token
10 type lexarg = string
11 type arg = lexarg
12 val linep = ref 1; (* Line pointer *)
```

  Lines 5 through 9 provide the basic glue. On line 9, lexresult returns the type of the result returned by the rule actions. If you are passing a parameter to the lexer, then you also need the additional glue in lines 10 through 11. The lexer offers the possibility of counting lines using value yylineno described in chapter 7.3.6. If you prefer to do this yourself with variable linep, you will need the declaration on line 12
- Running ML - Lex file: From the Unix shell, run

```
sml-lex myfile.lex
The output file will be myfile.lex.sml. The
extension .lex is not required but is recommended.
```

To get messages for lexer errors and unwelcome characters (note: l1 is the lineno. and l2 is the position in that line):

```
val error : string * int * int -> unit = fn
(e,l1,l2) => TextIO.output(TextIO.stdOut,"lex:line "
^Int.toString l1^" l2="^Int.toString l2
^": "^e^"\n")
val badCh : string * string * int * int -> unit = fn
(fileName,bad,line,col) =>
TextIO.output(TextIO.stdOut,fileName^"["
^Int.toString line^"."^Int.toString col
^"] Invalid character \""^bad^"\"\n");
```

A typical error is to forget to close an ongoing comment. If you allow ML style nested comments (* ... (* ... *) ... *) then you will need some management of nested comments and possible end-of-file errors in the lexer.

```
21 val mlCommentStack : (string*int) list ref = ref [];
22 val eof = fn fileName =>
23 (if (!mlCommentStack)=[] then ()
24 else let val (file,line) = hd (!mlCommentStack)
25 in TextIO.output(TextIO.stdOut,
26 " I am surprized to find the
27 ^" end of file \""fileName^"\"\n"
28 ^" in a block comment which began"
29 ^" at "^file^"["^Int.toString line^"].\n")
30 end;
31 T.EOF(!linep,!linep));
```

It assumes that the ML-Lex command %arg, chapter 7.2.7, has been specified and the name of the source file fileName has been passed to the lexer, see line 417 on page 43. If this is not the case, then fileName is replaced by (). For this treatment of nested commands to work well, additional measures are needed for the ends of lines in the rules section 7.3
The ML-Lex definitions section provides the following commands. They are all terminated with a semicolon ;
Use the specified code to create a functor header for the lexer structure. For example, if you are using ML-Yacc and you have specified %name My in the ML-Yacc declarations:

```
65 %header (functor MyLexFun(structure Tokens: My_TOKENS));
```

This has the effect of turning what would have been a structure into a functor. The functor is needed for the glue code which integrates the lexer into a project. The SML/NJ compiler uses this technique with ML in place of My. Our working example also uses the technique with Pi in place of My. See lines 317 on page 40 and 391 on page 42. If you prefer to create the lexer as an SML/NJ structure, then omit this command and use the command %structure If you prefer to create your lexer as an SML/NJ structure rather than a functor, when for example you are not using ML-Yacc, then use the command %structure identifier to name the structure in the output program my.lex.sml as identifier instead of the default Mlex
%count counts newlines using yylineno
%posarg Pass an initial-position argument to function makeLexer. See 10.4. %arg An extra (curried) formal parameter argument is to be passed to the lex functions, and to the eof function in place of (). See 7.3.2. For example:

```
66 %arg (fileName:string);
```

The argument value is passed in the call to the parser. See line 415 on page 43
lex() and continue(): If %arg, chapter 7.2.7, is not used, you may recursively call the lexing function with lex().

```
82 [\ \t]+ => ( lex() );
```

For example, line 82 ignores spaces and tabs silently; However, if %arg is used, the lexing function may be re-invoked with the same argument by using continue().

```
83 <COMMENT>. => (continue());
```

For example, line 83 silently ignores all characters except a newline when the parser is in the user-defined state COMMENT
yylineno: The value yylineno is defined only if command %count has been specified, chapter 7.2.5. yylineno provides the current line number. 7.3.6.1 Warning This function should be used only if it is really needed. Adding the yylineno facility to a lexer will slow it down by 20%. It is much more efficient to recognise nn and have an action that increments a line-number variable. For example, see chapter 11.2.3 on page 38 in our working example.

```
datatype lexresult= DIV | EOF | EOS | ID of string | LPAREN |
NUM of int | PLUS | PRINT | RPAREN | SUB | TIMES

val linenum = ref 1
val error = fn x => output(std_out,x ^ "\n")
val eof = fn () => EOF
%%
%structure CalcLex
alpha=[A-Za-z];
digit=[0-9];
ws = [\ \t];
%%
\n       => (inc linenum; lex());
{ws}+    => (lex());
"/"      => (DIV);
";"      => (EOS);
"("      => (LPAREN);
(* revfold (((('a * 'b)->'b) ->'a list ->'b -> 'b) is like fold
done from backwards (in this case from left) explode will
convert the string to list of characters. *)
{digit}+ => (NUM (revfold (fn(a,r)=>ord(a)-ord("0")+10*r)
(explode yytext) 0));
")"      => (RPAREN);
"+"      => (PLUS);
{alpha}+ => (if yytext="print" then PRINT else ID yytext);
"-"      => (SUB);
"*"      => (TIMES);
.        => (error ("calc: ignoring bad character "^yytext);
lex());
```

2 Parsing

The parser returns an abstract syntax tree of the expression being evaluated. The parser gets tokens from the scanner to parse the input and build the AST. When an AST is returned by the parser, the compiler calls the code generator to evaluate the tree and produce the target code.
There are two main parts to a compiler, the front end and back end. The front end reads the tokens and builds an AST of a program. The back end generates the code given the AST representation of the program. As presented in earlier chapters, the front end consists of the scanner and the parser.

1 $S \rightarrow S\ ;\ S$	4 $E \rightarrow \text{id}$	8 $L \rightarrow E$
2 $S \rightarrow \text{id} := E$	5 $E \rightarrow \text{num}$	9 $L \rightarrow L\ ,\ E$
3 $S \rightarrow \text{print}\ (\ L\ )$	6 $E \rightarrow E + E$	
	7 $E \rightarrow (\ S\ ,\ E\ )$	

GRAMMAR 3.1. A syntax for straight-line programs.

As before, we say that a language is a set of strings; each string is a finite sequence of symbols taken from a finite alphabet. For parsing, the strings are source programs, the symbols are lexical tokens, and the alphabet is the set of token types returned by the lexical analyzer.
A leftmost derivation is one in which the leftmost nonterminal symbol is always the one expanded; in a rightmost derivation, the rightmost nonterminal is always next to be expanded.
A parse tree is made by connecting each symbol in a derivation to the one from which it was derived, as shown in Figure 3.3. Two different derivations can have the same parse tree.
A grammar is ambiguous if it can derive a sentence with two different parse trees.
Parsers must read not only terminal symbols such as +, -, num, and so on, but also the end-of-file marker. We will use $ to represent end of file.

Suppose S is the start symbol of a grammar. To indicate that $ must come after a complete S-phrase, we augment the grammar with a new start symbol S' and a new production S' → S$.

Predictive Parsing: Some grammars are easy to parse using a simple algorithm known as recursive descent. Predictive parsing works only on grammars where the first terminal symbol of each subexpression provides enough information to choose which production to use.

$S \to E\ \$$

$E \to E + T$
$E \to E - T$
$E \to T$

$T \to T * F$
$T \to T / F$
$T \to F$

$F \to \text{id}$
$F \to \text{num}$
$F \to (\ E\ )$

GRAMMAR 3.10.

$S \to \text{if } E \text{ then } S \text{ else } S$
$S \to \text{begin } S\ L$
$S \to \text{print } E$

$L \to \text{end}$
$L \to ;\ S\ L$

$E \to \text{num} = \text{num}$

GRAMMAR 3.11.

```
datatype token = IF | THEN | ELSE | BEGIN | END | PRINT
               | SEMI | NUM | EQ

val tok = ref (getToken())
fun advance() =  tok := getToken()
fun eat(t) = if (!tok=t) then advance() else error()

fun S() = case !tok
           of IF => (eat(IF); E(); eat(THEN); S();
                     eat(ELSE); S())
            | BEGIN => (eat(BEGIN); S(); L())
            | PRINT => (eat(PRINT); E())
and L() = case !tok
           of END => (eat(END))
            | SEMI => (eat(SEMI); S(); L())
and E() = (eat (NUM); eat(EQ); eat(NUM))
```

ith suitable definitions of `error` and `getToken`, this program will parse ry nicely.

Emboldened by success with this simple method, let us try it with Gram- ar 3.10:

```
fun S() = (E(); eat(EOF))
and E() = case !tok
           of ? => (E(); eat(PLUS); T())
            | ? => (E(); eat(MINUS); T())
            | ? => (T())
and T() = case !tok
           of ? => (T(); eat(TIMES); F())
            | ? => (T(); eat(DIV); F())
            | ? => (F())
and F() = case !tok
           of ID => (eat(ID))
            | NUM => (eat(NUM))
            | LPAREN => (eat(LPAREN); E(); eat(RPAREN))
```

Given a string γ of terminal and nonterminal symbols, FIRST(γ) is the set of all terminal symbols that can begin any string derived from γ. For example, let $\gamma = T * F$. Any string of terminal symbols derived from γ must start with `id`, `num`, or `(`. Thus,

FIRST($T * F$) = {`id`, `num`, `(`}.

If two different productions $X \to \gamma_1 and X \to \gamma_2$ have the same left- hand-side symbol (X) and their right-hand sides have overlapping FIRST sets, then the grammar cannot be parsed using predictive parsing.

With respect to a particular grammar, given a string γ of terminals and nonterminals,

- nullable(X) is true if X can derive the empty string.
- FIRST(γ) is the set of terminals that can begin strings derived from γ.
- FOLLOW(X) is the set of terminals that can immediately follow X. That is,

$t \in$ FOLLOW(X) if there is any derivation containing Xt. This can occur i the derivation contains $X\,Y\,Z\,t$ where Y and Z both derive ϵ.

A precise definition of FIRST, FOLLOW, and nullable is that they are the smallest sets for which these properties hold:

For each terminal symbol Z, FIRST[Z] = {Z}.
for each production $X \to Y_1 Y_2 \cdots Y_k$
  if $Y_1 \ldots Y_k$ are all nullable (or if $k = 0$)
    then nullable[X] = true
  for each i from 1 to k, each j from $i + 1$ to k
    if $Y_1 \cdots Y_{i-1}$ are all nullable (or if $i = 1$)
      then FIRST[X] = FIRST[X] ∪ FIRST[Y_i]
    if $Y_{i+1} \cdots Y_k$ are all nullable (or if $i = k$)
      then FOLLOW[Y_i] = FOLLOW[Y_i] ∪ FOLLOW[X]
    if $Y_{i+1} \cdots Y_{j-1}$ are all nullable (or if $i + 1 = j$)
      then FOLLOW[Y_i] = FOLLOW[Y_i] ∪ FIRST[Y_j]

The set of sym- bols is the union of the non-terminal and terminal sets. Each production has a semantic action associated with it. A production with a semantic action is called a rule. Parsers perform bottom-up, left-to-right evaluations of parse trees using semantic actions to compute values as they do so Given a production $P = A \to \alpha$, the corre- sponding semantic action is used to compute a value for A from the values of the symbols in α. If A has no value, the semantic action is still evaluated but the value is ignored. Each parse returns the value associated with the start symbol S of the grammar. A parse returns a nullary value if the start symbol does not carry a value.

An ML-Yacc specification consists of three parts, each of which is separated from the others by a %% delimiter. The general format is: ML-Yacc user declarations %% ML-Yacc declarations %% ML-Yacc rules Comments have the same lexical definition as they do in Standard ML and can be placed in any ML-Yacc section

After the first %%, the following words and symbols are reserved:

```
of for = { } , * -> :  | ( )
```

code: This class is meant to hold ML code. The ML code is not parsed for syntax errors. It consists of a left parenthesis followed by all characters up to a balancing right parenthesis. Parentheses in ML comments and ML strings are excluded from the count of balancing parentheses.

In ML-YACC user declarations section you can define values available in the semantic actions of the rules in the user declarations section. It is recommended that you keep the size of this section as small as possible and place large blocks of code in other modules.

The ML-Yacc declarations section is used to make a set of required declarations and a set of optional declarations.

You must declare the type of basic payload position values using the %pos declaration. The syntax is %pos ML-type . This type MUST be the same type as that which is actually found in the lexer. It cannot be polymorphic.

You may declare whether the parser generator should create a verbose description of the parser in a ".desc" file. This is useful for debugging your parser and for finding the causes of shift/reduce errors and other parsing conflicts.

You may also declare whether the semantic actions are free of significant side-effects and always terminate. Normally, ML-Yacc delays the evaluation of semantic actions until the completion of a successful parse. This ensures that there will be no semantic actions to "undo" if a syntactic error-correction invalidates some semantic actions. If, however, the semantic actions are free of significant side-effects and always terminate, the results of semantic actions that are invalidated by a syntactic error-correction can always be safely ignored.

Parsers run faster and need less memory when it is not necessary to delay the eval- uation of semantic actions. You are encouraged to write semantic actions that are free of side-effects and always terminate and to declare this information to ML-Yacc.

A semantic action is free of significant side-effects if it can be re-executed a reasonably small number of times without affecting the result of a parse. (The re-execution occurs when the error-correcting parser is testing possible corrections to fix a syntax error, and the number of times re-execution occurs is roughly bounded, for each syntax error, by the number of terminals times the amount of lookahead permitted for the error-correcting parser).

You must specify the name of the parser with command %name name . If you decide to call your parser " *My* Parser " then you will need the declaration: 87 %name My This declaration must agree with the ML-Lex command %header

You must define the terminal and non-terminal sets using the %term and %nonterm declarations, respectively. These declarations are like an ML datatype definition. The types cannot be polymorphic. Do not use any locally defined types from the user declarations section of the specification

Terminals are written in Capitals whereas non terminals are written in small.

Consider

%nonterm elabel of (symbol * exp)

Here since the supplemented payload is a tuple, parenthesis is required.

You may want each invocation of the entire parser to be parameterised by a particular argument, such as the file name of the input being parsed in an invocation of the parser. The %arg declaration allows you to specify such an argument. (This is often cleaner than using "global" reference variables.) The declaration %arg Any-ML-pattern : ML-type specifies the argument to the parser, as well as its type. If %arg is not specified, it defaults to () : unit . Note that ML-Lex also has a %arg directive, but the two are independent and may have different types. For example:

107 %arg (fileName) : string

You should specify the set of terminals that may follow the start symbol, also called end-of-parse symbols, using the %eop declaration. The %eop keyword should be followed by the list of terminals. This is useful, for example, in an interactive system where you want to force the evaluation of a statement before an end-of-file (remember, a parser delays the execution of semantic actions until a parse is successful). ML-Yacc has no concept of an end-of-file. You must define an end-of-file terminal (EOF , perhaps) in the %term declaration. You must declare terminals which cannot be shifted, such as end-of-file, in the %noshift declaration. The %noshift keyword should be followed by the list of non-shiftable terminals. An error message will be printed if a non-shiftable terminal is found on the right hand side of any rule, but ML-Yacc will not prevent you from using such grammars.

You should list the precedence declarations in order of increasing (tighter-binding) prece- dence. Each precedence declaration consists of a % keyword specifying associativity followed by a list of terminals. You may place more than one terminal at a given prece- dence level, but you cannot specify non-terminals. The keywords are %left , %right , and %nonassoc , standing for their respective associativities.

The %nodefault declaration suppresses the generation of default reductions

Include the %pure declaration if the semantic actions are free of significant side effects and always terminate. It is suggested that you begin developing your language without this directive

You may define the start symbol using the %start declaration. Otherwise the non- terminal for the first rule will be used as the start non-terminal. The keyword %start should be followed by the name of the starting non-terminal. This non-terminal should not be used on the right hand side of any rules, to avoid conflicts between reducing to the start symbol and shifting a terminal.

Include the %verbose declaration to produce a verbose description of the LALR parser. The name of this file is the name of the specification file with a " .desc " appended to it, for example pi.yacc.desc . This file is helpful for debugging, and has the following format:

1. A summary of errors found while generating the LALR tables.

2. A detailed description of all errors.

3. A description of the states of the parser. Each state is preceded by a list of conflicts in the state.

Specify all keywords in your grammar here. The %keyword should be followed by a list of terminal names

List terminals to prefer for insertion after the command %prefer . Corrections which insert a terminal on this list will be chosen over other corrections, all other things being equal.

The error-correction algorithm may also insert terminals with values. You must supply a value for such a terminal. The keyword should be followed by a terminal and a piece of code (enclosed in parentheses) that when evaluated supplies the value. There must be a separate %value declaration for each terminal with a value that you wish may be inserted or substituted in an error correction. The code for the value is not evaluated until the parse is successful.

ML-YACC Rules

The rules section contains the context-free grammar productions and their associated semantic actions. A rule consists of a left hand side non-terminal, followed by a colon, followed by a list of right hand side clauses. The right hand side clauses should be separated by bars. Each clause consists of a list of non-terminal and terminal symbols, followed by an optional %prec declaration, and then followed by the code to be evaluated when the rule is reduced. The optional %prec consists of the keyword %prec followed by a terminal whose precedence should be used as the precedence of the rule. The values of those symbols in a right hand side clause which have values are available inside the code

```
141   path: IDE   ((Name (IDE,fileName,IDEleft,IDEright)
```

Each position value has the general form { symbol name }{ n+1 } , where { n } is the number of occurrences of the symbol to the left of the symbol. If the symbol occurs only once in the rule, { symbol name } may also be used. For example, if in rule " path " above, there had been two IDE 's in the list of symbols, we could have referred to their values as IDE1 and IDE2 .

Positions for all the symbols are also available. The payload positions are given by { symbol name }{ n+1 } left and { symbol name }{ n+1 } right . where { n } is defined as before. For example we see the use of IDEleft and IDEright on line 141. If in rule " path " above, there had been two IDE 's in the list of symbols, we could have referred to their left and right positions as IDE1left , IDE1right , IDE2left and IDE2right .

The position for a null right-hand-side of a production is assumed to be the leftmost position of the lookahead terminal which is causing the reduction. This position value is available in defaultPos

The value to which the code evaluates is used as the value of the non-terminal. The type of the value and the non-terminal must match. The value is ignored if the non-terminal has no value, but is still evaluated for side-effects.

```
142 datatype Life = Life of Year * Year      * int * int
143 and Year = Year of int * int * int  * int * int
```

The two integers at the end of each of lines 142 and 143 give the position of the con- structions in the source file. They will be needed for error messages. The tokens representing years, months and days will have a supplemental payload of one integer; they are declared in the ML-Yacc declarations section of the file my.yacc as:

```
144 %term YMD of int
```

This means that the tokens YMD generated by the lexer will have a payload of type int * int * int . The lexer rules in file my.lex might be

```
145 {int} => (T.YMD(stoi yytext,!line,!col))
```

where function stoi : string -¿ int converts a string of digits to the corresponding integer. Parser rules in file my.yacc will pull these three integers together to form the two years and the life:

```
146 life: year HYPHEN year
147 ((Life (year1,year2,year1left,year1right)))
148 year: YMD COLON YMD COLON YMD
149 ((Year (YMD1,YMD2,YMD3,YMD1left,YMD1right)
```

Add %prec declarations to the rules to say which terminal's precedence and as- sociativity are to be used with which rules

The language specified by this project requires that function application be left associative as in ML, i.e. f g 2 means (f g) 2 . Now function application is a non-terminal, fun_appl , in the ML-Yacc specification, and non-terminals cannot be placed in the %left , %right and %nonassoc declarations — what can we do? The solution is to create a new terminal, FUN_APPL , by declaring it in the %term declaration, and then defining the required precedence and associativity on line

```
170.  The corresponding rule becomes:
175 fun_appl:   (* Function application has precedence
176 and associativity defined by dummy
177 terminal FUN_APPL. *)
178 name e     %prec FUN_APPL  ((name,e))
```

Error recovery in predictive parsing (LL (1)):

A syntax error occurs when the string of input tokens is not a sentence in the language. Error recovery is a way of finding some sentence similar to that string of tokens. This can proceed by deleting, replacing, or inserting tokens.

For example, error recovery for T could proceed by inserting a num token. It's not necessary to adjust the actual input; it suffices to pretend that the num was there, print a message, and return normally.

```
and T()  = case !tok
             of ID     => (F(); T'())
              | NUM    => (F(); T'())
              | LPAREN => (F(); T'())
              | _      => print("expected id, num,
                                 or left-paren")
```

It's a bit dangerous to do error recovery by insertion, because if the error cascades to produce another error, the process might loop infinitely. Error recovery by deletion is safer, because the loop must eventually terminate when end-of-file is reached.

Simple recovery by deletion works by skipping tokens until a token in the FOLLOW set is reached. For example, error recovery for T' could work like this:

```
and T'() = case !tok
             of PLUS   => ()
              | TIMES  => (eat(TIMES); F(); T'())
              | RPAREN => ()
              | EOF    => ()
              | _      => (print "expected +, *, right-paren,
                                 or end-of-file";
                           skipto[PLUS,TIMES,RPAREN,EOF])
and skipto(stop) =
      if member(!tok, stop) then ()
      else (eat(!tok); skipto(stop))
```

2.1 LR(0)

Closure(I) =
  repeat
    for any item $A \to \alpha.X\beta$ in I
      for any production $X \to \gamma$
        $I \leftarrow I \cup \{X \to .\gamma\}$
  until I does not change.
  return I

Goto(I, X) =
  set J to the empty set
  for any item $A \to \alpha.X\beta$ in I
    add $A \to \alpha X.\beta$ to J
  return Closure(J)

Note: 'X' in Goto can be either a terminal or non terminal

Now here is the algorithm for LR(0) parser construction. First, augment the grammar with an auxiliary start production $S' \to S\$$. Let T be the set of states seen so far, and E the set of (shift or goto) edges found so far.

Initialize T to {**Closure**($\{S' \to .S\$\}$)}
Initialize E to empty.
repeat
  for each state I in T
    for each item $A \to \alpha.X\beta$ in I
      let J be **Goto**(I, X)
      $T \leftarrow T \cup \{J\}$
      $E \leftarrow E \cup \{I \xrightarrow{X} J\}$
until E and T did not change in this iteration

However, for the symbol \$ we do not compute **Goto**(I, \$); instead we will make an **accept** action.

For Grammar 3.20 this is illustrated in Figure 3.21.

Now we can compute set R of LR(0) reduce actions:

$R \leftarrow \{\}$
for each state I in T
  for each item $A \to \alpha.$ in I
    $R \leftarrow R \cup \{(I, A \to \alpha)\}$

We can now construct a parsing table for this grammar (Table 3.22). For each edge $I \xrightarrow{X} J$ where X is a terminal, we put the action *shift* J at position (I, X) of the table; if X is a nonterminal we put *goto* J at position (I, X). For each state I containing an item $S' \to S.\$$ we put an *accept* action at $(I, \$)$. Finally, for a state containing an item $A \to \gamma.$ (production n with the dot at the end), we put a *reduce* n action at (I, Y) for every token Y.

2.2 SLR

$R \leftarrow \{\}$
for each state I in T
  for each item $A \to \alpha.$ in I
    for each token X in FOLLOW(A)
      $R \leftarrow R \cup \{(I,\ X,\ A \to \alpha)\}$

The action $(I, X, A \to \alpha)$ indicates that in state I, on lookahead symbol X, the parser will reduce by rule $A \to \alpha$.

2.3 LR(1)

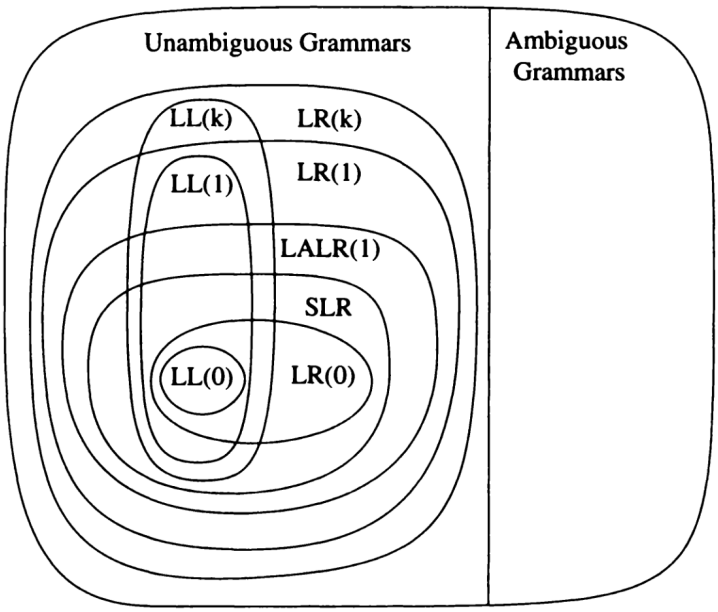
Closure(I) =
  repeat
    for any item $(A \to \alpha.X\beta,\ z)$ in I
      for any production $X \to \gamma$
        for any $w \in$ FIRST(βz)
          $I \leftarrow I \cup \{(X \to .\gamma,\ w)\}$
  until I does not change
  return I

Goto(I, X) =
  $J \leftarrow \{\}$
  for any item $(A \to \alpha.X\beta,\ z)$ in I
    add $(A \to \alpha X.\beta,\ z)$ to J
  return Closure(J).

The start state is the closure of the item $(S' \to .S\ \$,\ ?)$, where the lookahead symbol ? will not matter, because the end-of-file marker will never be shifted.

The reduce actions are chosen by this algorithm:

$R \leftarrow \{\}$
for each state I in T
  for each item $(A \to \alpha.\ ,\ z)$ in I
    $R \leftarrow R \cup \{(I, z, A \to \alpha)\}$



In examining a shift-reduce conflict such as

$E \to E * E.$	$+$
$E \to E.\ +\ E$	*(any)*

there is the choice of shifting a *token* and reducing by a *rule*. Should the rule or the token be given higher priority? The precedence declarations (`%left`, etc.) give priorities to the tokens; the priority of a rule is given by the last token occurring on the right-hand side of that rule. Thus the choice here is between a rule with priority * and a token with priority +; the rule has higher priority, so the conflict is resolved in favor of reducing.

When the rule and token have equal priority, then a `%left` precedence favors reducing, `%right` favors shifting, and `%nonassoc` yields an error action.

Instead of using the default "rule has precedence of its last token," we can assign a specific precedence to a rule using the `%prec` directive. This is commonly used to solve the "unary minus" problem. In most programming languages a unary minus binds tighter than any binary operator, so $-6 * 8$ is parsed as $(-6) * 8$, not $-(6 * 8)$. Grammar 3.35 shows an example.

The token UMINUS is never returned by the lexer; it is merely a placeholder in the chain of precedence (`%left`) declarations. The directive `%prec UMINUS` gives the rule `exp: MINUS exp` the highest precedence, so reducing by this rule takes precedence over shifting any operator, even a minus sign.

```
%%
%term INT | PLUS | MINUS | TIMES | UMINUS | EOF
%nonterm  exp
%start exp
%eop EOF

%left PLUS MINUS
%left TIMES
%left UMINUS
%%

exp :  INT              ()
    |  exp PLUS exp     ()
    |  exp MINUS exp    ()
    |  exp TIMES exp    ()
    |  MINUS exp    %prec UMINUS  ()
```

GRAMMAR 3.35.

Popping states from the stack can lead to seemingly "im
possible" semantic actions, especially if the actions contain side effects. Consider this grammar fragment:

```
statements:  statements exp SEMICOLON
          |  statements error SEMICOLON
          |  (* empty *)

exp : increment exp decrement
    | ID

increment:  LPAREN     (nest := nest + 1)
decrement:  RPAREN     (nest := nest - 1)
```

"Obviously" it is true that whenever a semicolon is reached, the value of `nest` is zero, because it is incremented and decremented in a balanced way according to the grammar of expressions. But if a syntax error is found after some left parentheses have been parsed, then states will be popped from the stack without "completing" them, leading to a nonzero value of `nest`. The best solution to this problem is to have side-effect-free semantic actions that build abstract syntax trees, as described in Chapter 4.

Global error repair finds the smallest set of insertions and deletions that would turn the source string into a syntactically correct string, *even if the insertions and deletions are not at a point where an LL or LR parser would first report an error*. In this case, global error repair would do a single-token substitution, replacing `type` by `var`.

Read page 79, 80, 81 from text.

3 AST

However, an LR parser does perform reductions, and associated semantic actions, in a deterministic and predictable order: a bottom-up, left-to-right traversal of the parse tree. In other words, the (virtual) parse tree is traversed in postorder.

It is possible to write an entire compiler that fits within the semantic action phrases of an ML-Yacc parser. However, such a compiler is difficult to read and maintain. And this approach constrains the compiler to analyze the program in exactly the order it is parsed.

Technically, a parse tree has exactly one leaf for each token of the input and one internal node for each grammar rule reduced during the parse. Such a parse tree, which we will call a concrete parse tree representing the concrete syntax of the source language, is inconvenient to use directly. Many of the punctuation tokens are redundant and convey no information - they are useful in the input string, but once the parse tree is built, the structure of the tree conveys the structuring information more conveniently. Furthermore, the structure of the parse tree depends too much on the grammar! The grammar transformations shown in Chapter 3 - factoring, elimination of left recursion, elimination of ambiguity - involve the introduction of extra nonterminal symbols and extra grammar productions for technical purposes. These details should be confined to the parsing phase and should not clutter the semantic analysis.

An abstract syntax makes a clean interface between the parser and the later phases of a compiler (or, in fact, for the later phases of other kinds of program-analysis tools such as dependency analyzers). The abstract syntax tree conveys the phrase structure of the source program, with all parsing issues resolved but without any semantic interpretation.

So the parser uses the concrete syntax to build a parse tree for the abstract syntax. The semantic analysis phase takes this abstract syntax tree; it is not bothered by the ambiguity of the grammar, since it already has the parse tree!

The lexer must pass the source-file positions of the beginning and end of each token to the parser. The ML-Yacc parser makes these positions available (in the semantic action fragments): for each terminal or nonterminal such as FOO or FOOl on the right-hand side of a rule, the ML variable FOOlef t or FOOllef t stands for the left-end position of the terminal or nonterminal, and FOOright or FOOl right stands for the right-end position.

The Tiger language treats adjacent function declarations as (possibly) mutually recursive. The FunctionDec constructor of the abstract syntax takes a list of function declarations, not just a single function. The intent is that this list is a maximal consecutive sequence of function declarations. Thus, functions declared by the same FunctionDec can be mutually recursive.

The TypeDec constructor also takes a list of type declarations, for the same reason;

There is no abstract syntax for "&" and "|" expressions; instead, $e_1 \& e_2$ is translated as if e_1 then e_2 else 0, and $e_1 | e_2$ is translated as though it had been written if $e_1 then 1 else e_2$ Similarly, unary negation (-x) should be represented as subtraction (0 - x) in the abstract syntax. Also, where the body of a Let Exp has multiple statements, we must use a SeqExp. An empty statement is represented by SeqExp[].

The semantic analysis phase of the compiler will need to keep track of which local variables are used from within nested functions. The escape component of a VarDec, ForExp, or formal parameter is used to keep track of this. The parser should leave each escape set to true, which is a conservative approximation. The field type is used for both formal parameters and record fields; escape has meaning for formal parameters, but for record fields it can be ignored. Having the escape fields in the abstract syntax is a "hack," since escaping is a global, nonsyntactic property. But leaving escape out of the Absyn would require another data structure for describing escapes.

4 Semantic Analysis

This phase is characterized by the maintenance of symbol tables (also called environments) mapping identifiers to their types and locations.

An environment is a set of *bindings* denoted by the $\mapsto$ arrow. For example, we could say that the environment σ_0 contains the bindings $\{\texttt{g} \mapsto \texttt{string}, \texttt{a} \mapsto \texttt{int}\}$; meaning that the identifier `a` is an integer variable and `g` is a string variable.

```
1        function f(a:int, b:int, c:int) =
2           (print_int(a+c);
3            let var j := a+b
4                var a := "hello"
5             in print(a); print_int(j)
6            end;
7            print_int(b)
8           )
```

Suppose we compile this program in the environment σ_0. The formal parameter declarations on line 1 give us the table σ_1 equal to $\sigma_0 + \{\texttt{a} \mapsto \texttt{int}, \texttt{b} \mapsto \texttt{int}, \texttt{c} \mapsto \texttt{int}\}$, that is, σ_0 extended with new bindings for a, b, and c. The identifiers in line 2 can be looked up in σ_1. At line 3, the table $\sigma_2 = \sigma_1 + \{\texttt{j} \mapsto \texttt{int}\}$ is created; and at line 4, $\sigma_3 = \sigma_2 + \{\texttt{a} \mapsto \texttt{string}\}$ is created.

How does the + operator for tables work when the two environments being "added" contain different bindings for the same symbol? When σ_2 and $\{\texttt{a} \mapsto \texttt{string}\}$ map a to int and string, respectively? To make the scoping rules work the way we expect them to in real programming languages, we want $\{\texttt{a} \mapsto \texttt{string}\}$ to take precedence. So we say that $X + Y$ for tables is not the same as $Y + X$; bindings in the right-hand table override those in the left.

Finally, in line 6 we discard σ_3 and go back to σ_1 for looking up the identifier b in line 7. And at line 8, we discard σ_1 and go back to σ_0.

```
structure M = struct
   structure E = struct
      val a = 5;
   end
   structure N = struct
      val b = 10
      val a = E.a + b
   end
   structure D = struct
      val d = E.a + N.a
   end
end
```

(a) An example in ML

```
package M;
class E {
    static int a = 5;
}
class N {
    static int b = 10;
    static int a = E.a + b;
}
class D {
    static int d = E.a + N.a;
}
```

(b) An example in Java

FIGURE 5.1. Several active environments at once.

MULTIPLE SYMBOL TABLES

$$\sigma_1 = \{a \mapsto int\}$$

$$\sigma_2 = \{E \mapsto \sigma_1\}$$

$$\sigma_3 = \{b \mapsto int, a \mapsto int\}$$

$$\sigma_4 = \{N \mapsto \sigma_3\}$$

$$\sigma_5 = \{d \mapsto int\}$$

$$\sigma_6 = \{D \mapsto \sigma_5\}$$

$$\sigma_7 = \sigma_2 + \sigma_4 + \sigma_6$$

5 Activation Records

For the remainder of this chapter we will consider languages with stackable local variables and postpone discussion of higher-order functions to Chapter 15.

In many languages (including C, Pascal, and Tiger), local variables are destroyed when a function returns. Since a function returns only after all the functions it has called have returned, we say that function calls behave in last-in-first-out (LIFO) fashion. If local variables are created on function entry and destroyed on function exit, then we can use a LIFO data structure – a stack – to hold them.

```
fun f(x) =
  let fun g(y) = x+y
   in g
  end

val h = f(3)
val j = f(4)

val z = h(5)
val w = j(7)
```

(a) Written in ML

```
int (*)() f(int x) {
  int g(int y) {return x+y;}
  return g;
}

int (*h)() = f(3);
int (*j)() = f(4);

int z = h(5);
int w = j(7);
```

(b) Written in pseudo-C

PROGRAM 6.1. An example of higher-order functions.

But in languages supporting both nested functions *and* function-valued variables, it may be necessary to keep local variables after a function has returned!

It is the combination of *nested functions* (where inner functions may use variables defined in the outer functions) and *functions returned as results* (or stored into variables) that causes local variables to need lifetimes longer than their enclosing function invocations.

Pascal (and Tiger) have nested functions, but they do not have functions as returnable values. C has functions as returnable values, but not nested functions. So these languages can use stacks to hold local variables.

ML, Scheme, and several other languages have both nested functions and functions as returnable values (this combination is called *higher-order functions*). So they cannot use stacks to hold all local variables. This complicates

Instead, we treat the stack as a big array, with a special register – the *stack pointer* – that points at some location. All locations beyond the stack pointer are considered to be garbage, and all locations before the stack pointer are considered to be allocated. The stack usually grows only at the entry to a function, by an increment large enough to hold all the local variables for that function, and, just before the exit from the function, shrinks by the same amount. The area on the stack devoted to the local variables, parameters, return address, and other temporaries for a function is called the function's *activation record* or *stack frame*. For historical reasons, run-time stacks usually start at a high memory address and grow toward smaller addresses. This can be rather confusing: stacks grow downward and shrink upward, like icicles.

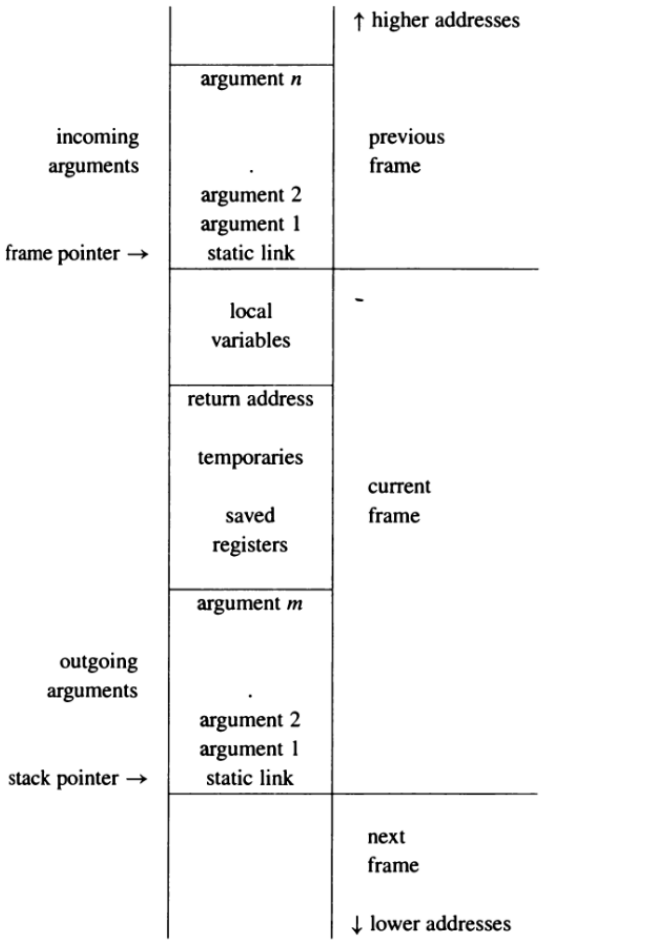


FIGURE 6.2. A stack frame.

Figure 6.2 shows a typical stack frame layout. The frame has a set of incoming arguments (technically these are part of the previous frame but they are at a known offset from the frame pointer) passed by the caller. The return address is created by the CALL instruction and tells where (within the calling function) control should return upon completion of the current function. Some local variables are in this frame; other local variables are kept in machine registers. Sometimes a local variable kept in a register needs to be saved into the frame to make room for other uses of the register; there is an area in the frame for this purpose. Finally, when the current function calls other functions, it can use the outgoing argument space to pass parameters.

Suppose a function g(...) calls the function f($a_1, \ldots, a_n$). We say g is the caller and f is the callee. On entry to f, the stack pointer points to the first argument that g passes to f. On entry, f allocates a frame by simply subtracting the frame size from the stack pointer SP. The old SP becomes the current frame pointer FP.

FP is a "fictional" register whose value is always SP+framesize. Why talk about a frame pointer at all? Why not just refer to all variables, parameters, etc. by their offset from SP, if the frame size is constant? The frame size is not known until quite late in the compilation process, when the number of memory-resident temporaries and saved registers is determined.

Suppose a function f is using register r to hold a local variable and calls procedure g, which also uses r for its own calculations. Then r must be saved (stored into a stack frame) before g uses it and restored (fetched back from the frame) after g is finished using it. But is it f's responsibility to save and restore the register, or g's? We say that r is a caller-save register if the caller (in this case, f) must save and restore the register, and r is callee-save if it is the responsibility of the callee (in this case, g).

therefore the first k arguments (for k = 4 or k = 6, typically) of a function are passed in registers $r_p, \ldots, r_{p+k-1}$ and the rest of the arguments are passed in memory.

Now, suppose $f(a_1, \ldots, a_n)$ (which received its parameters in $r_1, \ldots, r_n$, for example) calls $h(z)$. It must pass the argument z in r_1; so f saves the old contents of r_1 (the value a_1) in its stack frame before calling h. But there is the memory traffic that was supposedly avoided by passing arguments in registers! How has the use of registers saved any time?

There are four answers, any or all of which can be used at the same time:

1. **Some procedures don't call other procedures – these are called *leaf* procedures. What proportion of procedures are leaves? Well, if we make the (optimistic) assumption that the average procedure calls either no other procedures or calls at least two others, then we can describe a "tree" of procedure calls in which there are more leaves than internal nodes. This means that *most* procedures called are leaf procedures.**

   Leaf procedures need not write their incoming arguments to memory. In fact, often they don't need to allocate a stack frame at all. This is an important savings.
2. **Some optimizing compilers use *interprocedural register allocation*, analyzing all the functions in an entire program at once. Then they assign different procedures different registers in which to receive parameters and hold local variables. Thus $f(x)$ might receive x in r_1, but call $h(z)$ with z in r_7.**
3. **Even if f is not a leaf procedure, it might be finished with all its use of the argument x by the time it calls h (technically, x is a dead variable at the point where h is called). Then f can overwrite r_1 without saving it.**
4. **Some architectures have *register windows*, so that each function invocation can allocate a fresh set of registers without memory traffic.**

If f needs to write an incoming parameter into the frame, where in the frame should it go? Ideally, f's frame layout should matter only in the implementation of f. A straightforward approach would be for the caller to pass arguments $a_1, \ldots, a_k$ in registers and $a_{k+1}, \ldots, a_n$ at the end of its own frame – the place marked *outgoing arguments* in Figure 6.2 – which become the *incoming arguments* of the callee. If the callee needed to write any of these arguments to memory, it would write them to the area marked *local variables*.

To resolve the contradiction that parameters are passed in registers, but have addresses too, the first k parameters are passed in registers; but any parameter whose address is taken must be written to a memory location on entry to the function. To satisfy printf, the memory locations into which register arguments are written must all be consecutive with the memory locations in which arguments k + 1, k + 2, etc. are written. Therefore, C programs can't have some of the arguments saved in one place and some saved in another - they must all be saved contiguously.

When function g calls function f, eventually f must return. It needs to know where to go back to. If the call instruction within g is at address a, then (usually) the right place to return to is a + 1, the next instruction in g. This is called the return address.

On modern machines, the call instruction merely puts the return address (the address of the instruction after the call) in a designated register. A non- leaf procedure will then have to write it to the stack (unless interprocedural register allocation is used), but a leaf procedure will not.

Many of the local variables will be allocated to registers, as will the intermediate results of expression evaluation. Values are written to memory (in the stack frame) only when necessary for one of these reasons:

- **the variable will be passed by reference, so it must have a memory address (or, in the C language the & operator is anywhere applied to the variable);**
- **the variable is accessed by a procedure nested inside the current one;[2]**
- **the value is too big to fit into a single register;[3]**
- **the variable is an array, for which address arithmetic is necessary to extract components;**
- **the register holding the variable is needed for a specific purpose, such as parameter passing (as described above), though a compiler may move such values to other registers instead of storing them in memory;**
- **or there are so many local variables and temporary values that they won't all fit in registers, in which case some of them are "spilled" into the frame.**

We will say that a variable escapes if it is passed by reference, its address is taken (using C's & operator), or it is accessed from a nested function.

Unfortunately, the conditions in our list don't manifest themselves early enough. When the compiler first encounters the declaration of a variable, it doesn't yet know whether the variable will ever be passed by reference, accessed in a nested procedure, or have its address taken; and doesn't know how many registers the calculation of expressions will require An industrial-strength compiler must assign provisional locations to all formals and locals, and decide later which of them should really go in registers.

In languages that allow nested function declarations, the inner functions may use variables declared in outer functions. This language feature is called block structure.

There are several methods to allow inner functions to access outer ones:

- Whenever a function f is called, it can be passed a pointer to the frame of the function statically enclosing f; this pointer is the *static link*.
- A global array can be maintained, containing – in position i – a pointer to the frame of the most recently entered procedure whose *static nesting depth* is i. This array is called a *display*.
- When g calls f, each variable of g that is actually accessed by f (or by any function nested inside f) is passed to f as an extra argument. This is called *lambda lifting*.

I will describe in detail only the method of static links.

```
1       type tree = {key: string, left: tree, right: tree}
2
3       function prettyprint(tree: tree) : string =
4        let
5            var output := ""
6
7            function write(s: string) =
8                output := concat(output,s)
9
10           function show(n:int, t: tree) =
11              let function indent(s: string) =
12                      (for i := 1 to n
13                        do write(" ");
14                       output := concat(output,s); write("\n"))
15               in if t=nil then indent(".")
16                  else (indent(t.key);
17                        show(n+1,t.left);
18                        show(n+1,t.right))
19              end
20
21        in show(0,tree); output
22       end
```

ROGRAM 6.3. Nested functions.

Line #
21 `prettyprint` calls `show`, passing `prettyprint`'s own frame pointer as `show`'s static link.

10 `show` stores its static link (the address of `prettyprint`'s frame) into its own frame.
15 `show` calls `indent`, passing its own frame pointer as `indent`'s static link.
17 `show` calls `show`, passing its own static link (not its own frame pointer) as the static link.
12 `indent` uses the value n from `show`'s frame. To do so, it fetches at an appropriate offset from `indent`'s static link (which points at the frame of `show`).
13 `indent` calls `write`. It must pass the frame pointer of `prettyprint` as the static link. To obtain this, it first fetches at an offset from its own static link (from `show`'s frame), the static link that had been passed to `show`.
14 `indent` uses the variable `output` from `prettyprint`'s frame. To do so it starts with its own static link, then fetches `show`'s, then fetches `output`.[4]

```
signature FRAME =
sig type frame
    type access
    val newFrame : {name: Temp.label,
                    formals: bool list} -> frame
    val name : frame -> Temp.label
    val formals : frame -> access list
    val allocLocal : frame -> bool -> access
        :
end
```

The type frame holds information about formal parameters and local variables allocated in this frame. To make a new frame for a function f with k formal parameters, call newFrame {name=f formals=l}, where l is a list of k booleans: true for each parameter that escapes and false for each parameter that does not. The result will be a frame object.

The access type describes formals and locals that may be in the frame or in registers.

```
structure MipsFrame : FRAME = struct
        :
    datatype access = InFrame of int | InReg of Temp.temp
```

`InFrame(X)` indicates a memory location at offset X from the frame pointer; `InReg`(t_{84}) indicates that it will be held in "register" t_{84}. `Frame.access` is an abstract data type, so outside of the module the `InFrame` and `InReg` constructors are not visible. Other modules manipulate accesses using interface functions to be described in the next chapter.

The Frame. formals interface function extracts a list of it "accesses" denoting the locations where the formal parameters will be kept at run time, as seen from inside the callee. Parameters may be seen differently by the caller. For example, if parameters are passed on the stack, the caller may put a parameter at offset 4 from the stack pointer, but the callee sees it at offset 4 from the frame pointer. Or the caller may put a parameter into register 6, but the callee may want to move it out of the way and always access it from register 13.
because of this "shift of view" newFrame must calculate two things: • How the parameter will be seen from inside the function (in a register, or in a frame location); • What instructions must be produced to implement the "view shift."
Some local variables are kept in the frame; others are kept in registers. To allocate a new local variable in a frame f, the semantic analysis phase calls Frame.allocLocal(f)(true). This returns an InFrame access with an offset from the frame pointer. The boolean argument to allocLocal specifies whether the new variable escapes and needs to go in the frame; if it is false, then the variable can be allocated in a register. Thus, allocLocal (f) (false) might create InReg(t_{481}).
We use the word temporary to mean a value that is temporarily held in a register, and the word label to mean some machine-language location whose exact address is yet to be determined.

6 Translation To Intermediate Code

The front end of the compiler does lexical analysis, parsing, semantic analysis, and translation to intermediate representation. The back end does optimization of the intermediate representation and translation to machine language.

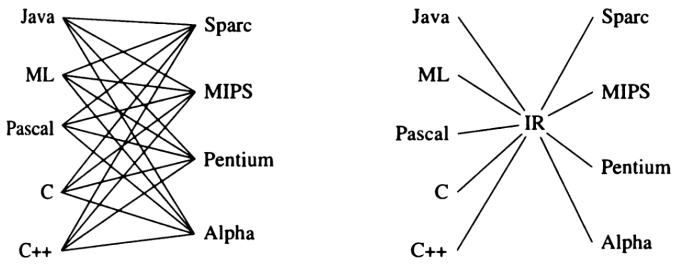


FIGURE 7.1. Compilers for five languages and four target machines: (left) without an IR, (right) with an IR.

For our compiler the intermediate representation tree language is defined by the signature tree, as shown in figure

```
signature TREE =
sig

datatype  exp  = CONST of int
               | NAME of Temp.label
               | TEMP of Temp.temp
               | BINOP of binop * exp * exp
               | MEM of exp
               | CALL of exp * exp list
               | ESEQ of stm * exp

     and  stm  = MOVE of exp * exp
               | EXP of exp
               | JUMP of exp * Temp.label list
               | CJUMP of relop * exp * exp * Temp.label * Temp.label
               | SEQ of stm * stm
               | LABEL of Temp.label

     and binop = PLUS | MINUS | MUL | DIV
               | AND | OR | LSHIFT | RSHIFT | ARSHIFT | XOR

     and relop = EQ | NE | LT | GT | LE | GE
               | ULT | ULE | UGT | UGE
end
```

IGURE 7.2. Intermediate representation trees.

Meaning:

expressions (exp), which stand for the computation of some value (possibly with side effects):

CONST(i) The integer constant i.

NAME(n) The symbolic constant n (corresponding to an assembly language label).

TEMP(t) Temporary t. A temporary in the abstract machine is similar to a register in a real machine. However, the abstract machine has an infinite number of temporaries.

BINOP(o, e_1, e_2) The application of binary operator o to operands e_1, e_2. Subexpression e_1 is evaluated before e_2. The integer arithmetic operators are PLUS, MINUS, MUL, DIV; the integer bitwise logical operators are AND, OR, XOR; the integer logical shift operators are LSHIFT, RSHIFT; the integer arithmetic right-shift is ARSHIFT. The Tiger language has no logical operators, but the intermediate language is meant to be independent of any source language; also, the logical operators might be used in implementing other features of Tiger.

MEM(e) The contents of *wordSize* bytes of memory starting at address e (where *wordSize* is defined in the `Frame` module). Note that when MEM is used as the left child of a MOVE, it means "store," but anywhere else it means "fetch."

CALL(f, l) A procedure call: the application of function f to argument list l. The subexpression f is evaluated before the arguments which are evaluated left to right.

ESEQ(s, e) The statement s is evaluated for side effects, then e is evaluated for a result.

The statements (`stm`) of the tree language perform side effects and control flow:

MOVE(TEMP t, e) Evaluate e and move it into temporary t.

MOVE(MEM(e_1), e_2) Evaluate e_1, yielding address a. Then evaluate e_2, and store the result into *wordSize* bytes of memory starting at a.

EXP(e) Evaluate e and discard the result.

JUMP($e, labs$) Transfer control (jump) to address e. The destination e may be a literal label, as in NAME(lab), or it may be an address calculated by any other kind of expression. For example, a C-language `switch(i)` statement may be implemented by doing arithmetic on i. The list of labels `labs` specifies all the possible locations that the expression e can evaluate to; this is necessary for dataflow analysis later. The common case of jumping to a known label l is written as JUMP(NAME l, [l]).

CJUMP(o, e_1, e_2, t, f) Evaluate e_1, e_2 in that order, yielding values a, b. Then compare a, b using the relational operator o. If the result is `true`, jump to t; otherwise jump to f. The relational operators are EQ and NE for integer equality and nonequality (signed or unsigned); signed integer inequalities LT, GT, LE, GE; and unsigned integer inequalities ULT, ULE, UGT, UGE.

SEQ(s_1, s_2) The statement s_1 followed by s_2.

LABEL(n) Define the constant value of name n to be the current machine code address. This is like a label definition in assembly language. The value NAME(n) may be the target of jumps, calls, etc.

What should the representation of an abstract syntax expression Absyn.exp be in the Tree language? At first it seems obvious that it should be Tree.exp. However, this is true only for certain kinds of expressions, the ones that compute a value. Expressions that return no value (such as some procedure calls, or while expressions in the Tiger language) are more naturally represented by Tree.stm. And expressions with Boolean values, such as $a > b$, might best be represented as a conditional jump - a combination of Tree.stm and a pair of destinations represented by Temp. labels. Therefore, we will make a datatype exp in the Translate module to model these three kinds of expressions:

```
datatype exp = Ex of Tree.exp
             | Nx of Tree.stm
             | Cx of Temp.label * Temp.label -> Tree.stm
```

Ex stands for an "expression," represented as a `Tree.exp`.

Nx stands for "no result," represented as a `Tree` statement.

Cx stands for "conditional," represented as a function from label-pair to statement. If you pass it a true-destination and a false-destination, it will make a statement that evaluates some conditionals and then jumps to one of the destinations (the statement will never "fall through").

For example, the Tiger expression `a>b|c<d` might translate to the conditional:

```
Cx(fn (t,f) => SEQ(CJUMP(GT,a,b,t,z),
                   SEQ(LABEL z, CJUMP(LT,c,d,t,f))))
```

for some new label z.

Sometimes we will have an expression of one kind and we will need to convert it to an equivalent expression of another kind. For example, the Tiger statement

flag := $(a > b | c < d)$

```
structure T = Tree

fun unEx (Ex e) = e
  | unEx (Cx genstm) =
      let val r = Temp.newtemp()
          val t = Temp.newlabel() and f = Temp.newlabel()
       in T.ESEQ(seq[T.MOVE(T.TEMP r, T.CONST 1),
                     genstm(t,f),
                     T.LABEL f,
                     T.MOVE(T.TEMP r, T.CONST 0),
                     T.LABEL t],
                 T.TEMP r)
      end
  | unEx (Nx s) = T.ESEQ(s,T.CONST 0)
```

PROGRAM 7.3. The conversion function `unEx`.

requires the conversion of a `Cx` into an `Ex` so that a 1 (for true) or 0 (for false) can be stored into `flag`.

It is helpful to have three conversion functions:

```
unEx :  exp → Tree.exp
unNx :  exp → Tree.stm
unCx :  exp → (Temp.label×Temp.label→Tree.stm)
```

Suppose e is the representation of `a>b|c<d`, so

```
e = Cx(fn(t,f) => ···)
```

Then the assignment statement can be implemented as

MOVE(TEMP$_{flag}$, unEx(e)).

The functions `unCx` and `unNx` are left as an exercise. It's helpful to have `unCx` treat the cases of CONST 0 and CONST 1 specially, since they have particularly simple and efficient translations. Also, `unCx(Nx _)` need not be translated, as it should never occur in compiling a well typed Tiger program.

7 Basic Blocks And Traces

It's useful to be able to evaluate the subexpressions of an expression in any order. If tree expressions did not contain ESEQ and CALL nodes, then the order of evaluation would not matter.

Some of the mismatches between `Trees` and machine-language programs are:

- The CJUMP instruction can jump to either of two labels, but real machines' conditional jump instructions fall through to the next instruction if the condition is false.
- ESEQ nodes within expressions are inconvenient, because they make different orders of evaluating subtrees yield different results.
- CALL nodes within expressions cause the same problem.
- CALL nodes within the argument-expressions of other CALL nodes will cause problems when trying to put arguments into a fixed set of formal-parameter registers.

We can take any tree and rewrite it into an equivalent tree without any of the cases listed above. Without these cases, the only possible parent of a SEQ node is another SEQ; all the SEQ nodes will be clustered at the top of the tree. This makes the SEQs entirely uninteresting; we might as well get rid of them and make a linear list of Tree.stms.

The transformation is done in three stages: First, a tree is rewritten into a list of canonical trees without SEQ or ESEQ nodes; then this list is grouped into a set of basic blocks, which contain no internal jumps or labels; then the basic blocks are ordered into a set of traces in which every CJUMP is immediately followed by its false label.

Thus the module `Canon` has these tree-rearrangement functions:

```
signature CANON =
sig
  val linearize : Tree.stm -> Tree.stm list

  val basicBlocks : Tree.stm list ->
                        (Tree.stm list list * Temp.label)

  val traceSchedule : Tree.stm list list * Temp.label ->
                        Tree.stm list
end

structure Canon : Canon
```

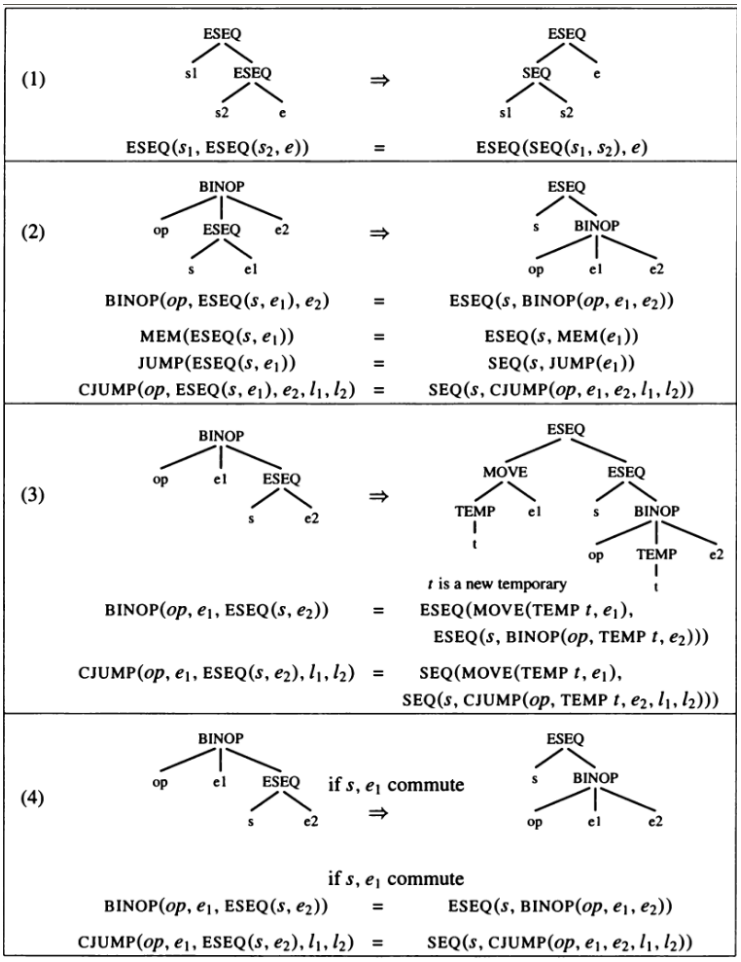
`Linearize` removes the ESEQs and moves the CALLs to top level. Then `BasicBlocks` groups statements into sequences of straight-line code. Finally `traceSchedule` orders the blocks so that every CJUMP is followed by its `false` label.

How can the ESEQ nodes be eliminated? The idea is to lift them higher and higher in the tree, until they can become SEQ nodes.

CANONICAL TREES

Let us define *canonical trees* as having these properties:

1. No SEQ or ESEQ.
2. The parent of each CALL is either EXP(...) or MOVE(TEMP t, ...).



It may happen that s causes no side effects that can alter the result produced by e_1. This will happen if the temporaries and memory locations assigned by s are not referenced by e_1 (and s and e_1 don't both perform external I/O). In this case, identity (4) can be used.

We cannot always tell if two expressions commute. For example, whether MOVE(MEM(x), y) commutes with MEM(z) depends on whether $x = z$, which we cannot always determine at compile time. So we *conservatively approximate* whether statements commute, saying either "they definitely do commute" or "perhaps they don't commute." For example, we know that any statement "definitely commutes" with the expression CONST(n), so we can use identity (4) to justify special cases like

BINOP(op, CONST(n), ESEQ(s, e)) = ESEQ(s, BINOP(op, CONST(n), e)).

The `commute` function estimates (very naively) whether two expressions commute:

```
fun commute(T.EXP(T.CONST _ ), _ ) = true
  | commute(_, T.NAME _ ) = true
  | commute(_, T.CONST _ ) = true
  | commute _ = false
```

The `reorder` function takes a list of expressions and returns a pair of (statement, expression-list). The statement contains all the things that must be executed before the expression-list. As shown in these examples, this includes all the statement-parts of the ESEQs, as well as any expressions to their left with which they did not commute. When there are no ESEQs at all we will use EXP(CONST 0), which does nothing, as the statement.

Algorithm. Step one is to make a "subexpression-extraction" method for each kind. Step two is to make a "subexpression-insertion" method: given an ESEQ-clean version of each subexpression, this builds a new version of the expression or statement.

ML makes this easy to express using little "`fn`" functions:

```
val reorder_stm: Tree.exp list * (Tree.exp list -> Tree.stm)
                                    -> Tree.stm

val reorder_exp: Tree.exp list * (Tree.exp list -> Tree.exp)
                                    -> Tree.stm * Tree.exp

fun do_stm(T.JUMP(e,labs)) =
          reorder_stm([e],fn [e] => T.JUMP(e,labs))
  | do_stm(T.CJUMP(p,a,b,t,f)) =
          reorder_stm([a,b], fn[a,b]=> T.CJUMP(p,a,b,t,f))
    ⋮ and so on

and do_exp(T.BINOP(p,a,b)) =
          reorder_exp([a,b], fn[a,b]=>T.BINOP(p,a,b))
  | do_exp(T.MEM(a)) =
          reorder_exp([a], fn[a]=>T.MEM(a))
    ⋮ and so on
```

`Reorder_stm` takes two arguments – a list l of subexpressions and a *build* function. It pulls all the ESEQs out of the l, yielding a statement s_1 that contains all the statements from the ESEQs and a list l' of cleaned-up expressions. Then it makes SEQ(s_1, *build*(l')).

`Reorder_exp`(l, *build*) is similar, except that it returns a pair (s, e) where s is a statement containing all the side effects pulled out of l, and e is *build*(l').

The left-hand operand of the MOVE statement is not considered a subexpression, because it is the *destination* of the statement – its value is not used by the statement. However, if the destination is a memory location, then the *address* acts like a source. Thus we have,

```
  | do_stm(T.MOVE(T.TEMP t,b)) =
          reorder_stm([b],fn[b]=>T.MOVE(T.TEMP t,b))
  | do_stm(T.MOVE(T.MEM e,b)) =
          reorder_stm([e,b],fn[e,b]=>T.MOVE(T.MEM e,b))
```

Now, given a list of extracted subexpressions, we pull the ESEQs out, from right to left.

MOVING CALLS TO TOP LEVEL

The `Tree` language permits CALL nodes to be used as subexpressions. However, the actual implementation of CALL will be that each function returns its result in the same dedicated return-value register TEMP(RV). Thus, if we have

BINOP(PLUS, CALL(...), CALL(...))

the second call will overwrite the RV register before the PLUS can be executed.

We can solve this problem with a rewriting rule. The idea is to assign each return value immediately into a fresh temporary register, that is

CALL(*fun*, *args*) → ESEQ(MOVE(TEMP t, CALL(*fun*, *args*)), TEMP t)

The rewriting rule is implemented as follows: `reorder` replaces any occurrence of CALL(f, *args*) by

ESEQ(MOVE(TEMP t_{new}, CALL(f, *args*)), TEMP t_{new})

and calls itself again on the ESEQ. But `do_stm` recognizes the pattern

MOVE(TEMP t_{new}, CALL(f, *args*)),

and does not call `reorder` on the CALL node in that case, but treats the f and *args* as the children of the MOVE node. Thus, `reorder` never "sees"

A LINEAR LIST OF STATEMENTS

Once an entire function body s_0 is processed with `do_stm`, the result is a tree s_0' where all the SEQ nodes are near the top (never underneath any other kind of node). The `linearize` function repeatedly applies the rule

SEQ(SEQ(a, b), c) = SEQ(a, *seq*(b, c))

The result is that s_0' is linearized into an expression of the form

SEQ(s_1, SEQ(s_2, ..., SEQ(s_{n-1}, s_n)...))

Here the SEQ nodes provide no structuring information at all, and we can just consider this to be a simple list of statements,

$s_1, s_2, \ldots, s_{n-1}, s_n$

where none of the s_i contain SEQ or ESEQ nodes.

These rewrite rules are implemented by `linearize`, with an auxiliary function `linear`:

```
fun linearize(stm0: T.stm) : T.stm list =
let
    ⋮ definitions of reorder_exp, reorder_stm, do_exp, do_stm, etc.

   fun linear(T.SEQ(a,b),l) = linear(a,linear(b,l))
     | linear(s,l) = s::l
 in
   linear(do_stm stm0, nil)
end
```

A *basic block* is a sequence of statements that is always entered at the beginning and exited at the end, that is:

- The first statement is a LABEL.
- The last statement is a JUMP or CJUMP.
- There are no other LABELs, JUMPs, or CJUMPs.

The algorithm for dividing a long sequence of statements into basic blocks is quite simple. The sequence is scanned from beginning to end; whenever a LABEL is found, a new block is started (and the previous block is ended); whenever a JUMP or CJUMP is found, a block is ended (and the next block is started). If this leaves any block not ending with a JUMP or CJUMP, then a JUMP to the next block's label is appended to the block. If any block has been left without a LABEL at the beginning, a new label is invented and stuck there.

We will apply this algorithm to each function-body in turn. The procedure "epilogue" (which pops the stack and returns to the caller) will not be part of this body, but is intended to follow the last statement. When the flow of program execution reaches the end of the last block, the epilogue should follow. But it is inconvenient to have a "special" block that must come last and that has no JUMP at the end. Thus, we will invent a new label done - intended to mean the beginning of the epilogue - and put a JUMP(NAME done) at the end of the last block.

Now the basic blocks can be arranged in any order, and the result of executing the program will be the same - every block ends with a jump to the appropriate place. We can take advantage of this to choose an ordering of the blocks satisfying the condition that each CJUMP is followed by its false label.

At the same time, we can also arrange that many of the unconditional jumps are immediately followed by their target label. This will allow the deletion of these jumps, which will make the compiled program run a bit faster.

A *trace* is a sequence of statements that could be consecutively executed during the execution of the program. It can include conditional branches. A program has many different, overlapping traces. For our purposes in arranging CJUMPs and false-labels, we want to make a set of traces that exactly covers the program: each block must be in exactly one trace. To minimize the number of JUMPs from one trace to another, we would like to have as few traces as possible in our covering set.

A very simple algorithm will suffice to find a covering set of traces. The idea is to start with some block – the beginning of a trace – and follow a possible execution path – the rest of the trace. Suppose block b_1 ends with a JUMP to b_4, and b_4 has a JUMP to b_6. Then we can make the trace b_1, b_4, b_6.

But suppose b_6 ends with a conditional jump CJUMP($cond, b_7, b_3$). We cannot know at compile time whether b_7 or b_3 will be next. But we can assume that some execution will follow b_3, so let us imagine it is that execution that we are simulating. Thus, we append b_3 to our trace and continue with the rest of the trace after b_3. The block b_7 will be in some other trace.

```
Put all the blocks of the program into a list Q.
while Q is not empty
    Start a new (empty) trace, call it T.
    Remove the head element b from Q.
    while b is not marked
        Mark b; append b to the end of the current trace T.
        Examine the successors of b (the blocks to which b branches);
        if there is any unmarked successor c
          b ← c
    End the current trace T.
```

ALGORITHM 8.2. Generation of traces.

FINISHING UP

An efficient compiler will keep the statements grouped into basic blocks, because many kinds of analysis and optimization algorithms run faster on (relatively few) basic blocks than on (relatively many) individual statements. For the Tiger compiler, however, we seek simplicity in the implementation of later phases. So we will flatten the ordered list of traces back into one long list of statements.

At this point, most (but not all) CJUMPs will be followed by their true or false label. We perform some minor adjustments:

- Any CJUMP immediately followed by its false label we let alone (there will be many of these).
- For any CJUMP followed by its true label, we switch the true and false labels and negate the condition.
- For any CJUMP($cond, a, b, l_t, l_f$) followed by neither label, we invent a new false label l'_f and rewrite the single CJUMP statement as three statements, just to achieve the condition that the CJUMP is followed by its false label:

```
CJUMP(cond, a, b, l_t, l'_f)
LABEL l'_f
JUMP(NAME l_f)
```