Secrets of the Glasgow Haskell Compiler inliner

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Abstract

Higher-order languages, such as Haskell, encourage the programmer to build abstractions by composing functions. A good compiler must inline many of these calls to recover an efficiently executable program.

In principle, inlining is dead simple: just replace the call of a function by an instance of its body. But any compiler-writer will tell you that inlining is a black art, full of delicate compromises that work together to give good performance without unnecessary code bloat.

The purpose of this paper is, therefore, to articulate the key lessons we learned from a full-scale "production" inliner, the one used in the Glasgow Haskell compiler. We focus mainly on the algorithmic aspects, but we also provide some indicative measurements to substantiate the importance of various aspects of the inliner.

1 Introduction

One of the trickiest aspects of a compiler for a functional language is the handling of inlining. In a functional-language compiler, inlining subsumes several other optimisations that are traditionally treated separately, such as copy propagation and jump elimination. As a result, effective inlining is particularly crucial in getting good performance.

The Glasgow Haskell Compiler (GHC) is an optimising compiler for Haskell that has evolved over a period of about ten years. We have repeatedly been through a cycle of looking at the code it produces, identifying what could be improved, and going back to the compiler to make it produce better code. It is our experience that the inliner is a lead player in many of these improvements. No other single aspect of the compiler has received so much attention.

This paper reports on selected algorithmic aspects of GHC's inliner, focusing on aspects that were not obvious to us — that is to say, aspects that we got wrong to begin with. For the sake of concreteness we focus throughout on GHC, but we stress that the lessons we learned are applicable to any compiler for a functional language, and indeed perhaps to compilers for other languages too.

1.1 Overview of the compiler

GHC uses the "compilation by transformation" approach to compiling Haskell (PJS98). After parsing, resolving lexical scopes, and typechecking, the Haskell source is desugared into a small, pure, explicitly-typed intermediate language called the *GHC Core language*. Many Core-to-Core transformations are applied to this intermediate form, before it is fed to the back end for code generation. Figure 1 illustrates this structure.

Some of the Core-to-Core transformations are global, module-at-a-time passes, such as strictness analysis (PP93), or let-floating (PPS96); these are depicted "Opt_i" in Figure 1. Many other useful transformations, are purely local, and are collected together into a single pass, called the *simplifier*. The most important single transformation performed by the simplifier is inlining, the focus of this paper. We often refer to the "inliner" meaning "that part of the simplifier that deals with inlining". There is no separate pass that deals with inlining.

The decision whether or not to inline a function clearly depends a great deal on how often the function is called – in particular, a very important special case is when the function has exactly one call site. So before each pass of the simplifier, GHC runs an *occurrence analyser* that decorates each binding site with occurrence information.

The local transformations implemented by the simplifier often cascade. In particular, inlining a function can often reveal new opportunities for (say) case elimination or constant folding, and these can in turn make a call site more attractive for further inlining. Accordingly, the simplifier tries hard to perform as many transformations as possible in a single pass. Despite this effort, we have found it impossible to guarantee to complete all transformations in a single pass, so we iterate the

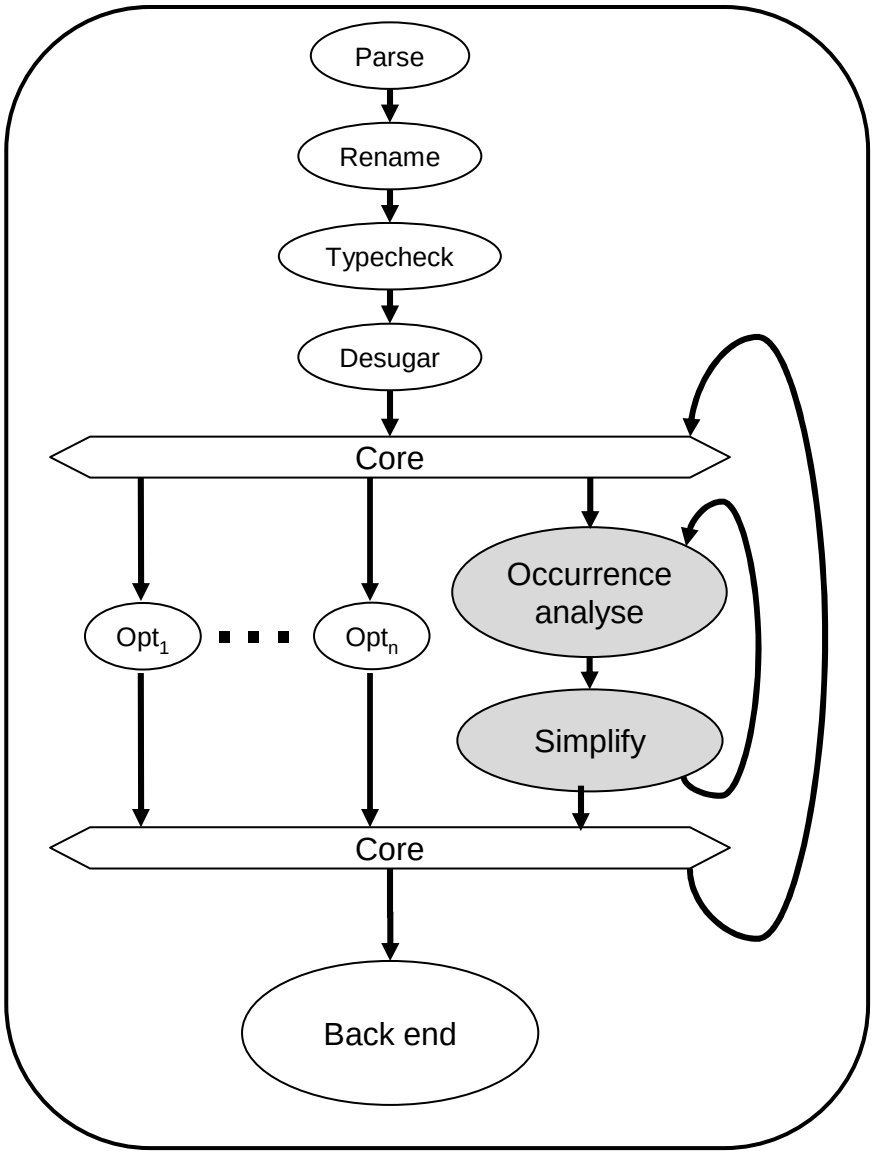


Fig. 1. Overall structure of GHC

occurrence-analysis/simplify pass, as suggested by the inner loop-back arrow in Figure 1. We iterate the loop until the simplifier indicates that no transformations occurred, or until some arbitrary number (currently 4) of iterations has occurred. This entire algorithm is applied between other major passes.

1.2 Contributions of the paper

Most papers about inlining focus on how to choose whether or not to inline a function called from many places. This is indeed an important question, but we have found that we had to deal with quite a few other less obvious, but equally interesting, issues.

Specifically, we describe the following, whose order of presentation roughly follows the flow of Figure 1:

Occurrence analysis. At first we were very conservative about inlining *recursive* definitions; that is, we did not inline them at all. But we found that this strategy sometimes behaves very badly. After a series of failed hacks we developed a simple, obviously-correct modification to the occurrence analyser, that does the job beautifully (Section 3).

Name capture. A major issue for any compiler, especially for one that inlines heavily, is *name capture*. Our initial brute-force solution involved inconvenient plumbing, but we have now evolved a simple and effective alternative, which we describe in Section 4.

Three-phase inlining. Because the compiler does so much inlining, it is important to get as much as possible done in each pass over the program. Yet one must steer a careful path between doing too *little* work in each pass, requiring extra passes, and doing too *much* work, leading to an exponential-cost algorithm. GHC now identifies *three* distinct moments at which an inlining decision may be taken for a particular definition, as we discuss in Section 5.

Evaluation state. When inlining an expression it is important to retain the expression's *lexical* environment, which gives the bindings of its free variables. But at the inline site, the compiler might know more about the *evaluation state* of some of those free variables — most notably, a free variable might be known to be (say) an evaluated pair at the inline site, but not at its original definition. Some key transformations make use of this extra information, and lacking it will cause an extra pass over the code. We describe how to exploit our name-capture solution to support accurate tracking of both lexical and evaluation-state environments in Section 6.

Implementation sketch. To make the story more concrete, we sketch our implementation in some detail, give some indicative performance measurements (Sections 6 and 7).

None of the algorithms we describe is individually very surprising. Perhaps because of this, the literature on the subject is very sparse, and we are not aware of published descriptions of any of our algorithms. Our contribution is to abstract some of what we have learned, in the hope that we may help others avoid the mistakes that we made.

2 Preliminaries

We begin by setting the scene. First, we say exactly what we mean by "inlining" (Section 2.1) and introduce the factors that affect the inlining decision (Section 2.2). Then we describe the GHC Core language (Section 2.3).

2.1 What is inlining?

Given a definition `x = E`, one can *inline* `x` at a particular occurrence by replacing the occurrence by `E`. (We use upper case letters, such as "`E`", to stand for arbitrary expressions, and "`==>`" to indicate a program transformation.) For example:

```
  let { f = \x -> x*3 } in f (a + b) - c
==>
  (a+b)*3 - c
```

We have found it useful to identify three distinct transformations that collectively implement what we informally describe as "inlining":

- *Inlining itself* replaces an occurrence of a `let`-bound variable by (a copy of) the right-hand side of its definition. Inlining `f` in the example above goes like this:

  ```
    let { f = \x -> x*3 } in f (a + b) - c
  ==> [inline f]
    let { f = \x -> x*3 } in (\x -> x*3) (a + b) - c
  ```

  Notice that not all the occurrences of `f` need be inlined, and hence that the original definition of `f` must, in general, be retained.
- *Dead code elimination* discards bindings that are no longer used; this usually occurs when all occurrences of a variable have been inlined. Continuing our example gives:

  ```
    let { f = \x -> x*3 } in (\x -> x*3) (a + b) - c
  ==> [dead f]
    (\x -> x*3) (a + b) - c
  ```

- *β-reduction* simply rewrites a lambda application `(\x->E) A` to `let {x = A} in E`. Applying β-reduction to our running example gives:

  ```
    (\x -> x*3) (a + b) - c
  ==> [beta]
    (let { x = a+b } in x*3) - c
  ```

The first of these is the tricky one; the latter two are easy. In particular, beta reduction simply creates a `let` binding.

In a lazy, purely functional language, inlining and dead-code elimination are both unconditionally valid, or meaning-preserving. (Neither is valid, in general, in a language permitting side effects, such as Standard ML or Scheme.) In particular, notice that inlining is valid, regardless of

- the number of occurrences of `x`,

- whether or not the binding for `x` is recursive,
- whether or not `E` has free variables (that is, inlining of nested definitions is perfectly fine), and
- the syntactic form of `E` (notably, whether or not it is a lambda abstraction).

Concerning the last of these items, notice that we (unconventionally) use the term "inline" equally for both functions and non-functions. Continuing the example, `x` can now be inlined, and then dropped as dead code, thus:

```
  (let { x = a+b } in x*3) - c
==> [inline x]
  (let { x = a+b } in (a+b)*3) - c
==> [dead x]
  (a+b)*3 - c
```

In this case, `x` is used exactly once, but we sometimes also inline non-functions that are used several times. Consider:

```
let x = (a,b)
in
...x...(case x of { (p,q) -> p+1 })...
```

By inlining `x` we can then eliminate the `case` to give

```
let x = (a,b)
in
...x...(a+1)...
```

In a similar way (when given bindings such as `x=y`), inlining subsumes copy propagation.

2.2 Factors affecting inlining

To say that inlining is *valid* does not mean that it is *desirable*. Inlining might increase code size, or duplicate work, so we need be careful about when to do it. There are three distinct factors to consider:

- *Does any code get duplicated, and if so, how much?* For example, consider

  ```
  let f = \v -> ...big... in (f 3, f 4)
  ```

  where "`...big...`" is a large expression. Then inlining `f` would not duplicate any work (`f` will still be called twice), but it will duplicate the code for `f`'s body. Bloated programs are bad (increased compilation time, decreased cache hit rates), but inlining can often *reduce* code size by exposing new opportunities for transformations. GHC uses a number of heuristics to determine whether an expression is small enough to duplicate.
- *Does any work get duplicated, and if so, how much?* For example, consider

  ```
  let x = foo 1000 in x+x
  ```

where `foo` is expensive to compute. Inlining `x` would result in two calls to `foo` instead of one.
Work can be duplicated even if `x` only appears once:

```
let x = foo 1000
    f = \y -> x * y
in ...(f 3)..(f 4)...
```

If we inline `x` at its (single) occurrence site, `foo` will be called every time `f` is. In general, we must be careful when inlining inside a lambda.
It is not hard to come up with examples where a single inlining that duplicates work gives rise to an arbitrarily large increase in run time. GHC is therefore very conservative about work duplication. In general, GHC never duplicates work unless it is sure that the duplication is a small, bounded amount.

- *Are any transformations exposed by inlining?* Often this is the case, as illustrated in Section 2.1. In general it is hard to predict whether any new transformations will be exposed and, like other compilers, we use a range of heuristics to answer this question (Section 7).

These considerations imply that inlining is not an optimisation "by itself". The *direct* effects of careful inlining are small: it may duplicate code or a constant amount of work, and usually saves a call or jump (albeit not invariably — see the example in the last bullet above). It is the *indirect* effects that we are really after: the main reason for inlining is that it often exposes new transformations by bringing together two code fragments that were previously separate. Thus, in general, inlining decisions must be influenced by context.

2.2.1 Trivial expressions

Sometimes we can be absolutely certain that inlining a variable will be beneficial: namely when the variable is bound to a *trivial* expression. A trivial expression is:

- A variable, or
- A literal, or
- A type application, where the function is trivial.

It is always good to inline a trivial expression: no code is duplicated, no work is duplicated, and new transformations may be enabled.

We mention the type-application case only for completeness. It is relevant only in a language that supports type abstraction and appplication, and then only if types are erased at run time, so that type application comes for free. This is the case for GHC.

2.2.2 Work duplication

If `x` is inlined in more than one place, or inlined inside a lambda, we have to worry about work duplication. When will such work duplication be bounded? Answer: at least in the cases when `x`'s right hand side is:

- A trivial expression.
- A constructor application.
- A lambda abstraction.
- An expression that is sure to diverge.

Constructor applications require careful treatment. Consider:

```
x = (f y, g y)
h = \z -> case x of
              (a,b) -> ...
```

It would plainly be a disaster, in general, to inline `x` inside the body of `h`, since that would potentially duplicate the calls to `f` and `g`. Yet we want to inline `x` so that it can cancel with the `case`. GHC therefore maintains the invariant that every `let`-bound constructor application has only arguments that can be duplicated with no cost: variables, literals, and type applications. We call such arguments *trivial* expressions, so the invariant is called the *trivial-constructor-argument invariant*. Once established, this invariant is easy to maintain (see Section 7.1).

The last case, that of divergent computations, is more surprising, but it is useful in practice. Consider:

```
sump = \xs ->
       let
         fail = error ("sump" ++ show xs)
       in let rec
         go = \xs ->
              case xs of
                []     -> 0
                (x:xs) -> if x<0 then fail
                          else x + go xs
       in
       go xs
```

Here `error` is the standard Haskell function that prints an error message and brings execution to a halt. Semantically, its value is just $\bot$, the divergent value. In this example, `sump` adds up the elements of a list, but reports an error if any element is negative. As it stands, a closure for `fail` will be allocated every time `sump` is called. It is perfectly OK to inline `fail`, because if `fail` is ever called, execution is going to halt anyway, so there is no work-duplication issue. If we do that, no closure is allocated; instead, `error` is called directly if an element turns out to be less than zero.

GHC has a predicate `whnfOrBot` that identifies expressions that are in WHNF or are certainly divergent:

```
whnfOrBot :: Expr -> Bool
```

One could easily imagine extending `whnfOrBot` to cover cases where a small amount of work other than allocation is duplicated, such as a few machine instructions.

2.3 The GHC Core Language

GHC is itself written in Haskell, so we define the Core language by giving its data type definition in Haskell:

```
type Program = [Bind]

      -- Bindings
data Bind = NonRec Var Expr
          | Rec [(Var, Expr)]

      -- Expressions
data Expr = Var   Var
          | Const Const
          | App   Expr Expr
          | Lam   Var Expr
          | Let   Bind Expr
          | Case  Expr Var [Alt]
          | Note  Note Expr

      -- Case alternatives
type Alt = (Const, [Var], Expr)

      -- Constants
data Const = Literal  Literal
           | DataCon  DataCon
           | PrimOp   PrimOp
           | DEFAULT

      -- Variables
data Var = MkVar String Unique
```

The Core language consists of the lambda calculus augmented with let-expressions (both non-recursive and recursive), case expressions, data constructors, literals, and primitive operations. In presenting examples we will use an informal, albeit hopefully clear, concrete syntax. We will feel free to use infix operators, and to write several bindings in a single non-recursive let-expression as shorthand for a sequence of let-expressions.

A program (`Program`) is simply a sequence of bindings, in dependency order. Each binding (`Bind`) can be recursive or non-recursive, and the right hand side of each binding is an expression (`Expr`).

The `Expr` data type has seven constructors, of which application (`App`), lambda abstraction (`Lam`), and let-expressions (`Let`) should be self-explanatory.

An expression may be a simple expression (constructor `Var`), containing a variable (of type `Var`)[1]. A variable (data type `Var`) consists of its name (a string) and a unique number; the name is used mostly for printing, and the unique number is for fast comparisons. The name by itself is not necessarily unique; in fact during compilation when it is necessary to change the unique number of a variable (see Section 4), the original name is retained, so that when dumping out intermediate code it is possible to relate variable names back to the original source program.

A constant expression is of the form `Const c`, where `c` is of type `Const`. The `Const` type includes literals, data constructors and primitive operators; the fourth case of the `Const` type, `DEFAULT` is illegal in an `Expr`.

The `Note` form of `Expr` allows annotations to be attached to the tree. The only impact on the inliner is discussed in Section 6.5.

Case expressions (`Case`) contain a list of alternatives, each of which is a triple of a constant, list of binders, and right-hand side. The number of binders must match the arity of the constant. The constant itself can be `DEFAULT`, `Literal` or `DataCon`, but it must not be `PrimOp`[2].

One other point about `Case` expressions is unusual: `Var` argument to `Case`. Consider the following Core expression,

```
case (reverse xs) of ys {
  (a:as) -> ys
  []     -> error "urk"
}
```

The unusual part of this construct is the binding occurrence of "`ys`", immediately after the "`of`" — the `Var` argument to the `Case` constructor records this unusual binder. The semantics is that `ys` is bound to the result of evaluating the scrutinee, `reverse xs` in this case, which makes it possible to refer to this value in the alternatives. We could instead insist that a `case` expression should only scrutinise a variable, using a `let`-binding to bind a non-variable scrutinee, but that would make it less structurally apparent that the scrutinee is sure to be evaluated (`let` is non-strict). In any case, this detail has no impact on the rest of this paper — indeed, we omit the extra binder in our examples — but we have found that it makes several transformations more simple and uniform, so we include it here for the sake of completeness.

GHC's actual intermediate language is very slightly more complicated than that given here. It is an explicitly-typed language based on System F_ω, and supports polymorphism through explicit type abstraction and application. It turns out that doing so adds only one new constructor to the `Expr` type, and adds nothing to the

[1] Note: Haskell allows a data constructor and a (perhaps-unrelated) type to have the same name.

[2] Philosophical aside. It would be possible to have a more refined type structure that did not have these side-conditions about what `Const` values can appear where, but we found that doing so made the code of the compiler signficantly longer, without really improving its robustness. There are many constraints on a Core program that are not statically checkable – for example, that every variable occurrence is in scope, and that the program is well-typed – so GHC optionally runs a type-checker after each optimisation phase. This type checker also checks the side conditions on `Const` values.

substance of this paper, so we do not mention it further. The main point is that this paper omits no aspect essential to a full-scale implementation of Haskell.

2.4 Separate compilation

GHC is capable of wholesale inlining across module boundaries. Whenever GHC compiles a module `M` it writes an "interface file", `M.hi`, that contains GHC-specific information about `M`, including the full Core-language definitions for any top-level definitions in `M` that are smaller than a fixed threshold. (This threshold is chosen so that few, if any, larger functions could possibly be inlined, regardless of the calling context.) When compiling any module, `A`, that imports `M`, GHC reads in `M.hi`, and is thereby equipped to inline calls in `A` to `M`'s exports. Since the definition of a function exported from `M` might refer to values *not* exported from `M`, GHC dumps into `M.hi` the transitive closure of all (sufficiently small) functions reachable from `M`'s exports. Values that are not exported from `M` may not be mentioned directly by the programmer, but may nevertheless be inlined by the inliner.

The consequence of all this is that `A` may need to be recompiled if `M` changes. There is no avoiding this, except by disabling cross-module inlining (via a command-line flag). GHC goes to some trouble to add version stamps to every inlining in `M.hi` so that it can deduce whether or not `A` *really* needs to be recompiled.

3 The occurrence analyser

It is clear that whether to inline `x` depends a great deal on how often `x` occurs in its scope. Before each run of the simplifier, GHC therefore runs an *occurrence analyser*, that performs two main functions:

- It decorates each binding site with an indication of how the bound variable occurs.
- It performs a dependency analysis of recursive binding groups, splitting them into their strongly-connected components.

This is all perfectly straightforward – we summarise the occurrence information we actually gather in Section 3.1. However, mutually recursive definitions present the inliner with a bit of a problem: the inliner may fail to terminate if it inlines them in an un-restrained way (Section 3.3). The main contribution of this section is to describe the simple new approach we have developed, which allows recursive definitions to be inlined without risking divergence. Our approach is by no means the only one possible — we review some in Section 3.6 — but it is simple, effective, and (so far as we know) not previously reported.

3.1 Simple occurrence analysis

The basic algorithm of the occurrence analyser is extremely simple. It performs a single bottom-up pass that annotates each binder with an indication of how it occurs, chosen from the following list:

`Dead`. The binder does not occur at all. For a `let` binder (whether recursive or not), the binding can be discarded, and the occurrence analyser does so immediately, so that it does not need to analyse the right hand side(s).

`Once`. The binder occurs exactly once, and that occurrence is not inside a lambda, nor is a constructor argument. Inlining is unconditionally safe; it duplicates neither code nor work. Section 2.2 explained why we must not inline an arbitrary expression inside a lambda, and also described the trivial-constructor-argument invariant.

`OnceInLam`. The binder occurs exactly once, but inside a lambda. Inlining will not duplicate code, but it might duplicate work (Section 2.2).

`ManyBranch`. The binder occurs at most once in each of several distinct case branches, and none of these occurrences is inside a lambda. For example:

```
case xs of
     []     -> y+1
     (x:xs) -> y+2
```

In this expression, `y` occurs only once in each case branch. Inlining `y` may duplicate code, but it will not duplicate work.

`Many`. The binder may occur many times, including inside lambdas. Variables exported from the module being compiled are also marked `Many`, since the compiler cannot predict how often they are used.

`LoopBreaker`. This binder breaks a mutually-recursive group, so do not inline it at all. The need for this item, and how we compute it, is the subject of much of the rest of this section.

Notice that we have three variants of "occurs once" (`Once`, `ManyBranch`, and `OnceInLam`). We have found all three to be important.

Some lambdas are certain to be called at most once. Consider:

```
let x = foo 1000
    f = \y -> x+y
in case a of
      []     -> f 3
      (b:bs) -> f 4
```

Here `f` cannot be called more than once, so no work will be duplicated by inlining `x`, even though its occurrence is inside a lambda. Hence, it would be better to give `x` an occurrence annotation of `Once`, rather than `OnceInLam`.

We call such lambdas *one-shot lambdas*, and mark them specially. They certainly occur in practice — for example, they are constructed as join points by the case-of-case transformation (PJS98). We are (still) working on a type-based analysis for identifying one-shot lambdas (WP99). Details of this analysis are beyond the scope of this paper, but our point here is that they are beautifully easy to exploit: the occurrence analyser simply ignores them when it is computing its "inside-lambda" information.

3.2 Ensuring termination

Inlining, together with beta reduction, corresponds closely to compile-time evaluation of the program, so we must clearly be concerned about ensuring that the compiler terminates. Non-termination may arise in two distinct ways:

Recursive bindings. If a recursively-bound variable is inlined at one of its occurrences, that will introduce a new occurrence of the same variable. Unless restricted in some way, inlining could go on for ever.

Recursive data types. Consider the following Haskell definition for `loop`:

```
data T = C (T -> Int)

g = \y -> case y of
            C h -> h y

loop = g (C g)
```

Here, `g` is small and non-recursive, so when processing `g (C g)`, `g` will be inlined. But the inlined call very soon rewrites to `g (C g)`, which is just the expression we started with.

The problem here is that the data type `T` is recursive, and *it appears contravariantly in its own definition* (How96).

Of these two forms of divergence, the former is an immediate and pressing problem, since almost any interesting Haskell program involves recursion.

In contrast, the latter situation is rather rare, and (embarrassingly) GHC can still be persuaded to diverge by such examples. The most straightforward solution is to spot such contravariant data types, and disable the case-elimination transformation

```
  case (C g) of { C h -> ...h... }
==>
  ...g...
```

The question of spotting contravariant data types is complicated by the fact that Haskell data types can be parameterised and mutually recursive. The MLj compiler (BKR98) restricts data types declarations somewhat, but does perform the analysis for exactly this reason.

Before discussing recursive bindings, it is worth noting two other possible sources of divergence that a Haskell compiler does *not* have to deal with. Firstly, in an untyped setting (such as a Scheme compiler) one can easily construct terms such as

```
  (\x -> x x) (\x -> x x)
```

This expression is not explicitly recursive, but it nevertheless reduces to itself. However, the strong-normalisation theorem for F_ω tells us that such terms simply must be ill-typed.

Secondly, side effects (which Haskell lacks) can create a recursive structure. For example[3]:

```
(let ((foo a-special-value)
      (bar a-special-value))
  (begin
     (set! foo (lambda ..bar..))
     (set! bar (lambda ..foo..))
     body))
```

Here, `foo` and `bar` are mutable locations, each of which is updated to refer to the other.

3.3 The problem

From now on we focus our attention on recursive bindings. We call a group of bindings wrapped in `rec` a *recursive group*. Unrestricted inlining of non-recursive bindings is safe, but unrestricted inlining of recursive bindings might lead to non-termination. One obvious thing to do, therefore, is to ensure that each recursive group really *is* recursive. To discover this, we regard each variable in the group as a *node*, and we record an *edge* from `f` to `g` if `f`'s right hand side mentions `g` (so `f` depends on `g`). The resulting collection of nodes and edges describes a graph, called the

[3] Thanks to Manuel Serrano for pointing this out.

dependency graph, whose strongly connected components are the smallest possible recursive groups (Pey87). To exploit this observation, GHC constructs the dependency graph for each `let rec`, and analyses its strongly-connected components. If there is more than one component, the `let rec` is split into a nest of recursive and non-recursive `let`s. GHC performs this analysis regularly; quite often, groups that were mutually-recursive fall into separate strongly-connected components as a result of earlier transformations.

So much is well known. But what do we do when we are faced with a genuinely recursive group? The simplest thing to do is not to inline any recursively-bound variables at all, and that is what earlier versions of GHC did. But this conservative strategy loses obviously-useful optimisation opportunities. Consider a recursive group of bindings:

```
let rec
  f = \ x -> ...g...
  g = \ y -> ...f...
in
...f...
```

By convention, other variables of interest, such as `g` in this case, are assumed not to be free in `...f...`. Since only `f` is called outside the `rec`, we can inline `g` at its unique call site to give:

```
let rec
  f = \ x -> ...(...f...)...
in
...f...
```

Here, the gain is modest. But sometimes inlining in `rec`s is critically important. Consider this:

```
let
  eq = ...
in
let rec
  d   = (eq,  neq)
  neq = \a b -> case d of
                  (e,n) -> not (e a b)
in
...
```

GHC generates code quite like this for an "`Eq` dictionary". A "dictionary" is a bundle of related "methods" for operating on values of a particular type. Here, the `Eq` dictionary, `d`, is a pair of methods (ordinary functions), `eq` and `neq`; the intention is that `eq` is a function that determines whether its arguments are equal, and `neq` determines whether they are unequal.

In this example, the `neq` method is specified by selecting the `eq` method from the dictionary `d`, calling it, and negating its result. You might think that it would be

more straightforward to call `eq` directly, but this code is generated by the compiler from `class` and `instance` declarations in the Haskell source code. We found that it was very hard, in general, to call the appropriate method directly; it was much easier to allow the front end to generate naive code, and let the simplifier take care of the rest.

In this particular example, `d` and `neq` are genuinely mutually recursive. Yet, if `d` were inlined in the body of `neq`, the `case` would cancel with the pair constructor, leading to the following:

```
let
  eq  = ...
  neq = \a b -> not (eq a b)
  d   = (eq,  neq)
in
...
```

Now everything is non-recursive, the definition of `neq` is improved, and inlining opportunities in the rest of the program are improved.

This is not an isolated or artificial example. Compiling Haskell's type-class-based overloading, using the dictionary-passing encoding sketched above, gives rise to pervasive recursion through these dictionaries. Failing to unravel the recursion has a devastating effect on performance, because overloaded functions include equality, ordering, and all numeric operations, some of which show up in almost any inner loop. We originally went to great lengths in the front end to avoid generating unnecessary dictionary recursion but, no matter how hard we tried, some unnecessary `rec`s *still* showed up. Our new approach uses a much simpler translation scheme, along with an inliner that does a good job of inlining `rec`-bound variables. This approach has the merit that it works equally well for complex recursions written by the programmer, though admittedly these are much less common.

3.4 The solution

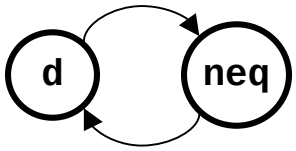
The real problem with recursive bindings is that they can make the inliner fall into an infinite loop. The key insight is this:

- *The inliner cannot loop if every cycle in the dependency graph is broken by a variable that is never inlined.*

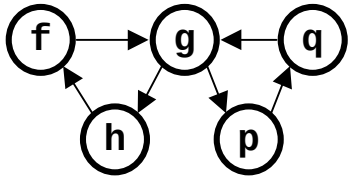
The conservative scheme works by never inlining *any* recursively-bound variable, but that is over-kill, as we saw in the example in Section 3.3:

```
rec
  d   = (eq,  neq)
  neq = \a b -> case d of
                  (e,n) -> not (e a b)
```

we obtained much better results by inlining `d` (but not `neq`) than by inlining neither. The dependency graph for this group forms a circle, thus:



To prevent the inliner diverging, it suffices to choose *either* of `d` or `neq`, and refrain from inlining it. In a more complicated situation, however, it might not be at all obvious which variable(s) suffice to break all the loops. For example, consider this more complex dependency graph:

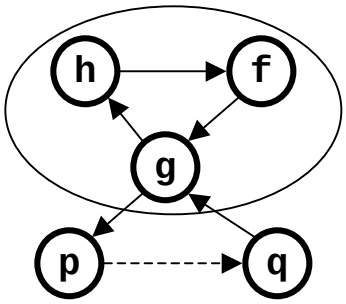


In this graph, we can break all the loops by picking `g` alone, or `f` and `q`, or `h` and `p`, or a variety of other pairs. To exploit this idea, we enhance the standard `rec`-breaking dependency analysis described above, in the following way. For each `rec` group, we construct its dependency graph, and then execute the following algorithm:

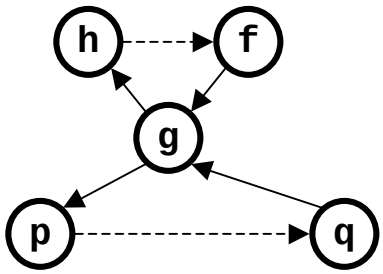
1. Perform a strongly-connected component analysis of the dependency graph.
2. For each strongly-connected component of the graph, perform the following steps, treating the components in topologically-sorted order; that is, deal first with the component that does not refer to any of the other components, and so on.
   (a) If the component is a singleton that does not depend on itself, do nothing.
   (b) Otherwise, choose a single variable, the *loop-breaker*, that will not be inlined. This choice is made using a heuristic we discuss shortly (Section 3.5).
   (c) Take the dependency graph of the component (a subset of the original graph), and delete all the edges in this graph that terminate at the loop-breaker.
   (d) Repeat the entire algorithm for this new dependency graph, starting with Step 1.

The algorithm is sure to terminate, because each iteration deletes at least one edge. Its worst-case efficiency is undoubtedly poor, and one could imagine various ways to improve it; for example, it may be possible to modify the existing strongly-connected component analysis after deleting some edges, rather than starting again from scratch. There is much to be said for simplicity, however, and so far we have not found a case where the algorithm behaves badly enough to be noticed.

Here is an example of the algorithm in action: consider the five-node dependency graph given above. It forms a single strongly-connected component. Suppose we pick `q` as a loop breaker; we delete arcs leading to it and perform the strongly-connected component analysis again. The reduced dependency graph has three strongly-connected components, namely {`p`}, {`f`, `g`, `h`}, and {`q`}



(We use dashed arcs for the arcs that are deleted in step (c).) Suppose now that we choose `f` as the loop breaker. Now we have no strongly connected components left in the reduced graph:



Notice that the only forward arcs are the dashed arcs leading to loop breakers. Reconstructing the recursive group in topologically sorted order (left to right in the diagrams) gives the following, where the "`*`" indicates the loop breakers:

```
rec
  p  = ...q...
  h  = ...f...
  g  = ...h...
  f* = ...g...
  q* = ...g...
```

The result of the algorithm is an ordered list of bindings with the following property: *the only forward references are to loop-breakers.* The bindings are still, of course, mutually recursive, but all the non-loop-breakers can be treated *exactly like non-recursive* `lets` so far as the inliner is concerned: their definition occurs before any of their uses, and inlining them cannot cause non-termination. The beauty of the loop-breaking algorithm is that the treatment of recursive `lets` is thereby factored into two independent pieces: first cut the loops, and then treat recursive and non-recursive bindings uniformly.

3.5 Selecting the loop breaker

There are two criteria that one might use to select a loop breaker:

- Try not to select a variable that it would be very beneficial to inline.

- Try to select a variable that will break many loops.

GHC currently uses only the first of these criteria. The second is a bit tricky to predict, and we have not explored using it. To evaluate the first criterion, GHC crudely "scores" each variable by how keen GHC is to inline it. Specifically, we pick the first of the following criterion that applies to the binding in question:

Score = 4, if the right hand side is trivial (Section 2.2.1). In this case the binding will certainly be inlined.

Score = 3, if the right hand side is a constructor application. Thus, we avoid selecting "`d`" in the example in Section 3.3, because its right hand side is a pair.

Score = 2, if the variable is marked with an INLINE pragma, indicating that the programmer was keen to inline it.

Score = 2, if the variable occurs just once (counting both the right hand sides of the `rec` itself and the body of the `let`). The variable is likely to be inlined if it occurs only once.

Score = 1, if the variable has rewrite rules or specialisations attached to it. Details of this are beyond the scope of this paper.

Score = 0, otherwise.

Then we pick a loop breaker by arbitrarily choosing one of the variables with lowest score. While this scoring mechanism is very crude, it seems adequate. In practice, we have never come across a `rec` in which a different choice of loop breaker would have made a significant difference. This amounts to anecdotal evidence only; we have not tried systematically to measure the effectiveness of loop-breaker choice.

3.6 Other approaches

A much more common approach to termination, taken by both Serrano (Ser97) and Wadell & Dybvig (WD97), is to bound both the *effort* that the inliner is prepared to invest, and the *size* of the expression it is prepared to build, when inlining a particular call. If either limit is exceeded, the inliner abandons the attempt to inline the call. Bounding effort deals with expressions, such as `(\x->x x)(\x->x x)`, that do not grow, but do not terminate either. The effort bound is typically set quite high, to allow for cascading transformations, so an effort bound alone might produce very large residual programs; that is why the size bound is necessary as well.

A variant of the approach retains a stack of inlinings that have been begun but not completed. When examining a call, the function is not inlined if an inlining of that same function is already in progress, or "pending". In effect, that function becomes the loop breaker, but it is chosen dynamically rather than statically.

This approach has the very great merit that it deals readily with all forms of non-termination: recursive functions, recursive data types, untyped languages and side effects, for example, all cause no problems. Even pathological programs that are not actually recursive, but which grow exponentially if one unconditionally inlines all non-recursive functions, can be dealt with.

The difficulty with this approach in our setting is that the simplifier is applied repeatedly, a dozen times or more, between applying other transformations (strictness

analysis, let-floating, etc). If each iteration accepts a given amount of code growth, or effort applied, then *each iteration might unroll a recursive function further.* The effort/size bound mechanism uses an auxiliary parameter (the effort/size budget) that is not recorded in the tree between successive iterations of the simplifier; it records the state of the inliner itself. Appel solves this by adding a fudge factor that makes successive applications of the inliner less and less keen to inline (parameter "E" on p92 of (App92)) — but that means that whole-module transformations applied late in the day are less likely to have their results exploited by the inliner.

Our approach does not have this problem: applications of the simplifier will eventually terminate. However, our more static analysis required that recursive functions and recursive data types be handled differently, which is undesirable. And yet more would be needed in an untyped or impure setting.

A quite separate, complementary, approach to inlining recursive functions is variously described as "loop headers" (App94), "labels-inline" (Ser97), "lambda-dropping" (DS97), and "the static argument transformation" (San95). The common idea is to turn a recursive function definition into a non-recursive function containing a local, recursive definition. Thus we can, for example, transform the standard recursive definition of `map`:

```
map = \f xs -> case xs of
                 [] -> []
                 (x:xs) -> f x : map f xs
```

into the following non-recursive definition:

```
map = \f xs ->
      let mp = \xs -> case xs of
                        []     -> []
                        (x:xs) -> f x : mp xs
      in mp xs
```

With the original definition, inlining would simply unroll a finite number of iterations of `map`. With the new definition, inlining `map` creates a new, specialised function definition for `mp` into which the particular `f` used at the call site can be inlined, perhaps resulting in better code — claimed benefits range from 1% to 10%. The overall effect is much better than that achieved by simply unrolling the original definition of `map`; unrolling a loop reduces the overheads of the loop itself, whereas creating a specialised function, `mp`, reduces the cost the computation in each iteration of the loop.

The static argument transformation may indeed be useful, but it is orthogonal to the main thrust of this paper. It is best considered as a separate transformation, performed on `map` before inlining is begun, that enhances the effectiveness of inlining.

Another orthogonal question is that of loop unrolling. A loop breaker could be inlined a fixed number of times to gain the effect of loop unrolling.

Identifying loop breakers can be useful for other purposes besides guiding the inliner. Compilers that support pre-emptive concurrency sometimes require that a

Allocations	No libs	Libs too
Mean	+23%	+78%
Min	−15%	0%
Max	+200%	+1125%

Fig. 2. Effect on total allocation of switching off the loop-breaker algorithm

garbage-collection safe point breaks every loop, and loop-breaker information can clearly be used to identify where such safe points must be inserted. The SML NJ compiler does exactly this, using an (unpublished) branch and bound algorithm, although it does not use the information to guide inlining.

3.7 Results

It is hard to offer convincing measurements for the effectiveness of the loop-breaker algorithm, because GHC is now built in the expectation that `rec`s that can be broken will be. Nevertheless, Figure 2 gives some indicative results.

For a deterministic measure of runtime, we use the amount of memory allocation performed by the compiled Haskell program. For Haskell programs, memory allocation tends to vary reasonably linearly with execution time.

To measure the effects of switching the loop-breaker algorithm off, we arranged to mark *every* `rec`-bound variable as a loop breaker. In the results table, the "Mean" row shows the *geometric* mean of the ratio between the switched-off version and the baseline version — we use a geometric mean because we are averaging ratios (FW86). The "Min" and "Max" rows show the most extreme ratios we found.

The effects are dramatic. The column headed "No libs" has the loop-breaking algorithm switched off when compiling the application, but not when compiling the standard libraries. The column "Libs too" shows the effect of switching off the loop-breaking algorithm when compiling the standard libraries as well. The importance of the libraries is that they contain implementations of arithmetic over basic types; if that is compiled badly then performance suffers horribly. (We are investigating the strange −15% figure, which suggests that switching off loop breakers improved at least one program.)

3.8 Summary

In retrospect, the algorithm is entirely obvious, yet we spent ages trying half-baked hacks, none of which quite worked, before finally biting the bullet and finding it quite tasty. It is more likely to be important for compilers for lazy languages than for strict ones, because only non-strict languages allow recursive data structures, and it is there that the most important performance implications show up. However, as our first example demonstrated, even where no data structures are involved, useful improvements can be had.

4 Name capture

We now turn our attention to the inliner proper, beginning with the tiresome but pervasive problem of name capture. It is well known that any transformation-based compiler must be concerned about *name capture* (Bar85). Consider, for example:

```
let x = a+b in
   let a = 7 in
      x+a
```

It is obviously quite wrong to inline `x` to give:

```
let a = 7 in
   (a+b) + a
```

because the `a` that was free in `x`'s right hand side has been captured by the `let` binding for `a`.

4.1 The sledge hammer

Earlier versions of GHC used a brutal approach to avoid the name-capture problem: during inlining, GHC would simply rename, or *clone*, every single bound variable, to give:

```
let s796 = 7
in (a+b) + s796
```

This renaming made use of a supply of fresh names that, in this example, has arbitrarily renamed `a` to `s796`. This approach suffers from two disadvantages:

- It allocates far more fresh names than are actually necessary, and there is sure to be a compile-time performance cost to this.
- Plumbing the supply of fresh names to the places those names are required is sometimes very painful.

Why is there a compile-time performance cost to the sledge-hammer approach? Firstly, because a variable is a structure containing a name; to rename the variable we must copy the structure, inserting the new name. Secondly, the substitution mapping old names to new names becomes larger. Lastly, if the substitution is empty we can sometimes avoid looking at an expression or type at all — but if all names are cloned the substitution is never empty.

If the compiler were written in an impure language, fresh names could be allocated by side effect, but GHC is written in Haskell, which does not have side effects. Using the trees of Augustsson et al. is the best solution we know of (ARS94), but it still involves plumbing a tree of fresh names everywhere they *might* be needed. Worse, the fresh names usually *aren't* needed, but the tree is nevertheless built. This unnecessary work is deeply irritating. Finally, even if we were not worried about performance, it is sometimes extremely painful to get the name supply to where it is needed. For example, in a typed intermediate language it should be possible to have a function:

```
exprType :: Expr -> Type
```

that figures out the type of an expression. But suppose the expression is something like:

```
filter Int pred xs
```

The function `filter` has the polymorphic type

```
filter :: forall a. (a -> Bool) -> [a] -> [a]
```

So to figure out the type of the subexpression (`filter Int`) we must instantiate `filter`'s type, substituting `Int` for `a`. Oh no! Substitution! That can, in general, give rise to name capture. So we need to feed a name supply to `exprType`:

```
exprType :: NameSupply -> Expr -> Type
```

This "solution" is deeply unattractive, and the situation is only different in its cosmetics if the name supply is hidden in a monad. Something better is required.

4.2 The rapier

Suppose we write the call $subst\ M\ [E/x]$ to mean the result of substituting E for x in M. The standard rule for substitution (Bar85) when M is a lambda abstraction is:

$$\begin{array}{lcl} subst\ (\lambda x.M)\ [E/x] & = & \lambda x.M \\ subst\ (\lambda x.M)\ [E/y] & = & \lambda x.(subst\ M\ [E/y]) \\ & & \text{if } x \text{ does not occur free in E} \end{array}$$

If the side condition does not hold, one must rename the bound variable x to something else. The brute-force solution does this renaming regardless.

Suppose that we lacked a name supply, but instead knew the free variables of E. Then we could test the side condition easily and, in the common case where there is no name capture, find that there was no need to rename x. But what if x *was* free in E? Then we need to come up with a fresh name for x that is not free in E. A simple approach is to try a variant of x, say "x_1". If that, too, is free in E, try "x_2", and so on.

When we finally discover a name, x_n, that is not free in E, we can augment the substitution to map x to x_n and apply this substitution to M, the body of the lambda. In general, then, we must simultaneously substitute for several variables at once.

To make this work at all, though, we need to know the free variables of E, or, more generally, the free variables of the range of the substitution. One way to find this is simply to compute the free variables directly from E, but if E is large this might be costly. However, it suffices to know any *superset* of these free variables. One obvious choice is *the set of all variables that are in scope*. If we made this choice, then we would end up renaming any bound variable for which there was an enclosing binding. We call this the *no-shadowing strategy*, for obvious reasons. The no-shadowing strategy will rename some variables when it is not strictly necessary

to do so, but it has the desirable property of idempotence: a complete pass of the simplifier that happens to make no transformations will clone no variables. This is a good thing. Usually, some parts of the program being compiled are fully-transformed before others; the no-shadowing strategy reduces gratuitous "churning" of variable names.

Thus, we are led to a substitution algorithm that has three parameters, instead of two: the expression to which the substitution is applied, the substitution itself, ϕ, and the set of in-scope variables, θ:

$$\begin{array}{lcl} subst\ (\lambda x.M)\ \phi\ \theta & = & \lambda x.subst\ M\ (\phi \setminus x)\ (\theta \cup \{x\}) \\ & & \text{if } x \notin \theta \\ subst\ (\lambda x.M)\ \phi\ \theta & = & \lambda y.subst\ M\ (\phi[x \mapsto y])\ (\theta \cup \{y\}) \\ & & \text{where } y \notin \theta \end{array}$$

Notice how conveniently the set of in-scope variables can be maintained. Almost all the time, it simply travels everywhere with the substitution; we shall see some interesting exceptions to this general rule in Section 7.

There is one other important subtlety in this algorithm: in the case where x is not in θ we must *delete* x from the substitution, denoted $\phi \setminus x$. How could x be in the domain of the substitution, but not be in scope? Here is an example:

```
let x = a*b in (x, \x -> x+3)
```

The outer `x` occurs exactly once, so the simplifier discards the outer `let` binding and simply adds [`x` $\mapsto$ `a*b`] to the substitution. This mapping must be deleted from the substitution inside the `\x` abstraction, else we will erroneously get `\x -> a*b+3`. Situations like this certainly occur in practice — we have the scars to show for it.

Occasionally, the set of in-scope variables is not conveniently to hand when starting a substitution. In that case, it is easy to find the set of free variables of the range of the substitution, and use that to get the process started.

4.3 Choosing a new name

The other choice that must be made in the algorithm is to choose a fresh name, in the (hopefully rare) cases where that proves necessary. We could just try x_1, x_2, and so on, but there is a danger that once $x_1 \ldots x_20$ are in scope, then any new x will make 20 tries before finding x_{21}. A simple way out is to compute some kind of hash value from the set of in-scope variables, and use that, together perhaps with the variable to be renamed, to choose a new name. Indeed, simply using the number of enclosing binders as the new variable name gives something not unlike de Bruijn numbers (see Section 4.5). The nice thing is that any old choice will do; the only issue is how many iterations it takes to find an unused variable.

4.4 Measurements

We made some simple measurements of the effectiveness of our approach. We compiled the entire `nofib` suite, some 370 Haskell modules, comprising around 50,000

	Number of attempts				
	0	1	2	3 − 9	10+
Mean	93.2%	1.3%	0.7%	1.6%	3.2%
Min	0.94%	0%	0%	0%	0%
Max	100%	10%	6.13%	18.2%	94%

Fig. 3. Cloning rates

lines of code in total (Par92). The size of each module varied from a few dozen lines to a thousand lines or so.

Figure 3 summarises how many "tries" it took to find a variable name that was not in scope. The columns show what proportion of binders required zero, one, two, 3-9, and 10 or more attempts, to find a variable name that was not already in scope. We measured these proportions separately for each module, and then took the arithmetic mean of the resulting figures. The "min" (resp "max") rows show the smallest (resp largest) proportions encountered among the entire set of modules.

The zero column corresponds to the situation where the binder is not shadowed; as expected, this is the case for the vast majority (93%) of binders. Our hash function (we simply picked an arbitrary member of the in-scope set as a hash value) is obviously too simple, though: on average 3.2% of all binders required more than ten attempts to find a fresh name, and in one pathological module almost all binders required more than ten attempts. This pathological case suggests that there is plenty of room for improvement in the hash function.

4.5 Other approaches

Another well-known approach to the name-capture problem to use de Bruijn numbers (dB80). Apart from being entirely unreadable, this approach suffers from the disadvantage that *when pushing a substitution inside a lambda, the entire range of the substitution must have its de Bruijn numbers adjusted.* That operation can be carried out lazily, to avoid a complexity explosion when pushing a substitution inside multiple lambdas, but that means yet more administration.

It is far from clear that using de Bruijn numbers gains any efficiency, and they carry a considerable cost in terms of the opacity of the resulting program. (Programmers will not care about this, but compiler writers do.)

There is one fairly compelling reason for using de Bruijn numbers. Precisely because they do discard the original variable names, many more common subexpressions can arise. These CSEs increase sharing of the compiler's representation of the program; they do not in general represent run-time sharing. However this compile-time sharing can be particularly important when dealing with types, which can get large. Shao, for example, reports substantial savings when using de Bruijn numbers (for types) together with hash-consing (SLM98). However, our types are

smaller than his (we are not compiling SML modules) so type sizes only become an issue for deliberately pathological programs whose types are exponential in the size of the program (LP91).

Another popular approach to the name-capture problem is this: establish the invariant that every bound variable is unique in the whole program. Appel *et al* only inline functions called exactly once, and then the situation is even easier: inlining preserves the unique-variable invariant without any cloning at all (AJ97). GHC inlines functions called more than once, but it could still maintain the invariant by *cloning all the locally-bound variables of an inlined expression*. There are three difficulties here. First, we found in practice that (in GHC at least) there were many transformations (other than inlining) that had to do extra work to maintain the global-uniqueness invariant. Secondly, this strategy will do more cloning than is really necessary. Thirdly, cloning the local binders of an inlined expression implies a whole extra pass over that expression, prior to simplifying the expression in its new context. Our approach, of maintaining an in-scope set, combines the cloning pass with the simplification pass, and simultaneously reduces the amount of cloning that has to be done.

4.6 Summary

Our new substitution algorithm is a simple re-working of the standard algorithm in Barendregt (Bar85). What is interesting is that the resulting algorithm seems quite practical. Even if the compiler were written in a language where name-supply plumbing was not an issue, maintaining the set of in-scope variables makes it easy to reduce the amount of cloning that is done.

5 The three-phase inlining strategy

The simplifier tries hard to perform as many transformations in a single pass as possible. Driven by this goal, the simplifier now makes an inlining decision about a particular `let` bound variable at no fewer than three distinct moments. Consider again the expression:

```
let x = E in B
```

Here are the three occasions on which we may consider inlining `x`:

PreInlineUnconditionally. When the simplifier meets the expression for the first time, it considers whether to inline `x` unconditionally in `B`. It does so if and only if `x` is marked `Once` (see Section 3). In this case, the simplifier does not touch `E` at all; it simply binds `x` to `E` in its current substitution, discards the binding completely, and simplifies `B` using this extended substitution. This is the main use of the substitution beyond dealing with name capture, but it needs a little care, as we discuss in Section 6.2.

Notice, crucially, that *the right hand side of the definition is processed only once*, namely at the occurrence site. It turns out that this is very important. If the right hand side *is* processed when the `let` is encountered, and then again at the occurrence of the variable, the complexity of the simplifier becomes exponential in program size. Why? Because the right hand side is processed twice; and it might have a `let` whose right hand side is then processed twice *each time*; and so on. In retrospect this is obvious, but it was very puzzling at the time!

PostInlineUnconditionally. If the pre-inline test fails, the simplifier next simplifies the right hand side, `E`, to produce `E'`. It then again considers whether to inline `x` unconditionally in `B`. It decides to do so if and only if

- `x` is not exported from this module (exported definitions must not be discarded), and
- `x` is not a loop breaker (Section 3.5), and
- `E'` is *trivial* (Section 2.2.1).

If so, then again the binding is dropped, and `x` is mapped to `E'` in the substitution. This case is quite common; it corresponds to copy propagation in a conventional compiler. It often arises as a result of β-reduction. For example, consider the definitions:

```
f = \x -> E
t = f a
```

If `f` is inlined, we get a β redex, and thence

```
f = \x -> E
t = let x = a in E
```

The interesting question is why we do not make this test at the PreInlineUnconditionally stage, something we discuss in Section 5.1.

CallSiteInline. If neither of the above holds, GHC retains the `let` binding, adds `x` to the in-scope set. While processing `B`, at every *occurrence* of `x`, GHC considers whether to inline `x`. This decision is based on a fairly complex heuristic, that we discuss in Section 7. If the decision is "Yes", then GHC needs to have access to `x`'s definition; this can be achieved quite elegantly, as we discuss in Section 6.3.

5.1 Why three-phase?

One obvious question is this: why not combine PostInlineUnconditionally with PreInlineUnconditionally? That is, before processing `E`, why not look to see if it is trivial (e.g. a variable), and if so inline it unconditionally? Doing so is a huge, but rather subtle, mistake.

The mistake is to do with the correctness of the pre-computed occurrence information. Suppose we have:

```
let
  a = ...big...
  b = a
in
...b...b...b...
```

`a` will be marked `Once`, and hence will be inlined unconditionally. But if PreInlineUnconditionally now sees that `b`'s right-hand side is just `a`, and inlines `b` everywhere, `a` now effectively occurs in many places. This is a disaster, because `a` is now inlined unconditionally in many places.

The cause of this disaster is that `a`'s occurrence information was rendered invalid by our decision to inline `b`. Several solutions suggest themselves — for example, provide some mechanism for fixing `a`'s occurrence information; or get the occurrence analyser to propagate `b`'s occurrences to `a` — and we tried some of them. They were all complicated, and the result was a bug farm. Appel and Jim describe one way to update occurrence information on the fly, depending critically on their assumption that functions are only inlined if they are called exactly once, but even that was very tricky (AJ97).

We finally discovered the three-phase inline mechanism we have described. It is simple, and obviously correct. The PreInlineUnconditionally phase only inlines a variable `x` if `x` occurs once, not inside a lambda. That means that *the occurrence information for any variable,* `y`*, free in* `x`*'s right hand side is unaffected by the inlining.*

On the other hand, once the right hand side has been processed, if `y` is going to be inlined unconditionally, then that will have happened already. In our example, PreInlineUnconditionally will decide to inline `a`. Now the simplifier moves on to the binding for `b`. PreInlineUnconditionally declines to inline, so the right hand side of `b` is processed; `a` is inlined, and (a processed version of) `...big...` is produced. This is not trivial, so PostInlineUnconditionally declines too.

Another obvious question is whether PostInlineUnconditionally could be omitted altogether, leaving CallSiteInline to do its work. Here the answer is clearly "yes";

	Pre	Post	CallSite
Mean	47.4%	17.4%	35.2%
Min	0.25%	0.92%	0.72%
Max	80%	95%	98%

Fig. 4. Relative frequency of inlining

PostInlineUnconditionally is just an optimisation that allows trivial bindings to be dropped a little earlier than would otherwise be the case.

To summarise, the key feature of our three-phase inlining strategy is that it allows the use of simple, pre-computed occurrence information, while still avoiding the exponential blowup that can occur if PreInlineUnconditionally is omitted.

5.2 Measurements

Figure 4 gives some simple measurements of the relative frequency of each form of inlining. We used the same set of benchmark programs in in Section 4.4, gathered statistics on how often each sort of inlining was used, and averaged these separately-calculated proportions. We took *arithmetic* means of the percentages, because here we are averaging "slices of the pie", so the "Mean" line should still sum to 100%.

The figures indicate that on average, each sort of inlining is actually used in practice, and that each dominates in some programs.

6 The inliner's data types

We have now described the key design decisions in our inliner. In this section we make the description more concrete by sketching the implementation itself. In this section we say more about the types involved: the type of the simplifier itself (Section 6.1), the substitution (Section 6.2), the in-scope set (Section 6.3) and the context (Section 6.4). In the next Section we return to the implementation of the inlining algorithm itself.

6.1 The type of the simplifier

The simplifier takes a substitution, a set of in-scope variables, an expression, and a "context", and delivers a simplified expression:

```
simplExpr :: Subst -> InScopeSet
          -> InExpr -> Context
          -> OutExpr
```

The real simplifier's type is a bit more complicated than this: it takes an argument that enables or disables individual transformations; it gathers statistics about how many transformations are performed; and it takes a name supply, to use when it has to conjure up a fresh name not based on an existing name[4]. However, we will not need to consider these aspects here.

The substitution and in-scope set perform the roles described in Section 4, and we describe their implementation in more detail in Sections 6.2 and 6.3 respectively. they both have further uses. The *context* tells the simplifier something about the context in which the expression appears (e.g. it is applied to some arguments, or it is the scrutinee of a `case` expression). This context information is important when making inlining decisions (Section 6.4).

We refer to an un-processed expression as an "in-expression", and an expression that has already been processed as an "out-expression", and similarly for variables. The reasons for making these distinctions are described in Section 6.2.

```
type InVar  = Var
type InExpr = Expr
type InAlt  = Alt

type OutVar  = Var
type OutExpr = Expr
type OutAlt  = Alt
```

As indicated in Section 2.3, the simplifier treats an entire Haskell module (which GHC treats as a compilation unit) as a sequence of bindings, some recursive and some not. It deals each of these bindings in turn, just as if they were in a nested sequence of `lets`.

[4] We could certainly do without this name supply, by conjuring up names based on an arbitrary base name, but it turns out that it can conveniently piggy-back on the (monadic) plumbing for the other administrative arguments.

6.2 The substitution

As we have just seen, the simplifier (`simplExpr`) carries along (a) the current substitution, and (b) the set of variables in scope. But since the simplifier is busy transforming the expression and cloning variables, we have to be more precise:

- The domain of the substitution is *in-variables*.
- The in-scope set consists of *out-variables*.

But what is the *range* of the substitution? When used for cloning or PostInlineUnconditionally the range was an *out-expression*, but when used in PreInlineUnconditionally the range was an *in-expression*. We have to distinguish these two cases, because an in-expression makes no sense without the substitution in force at the original site of that in-expression. Thus we are led to the following definition for the substitution:

```
type Subst    = FiniteMap InVar SubstRng
data SubstRng = DoneEx OutExpr
              | SuspEx InExpr Subst
```

A `DoneEx` is straightforward, and is used both by the name-cloning mechanism, and by PostInlineUnconditionally. A `SuspEx` (`Susp` for "suspended") is used by PreInlineUnconditionally, and pairs an in-expression with the substitution appropriate to its `let` binding; you can think of it as a suspended application of `simplExpr`. Notice that we do *not* capture the in-scope set as well. Why not? Because we must use the in-scope set appropriate to the occurrence site — Section 7 amplifies this point.

6.3 The in-scope set

We mentioned at the beginning of Section 5 that the simplifier needs access to a `let`-bound variable's right-hand side at its occurrence site(s). A simple way to achieve this is to turn the in-scope set into a finite mapping:

```
type InScopeSet = FiniteMap OutVar Definition
data Definition = Unknown
                | BoundTo OutExpr OccInfo Level
                | NotAmong [Const]
data Level = TopLevel | Nested
```

Whether or not a variable is in scope can be answered by looking in the domain of the in-scope set (we still call it a "set" for old times sake). But the range of the mapping records what value the variable is bound to:

`Unknown` is used for variables bound in lambda and case patterns. We don't know what value such a variable is bound to.

`BoundTo` is used for `let` bound variables (both recursive and non-recursive), and records the right-hand side of the definition and the occurrence information left

with the binding by the occurrence analyser, and whether the binding is a top-level definition or a nested one. All this information is needed when making the inlining decision at occurrence sites.

`NotAmong` lists constants that the variable is *not* bound to (Section 6.3.1).

The in-scope set is also a convenient place to record information that is valid in only *part* of a variable's scope. Consider:

```
\x -> ...(case x of (a,b) -> E)...
```

When processing `E`, but not in the "..." parts, `x` is known to be bound to `(a,b)`. So, when processing the alternative of a `case` expression whose scrutinee is a variable, it is easy for the simplifier to modify the in-scope set to record `x`'s binding. Why is this useful? Because `E` might contain another `case` expression scrutinising `x`:

```
...(case x of (p,q) -> F)...
```

By inlining `(a,b)` for `x`, we can eliminate this `case` altogether. This turns out to be a big win (PJS98).

To summarise, the in-scope set, extended to be an in-scope mapping, plays the role of a *evaluation-state environment*. It records knowledge of the value of each in-scope variable, including knowledge that may be true for only part of that variable's scope. The nice thing is that this evaluation-state knowledge can elegantly be carried by the in-scope set, which we need anyway. The details of the transformations that exploit that evaluation-state knowledge are beyond the scope of this paper, but one simple one is this: if a `case` expression scrutinises a variable whose value is known, the `case` can be eliminated. For example:

```
case x of (a,b) ->
     ....(case x of (p,q) -> ...) ...
```

At the inner `case`, the value of `x` is known to be `(a,b)`, so the inner case can be eliminated.

6.3.1 The `NotAmong` *variant*

The `NotAmong` variant of the `Definition` type allows the simplifier to record negative information:

```
case x of
  Red     -> ...
  Blue    -> ...
  Green   -> ...
  DEFAULT -> E
```

The `DEFAULT` alternative matches any constructors other than `Red`, `Blue`, and `Green`. GHC supports such `DEFAULT` alternatives directly, rather than requiring `case` expressions to be exhaustive, which is dreadful for large data types. Inside `E`, what is known about `x`? What we know is that it is *not* bound to `Red`, `Blue`, or `Green`. This can be useful; if `E` contains a `case` expression that scrutinises `x`, we can eliminate

any alternatives that cannot possibly match. Similarly, the expression `x `seq` F` inside `E` can be transformed to just `F`, since `NotAmong` implies that `x` is evaluated[5]. Even the value `NotAmong []` is useful: it signals that the variable is evaluated, without specifying anything about its value.

6.4 The context

It should by now be clear that the *context* of an expression plays a key role in inlining decisions. For example, a function that is called should be inlined more vigorously than a function that is simply passed as an argument to another function.

For a long time we passed in a variety of *ad hoc* flags indicating various things about the context, but we have now evolved a much more satisfactory story. The context is a little like a continuation, in that it indicates how the result of the expression is consumed. But this continuation *must not be represented as a function* because we must be able to ask questions of it, as the earlier sub-sections indicate.

So GHC's contexts are defined by the following data type:

```
data Context
  = Stop
  | AppCxt InExpr Subst Context
  | CaseCxt  InVar [InAlt] Subst Context
  | ArgCxt   (OutExpr -> OutExpr)
  | InlineCxt  Context
```

The `Stop` context is used when beginning simplification of a lazy function argument, or the right hand side of a `let` binding. The `AppCxt` context indicates that the expression under consideration is to be applied to an argument. The argument is as yet un-simplified, and must be paired with its substitution. Similarly, the `CaseCxt` context is used when simplifying the scrutinee of a `case` expression.

`simplExpr` simply recurses into the expression, building a context "stack" as it goes. Here, for example, is what `simplExpr` does for `App` and `Case` nodes:

```
simplExpr sub ins (App f a) cont
  = simplExpr sub ins f (AppCxt a sub cont)

simplExpr sub ins (Case e b alts) cont
  = simplExpr sub ins e (CaseCxt b alts sub cont)
```

We have already seen how useful it is to know the context of a variable occurrence. The context also makes it easy to perform other transformations, such as the case-of-known-constructor transformation:

```
  case (a,b) of { (p,q) -> E }
==>
  let {p=a; q=b} in E
```

[5] The expression `E1 `seq` E2` evaluates `E1`, discards the result, and then evaluates and returns `E2`.

`simplExpr` just matches a constructor application with a `CaseCxt` continuation.

The next case, `ArgCxt`, is used when simplifying the argument of a strict function or primitive operator. Here, a genuine, functional continuation is used, because no more needs to be known about the continuation.

The `InlineCxt` context is discussed in the next subsection. In practice, GHC's simplifier has another couple of constructors in the `Context` data type, but they are more peripheral so we do not discuss them here.

6.5 `INLINE` *pragmas*

Like some other languages, GHC allows the programmer to specify that a function should be inlined at all its occurrences using a *pragma* in the Haskell source language:

```
{-# INLINE f #-}
f x = ...
```

GHC also allows the Haskell programmer to ask the compiler to inline a function at a particular call site, thus:

```
...(inline f a b)...
```

The function `inline` has type $\forall\alpha.\alpha$ `->` α, and is semantically the identity function. Operationally, though, it asks that `f` be inlined at this call site. Such per-occurrence inline pragmas are less commonly offered by compilers (Bak92).

Both these pragmas are translated to constructors in the `Note` data type, which itself can be attached to an expression (Section 2.3):

```
data Note = ...
          | InlineMe       -- {-# INLINE #-}
          | InlinePlease   -- inline
```

If they are so similar in the Core language, why do they appear so different in Haskell? Haskell allows functions to be defined by pattern-matching, using multiple equations, so there is no convenient syntactic place to ask for `f` to be inlined everywhere. At an occurrence site, however, it is natural just to use a pseudo-function.

The effects of `InlineMe` and `InlinePlease` are as follows:

- The effect of `InlineMe` is to make the enclosed expression look very small, which in turn makes the `smallEnough` predicate reply `True`. When `simplExpr` finds an `InlineMe` in a context where `someBenefit` is `True`, it drops the `InlineMe`, because its work is done.
- The effect of `InlinePlease` is to push an `InlineCxt` onto the context stack. The `smallEnough` predicate returns `True` if it finds such a context, regardless of the size of the expression.

There is an important subtlety, however. Consider

```
g = \a b -> ...big...
{-# INLINE f #-}
f = \x -> g x y
```

and suppose that this is the only occurrence of `g`. Should we inline `g` in `f`'s right hand side? By no means! The programmer is asking that `f` be replicated, but not `g`! The right thing to do is to switch off all inlining when processing the body of an `InlineMe`; when `f` is inlined, then (and only then) `g` will get its chance.

7 Implementing the three-phase algorithm

We can now say what happens when the simplifier encounters a binding

```
let x = E in B
```

with substitution `subst`, and an in-scope set `in-scope`. It uses the three-phase strategy described in Section 5, with the following effect on the substitution and in-scope set.

PreInlineUnconditionally. The substitution is extended by binding `x` to `SuspEx E subst`. The in-scope set is not changed.

PostInlineUnconditionally. The substitution is extended by binding `x` to `DoneEx E'`, where `E'` is the simplified version of `E`. The in-scope set is not changed.

Otherwise. If `x` is not already in scope, the substitution is not changed, but the in-scope set is extended by binding `x` to `E'`. If `x` *is* already in scope, then a new variable name `x'` is invented (Section 4.3); the substitution is extended by binding `x` to `DoneEx x'`, and the in-scope set is extended by binding `x'` to `E'`.

So much for what happens at the *binding site* of a variable. Next we consider what happens at its *occurrence(s)*, which is the third occasion on which the simplifier considers inlining.

When the simplifier encounters the occurrence of a variable, the latter (being an `InVar`) must be looked up in the substitution:

```
simplExpr sub ins (Var v) cont
  = case lookup sub v of
      Just (SuspEx e s) -> simplExpr     s ins e cont
      Just (DoneEx e)   -> simplExpr empty ins e cont
      Nothing           -> callSiteInline ins v cont
```

There are three cases to consider. If the substitution maps the variable to a `SuspEx`, then the simplifier is (tail) called again, passing the captured substitution, and the *current* in-scope set. The substitution and the in-scope set usually travel together, but here they do not. We must use the in-scope set from the *occurrence site* (because that describes what variables are in scope there), and the substitution from the *definition site*.

The second case is when the variable maps to `DoneEx e`. In this case you might think we were done. But suppose `e` was a variable. Then we should consider inlining it, given the current context `cont`, which may differ from that at the variable's definition site. What if `e` was a partial application of a function? Again, the context might now indicate that the function should be inlined. So the simple thing to do is simply to pass `e` to `simplExpr` again.

However, notice that that *the substitution must be discarded at this point* — we pass `empty` to `simplExpr` — because the expression `e` is already an out-expression. To see why, consider this example:

```
\x -> let
```

```
      f = x
in
\x -> ...f...
```

When the binding for `f` is encountered, PostInlineUnconditionally will extend the substitution, binding `f` to `DoneEx x`. When the inner `\x` is encountered, the substitution will again be extended to bind `x` to `DoneEx x1`, because `x` is already in scope. Now, at the occurrence of `f`, we will look up `f` in the substitution, finding `DoneEx x`. We must not definitely not apply the same substitution again, lest we replace `x` by `x1`! The right thing to do is to continue with the empty substitution.

The third case is that the variable `v` might not not be in the substitution at all – for example, it might be a variable that did not need to be renamed. In that case, the next thing to do is to consider inlining it, a task that is performed by `callSiteInline`, which we will discuss next. Before we do, it is worth noticing two further points. First, the substitution is again discarded (i.e. not passed to `callSiteInline`) for the same reason as above. Second, the variable we previously thought of as an `InVar` is now an `OutVar`. This is one reason that `InVar` and `OutVar` are simply synonyms for `Var`, rather than being truly distinct types.

The code is simple enough, but it took us a long time before the interplay between the substitution and the in-scope set became as simple and elegant as it now is.

7.1 Inlining at an occurrence site

Once the simplifier has found a variable that is not in the substitution (and hence is an `OutVar`), we need to decide whether to inline it. The first thing to do is to look up the variable in the in-scope set:

```
callSiteInline ins v cont
 = case lookup ins v of
    Nothing     -> error "Not in scope"

    Just (BoundTo rhs occ lvl)
                |  inline rhs occ lvl cont
                -> simplExpr empty ins rhs cont

    Just other  -> rebuild (Var v) cont
```

If the value information is `BoundTo`, and the predicate `inline` says "yes, go ahead", we simply tail-call the simplifier, passing the in-scope set and the empty substitution (as in the `DoneEx` case of the substitution). In all other cases we give up on inlining. The function `rebuild`, which we do not discuss further here, simply combines the variable with its context.

The `inline` predicate is the interesting bit. It looks first at the variable's occurrence information:

```
inline :: OutExpr -> OccInfo -> Level -> Context -> Bool
inline rhs LoopBreaker lvl cont = False
```

```
inline rhs Once lvl cont        = error "inline: Once"

inline rhs OnceInLam lvl cont   = whnfOrBot rhs &&
                                  someBenefit rhs lvl cont

inline rhs ManyBranch lvl cont  = inlineMulti rhs lvl cont

inline rhs Many lvl cont        = whnfOrBot rhs &&
                                  inlineMulti rhs lvl cont
```

The `LoopBreaker` case is obvious. The `Once` case should never happen, because PreInlineUnconditionally will have already inlined the binding.

The `OnceInLam` case uses the `whnfOrBot` predicate (Section 2.2), to ensure that inlining will not happen if there is any work duplication. However, as noted in Section 2.2, even if the variable occurs just once, it is not always a good idea to inline it. The `someBenefit` predicate is discussed in Section 7.2.

The `ManyBranch` and `Many` cases deal with the situation where there is more than one occurrence of the variable. Both make use of `inlineMulti` to do the bulk of the work; in addition, `Many` uses `whnfOrBot` to avoid work duplication.

Incidentally, since `whnfOrBot rhs` depends only on `rhs`, it is actually (lazily) cached in the `BoundTo` constructor rather than being re-calculated at each occurrence site.

7.2 Checking for benefits

The predicate `someBenefit` tries to guess whether some benefit will arise from inlining the expression in a given context:

```
someBenefit :: OutExpr -> Level -> Context -> Bool
```

It is intended to deal with situations like this:

```
g = \a -> E
f = \xs -> map g xs
```

where we suppose that the only call to `g` is the one in `f`'s right-hand side. Should we inline `g`? No! Then we would get

```
f = \xs -> map (\a -> E) xs
```

and now we will allocate the lambda abstraction `(\a -> E)` each time `f` is called. Nothing gained by inlining, but something is lost.

Here is another example:

```
f = \x -> let g = \y -> E
              h = \a b -> (g y, g z)
          in ...
```

There is very little point in inlining `g` at its two call sites, because we can guarantee that no new transformations (beyond those already performed on `E` itself) will be enabled by doing so; the only saving is the call to `g`, and there is a code duplication cost to pay. How do we know that no transformations will be enabled? Because: (a) the arguments `y` and `z` are lambda-bound and hence uninformative; and (b) the result of both calls are simply stored in a data structure.

So `someBenefit rhs lvl cont` returns `True` if any of the following holds:

(a) The expression `rhs` is a constructor, and the context `cont` scrutinises the constructor with a `case`. In this case, inlining will allow us to eliminate the `case`.

(b) The expression is a function (lambda abstraction), and at least one of the arguments in the context is a non-trivial expression, or is a variable with value information other than `Unknown`.

(c) The expression is a function, and after consuming enough arguments from the context to satisfy the lambdas at the top of the function, the remaining context scrutinises the result with a `case`, or applies the result to further arguments.

(d) The expression is a `Nested` function, and the context has enough arguments to saturate the lambdas in the function. In this case, inlining the function may decrease allocation by eliminating the nested binding (which would otherwise lead to allocation), and without increasing allocation at the call sites. This is the main reason we record the `Level` in the function's `Definition`.

As an example of (d), consider this:

```
f = \xs -> let g = \y -> x+y
           in
           zipWith (\a b -> g b a) xs xs
```

Should we inline `g`? It's calling context is completely uninteresting, but if we inline it we can eliminate `g`'s binding altogether, giving:

```
f = \xs -> zipWith (\a b -> a+b) xs xs
```

If `g` had been a `TopLevel` definition, however, we would have had no allocation benefit to trade for the increase in code size, simply the elimination of the call itself. A `Nested` definition is also likely to have fewer call sites than a top-level one, and we have to inline at all of them to eliminate the definition.

Notice that if a variable is the argument of a constructor, `someBenefit` will return `False`, and so the variable will not be inlined, thus maintaining the trivial-constructor-argument invariant (Section 2.2).

7.3 Inlining multiple-occurrence variables

Now we are left with the case of inlining a variable that occurs many times.

```
inlineMulti :: OutExpr -> Level -> Context -> Bool
inlineMulti rhs lvl cont
    = noSizeIncrease rhs cont
      || (someBenefit rhs lvl cont && smallEnough rhs cont)
```

`smallEnough` is the function that every inliner has: is the function small enough to inline? We discuss `noSizeIncrease` in the next section.

The call to `someBenefit` prevents inlining if there is no benefit (even if `smallEnough` would have returned `True`).

Even if there is no direct benefit, however, it is still worth while inlining the function if the result of doing so is no bigger than the call (App92). That is what the predicate `noSizeIncrease` tests. Again, one might expect this case to be rare, but it isn't. For example, Haskell data constructors are curried functions, but in GHC's intermediate language constructor applications are saturated (Section 2.3). We bridge this gap by producing a function definition for each constructor such as:

```
cons = \x xs -> Cons {x,xs}
```

where the `Cons {x,xs}` is the saturated constructor application. (In reality there are a few type abstractions and applications too, but the idea is the same.) These definitions also make a convenient place to perform argument evaluation (and perhaps unboxing) for strict constructors. For the simple definitions, such as `cons`, it is clearly better to inline the definition, even in the case when `someBenefit` is `False`.

7.4 Size matters

We have now finally arrived at the `smallEnough` predicate, the main aspect of this paper for which there is a reasonable (albeit small) literature. We do not claim any new contribution here, though (unlike some proposals) `smallEnough` is context-sensitive:

```
smallEnough :: Expr -> Context -> Bool
```

For the record, however, the algorithm is as follows. We compute the size of the function body (having first split off its formal parameters, namely the lambdas at the top). From this size we subtract:

- The size of the call.
- An *argument discount* for each argument (extracted from the context) that (a) has evaluation-state information other than `Unknown`, and (b) is scrutinised by a `case`, or applied to an argument, in the function body.
- A *result discount* if `someBenefit` is `True` and the function body returns an explicit constructor or lambda.

Actually, the two discounts are multiplied by a *keenness factor*, to allow us to experiment with the weighting given to the discounts. If the result of this computation is smaller than the *inline threshold*, then we inline the function:

$$\text{Inline the function iff}$$
$$(\text{Body-size} - \text{Size-of-call} - \text{Keenness} * \text{Discounts}) \leq \text{Threshold}$$

The argument discount, result discount, keenness factor, and inline threshold are all settable from the command line, though we expect that only experts will wish to do so. Santos gives more details of GHC's heuristics in Section 6.3 of his thesis (San95).

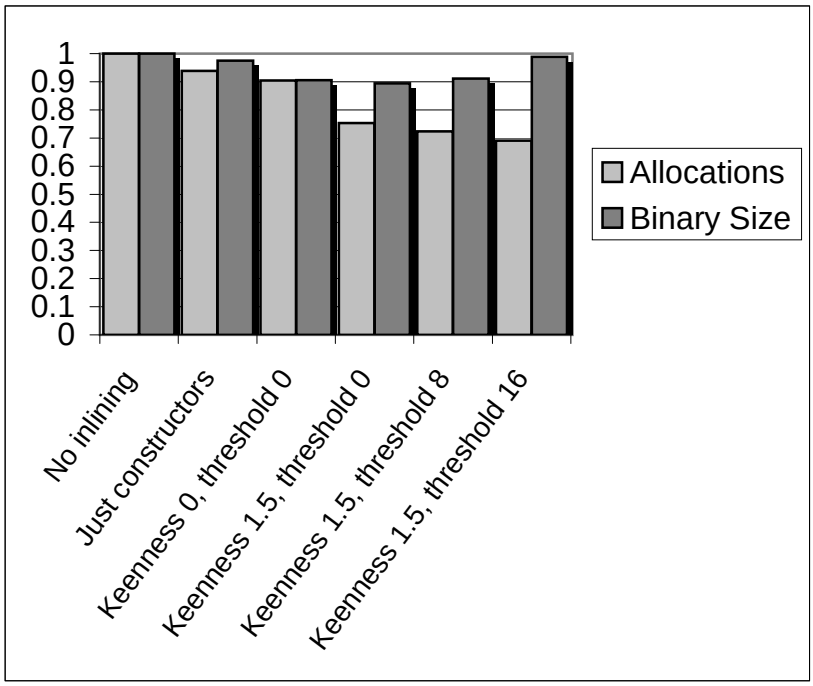


Fig. 5. Effect of inlining threshold on allocations and binary size

7.5 Measurements

As mentioned in Section 7.4, our implementation makes use of an "inline threshold" and a "keenness factor" to determine whether a given expression is small enough to inline. Figure 5 shows the effect of different inlining strategies on the amount of memory allocation done by several test programs, and also the size of the binary generated by the compiler. We measure allocations rather than run-time because we have found that both values tend to vary together, but measuring allocations is repeatable. Our measurements are obtained by taking the geometric mean of the normalised allocations and binary sizes from 18 test programs taken from the "real" section of the nofib test suite.

The strategies we measured were applied to the program only; the libraries in all cases were compiled with the default inlining settings. The measurements we took were:

- No inlining at all. This is likely to produce poor results, because the compiler is built around the assumption that certain essential inlining will be performed, even when all other optimisation is disabled.
- Inlining of constructor wrappers only.
- Keenness zero, threshold zero. This should cause the compiler to inline only when it can be sure that doing so will *reduce* the code size; the keenness of zero means that possible subsequent transformations will not be taken into account.

- Keenness 1.5, threshold zero. Inlinings that reduce the code size may happen, discounts for possible subsequent transformations are taken into account.
- Keenness 1.5, threshold 8 (this is the default setting). Increasing the threshold to 8 allows inlinings that may increase the code size slightly.
- Keenness 1.5, threshold 16.

The actual values for the threshold are fairly arbitrary, and are affected by some of the other parameters: discounts for evaluated arguments and so on.

The results shown here aren't particularly dramatic; this is partly due to the fact that the libraries weren't recompiled each time. Any time spent in library code will be benefitting from the normal inlining behaviour. Nevertheless, we achieve a respectable 30% speedup with the default settings.

There is a rather pronounced jump in performance as soon as we move to a non-zero keeness factor. This appears to be the point at which certain important inlinings (such as integer arithmetic operations) start to happen.

The effects on binary size are small, but there is a pronounced "dip" in the results. Early on, beneficial inlinings lead to further transformations with a net reduction in code size. Beyond a certain point, fewer inlinings are beneficial, leading to an increase in the binary size. Ideally, we should be aiming to achieve maximum performance at the same point as the smallest binary size; in fact the maximum performance point is reached slightly later. This simply means that our heuristics for deciding when an inlining is beneficial aren't perfect; sometimes we have to inline "blind" and see what happens.

8 Related work

There is a modest literature on inlining applied to imperative programming languages, such as C and FORTRAN — some recent examples are (DH92; CMCH92; CHT91; CHT92). In these works the focus is exclusively on *procedures* defined at the *top level*. The benefits are found to be fairly modest (in the 10-20% range), but the cost in terms of code bloat is also very modest. Considerable attention is paid to the effect on architecture-specific effects, such as cache behaviour and register pressure in larger basic blocks, which we do not consider at all.

It seems self-evident that the benefits of inlining are strongly related to both language and programming style. Functional languages encourage the use of abstractions, so the benefits of inlining are likely to be greater. Indeed, Appel reports benefits in the range 15-25% for the Standard ML of New Jersey compiler (App92), while Santos reports average benefits of around 40% for Haskell programs (San95). Chambers reports truly dramatic factors of 4 to 55 for his SELF compiler (Cha92); SELF takes abstraction very seriously indeed!

The most detailed and immediately-relevant work we have found is for two Scheme compilers. Waddell and Dybvig reports performance improvements of 10-100% in the *Chez Scheme* compiler (WD97), while Serrano found a more modest 15% benefit for the Bigloo Scheme compiler (Ser95; Ser97). Both use a dynamic, effort/size budget scheme to control termination. The *Chez Scheme* inliner uses an explicitly-encoded context parameter that plays exactly the role of our `Context` (Section 6.4).

A completely different approach to the inlining problem is discussed by Appel and Jim (AJ97). In this paper the focus is on inlining functions that are called precisely once, something that we have been very concerned with. Appel and Jim show that this transformation, along with a handful of others (including dead-code elimination), are normalising and confluent, a very desirable property. Their focus is then on finding an efficient algorithm for applying the transformations exhaustively. Their solution involves adjusting the results of the occurrence analysis phase as transformations proceed. Their initial algorithm has worst-case quadratic complexity, but they also propose a more subtle (and unimplemented) linear-time variant. We too are concerned about efficient application of transformation rules, but our set of transformations is much larger, and includes general inlining, so their results are not directly applicable to our setting. Nevertheless, it is a unique and inspiring approach.

Copious measurements of many transformations in GHC (not only inlining) can be found in Santos's thesis (San95); although these measurements are now several years old, we believe that the general outlines are unlikely to have changed dramatically. Another paper contains briefer, but more up-to-date, measurements (PJS98).

This paper has focused on the task of compiling a *lazy, purely functional* language, but almost all of it applies to *strict, impure* languages as well. From an inlining point of view the principal difference is that only *values* (lambda abstrac-

tions, constructors, variables) can be inlined freely. Non-values need more careful treatment:

```
let x = f y in ...x...
```

Even if `x` occurs exactly once in its scope, it is only valid to inline it if function `f` has no side effects. Sophisticated compilers for such languages may perform an effects analysis to identify such functions.

9 Conclusion

This paper has told a long story. Inlining seems a relatively simple idea, but in practice it is complicated to do a good job. The main contribution of the paper is to set down, in sometimes-gory detail, the lessons that we have learned over nearly a decade of tuning our inliner. Everyone who tries to build a transformation-based compiler has to grapple with these issues but, because they are not crisp or sexy, there is almost no literature on the subject. This paper is a modest attempt to address that lack.

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