Annotating Deeply Embedded Languages

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June 1, 2022

This thesis explores the idea of annotating deeply embedded languages with additional information. A common problem with deeply embedded languages is that there is no strong connection between an embedded program, the abstract syntax tree for that embedded program, and a compiled version of the program. By gathering source location information for embedded expressions using a novel implicit parameter-based approach and annotating the embedded program's AST with this information, it becomes possible to map the embedded program back to the original source code. This new information can be used to improve compiler diagnostics, and to provide better profiling and debugging experiences for the language. In addition, the idea of extending that approach by also annotating the language with instructions for the compiler as a way to allow the programmer to hand optimize parts of a program will be explored. Storing these instructions as annotations on the abstract syntax trees enables new optimization workflows without requiring any changes to the language itself. Finally, an implementation of this annotation system in an existing language is examined, and its consequences on the experience of using the language are evaluated.

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1 Introduction

Embedded languages are programming languages that are written inside of another language, called the host language. These languages may be domain specific, and because they are embedded in an existing language they are able to leverage most of the host language's features and tooling. Within embedded languages there is a distinction between shallow embeddings and deep embeddings. The difference between those is that shallowly embedded languages immediately compute a result as the embedded program is evaluated, while deeply embedded languages merely produce a representation of the embedded program in the form of an abstract syntax tree (*AST*). This AST can later be type checked, optimized, and eventually interpreted or compiled, just like a compiler would be able to do with the AST from a standalone parsed language. Because deeply embedded languages reuse the syntax and features from the host language, they can integrate better with other programs written in the host language than an ordinary standalone language would. But this approach also comes with its share of drawbacks. One of those drawbacks is that the language is restricted to the syntax and semantics of the host language, as embedded languages do not have a separate parsing step. A consequence of not having this parsing step is that there is also little connection between the source code the programmer originally wrote, the AST the embedded language's compiler needs to compile, and an executable binary for the compiled embedded program. This makes it difficult to add descriptive error messages and compiler diagnostics, and debugging and profiling the compiled program also become more difficult.

The goal of this thesis is to solve these issues by extending an existing deeply embedded language with an annotation system. This annotation system is then used to make it possible to map parts of the AST back to the original program. Once that is in place, the idea of extending the annotation system to provide other language features will be explored. The main focus there will be on allowing the programmer to manually influence the compiler's optimization process, either globally or on a smaller local scope. Lastly, an implementation of this annotation system in an existing and more complex language will be examined along with its effect on the experience of using that language.

2 Background

Before taking a look at how an annotation system for a deeply embedded language could work, it is important to first have a basic understanding of the some underlying concepts. All code examples in this thesis are written in Haskell for the GHC compiler. It is also be possible to implement these techniques in any other language as long as it supports the same concepts.

2.1 Implicit parameters

Implicit parameters [1] is a programming language feature that has been implemented in Haskell through a GHC extension. It forms the basis of a number of the algorithms and techniques described in this thesis, so it is important to be familiar with the concept and its uses. In terse jargon, implicit parameters are statically typed dynamically scoped bindings. This is similar to the `(let (...)  ...)` construct in older lisp dialects. These implicit parameters allow functions to define a set of special variables that need to be set from an outer scope before the function is called. The function can then access those variables almost as if they were regular bindings. This is semantically equivalent to adding those variables as function parameters and manually passing them around, except that all of this is done implicitly. Hence, they are called implicit parameters. The syntax for implicit parameters in Haskell involves a new form of constraints and let bindings shown in Listing 1.

```haskell
addConstant :: (?constant :: Int) => Int -> Int
addConstant n = n + ?constant

addFour :: Int -> Int
addFour n = let ?constant = 4 in addConstant n
```

Listing 1: Implicit parameters in Haskell.

All of this may not look particularly useful at first, but it can be used as a less verbose alternative to the *reader monad* that was first proposed by Jones [2] as an abstraction to allow computations to access values held by their enclosing environment. Some common use cases for these reader monads are passing a fixed value down a function call tree, and having each function in that call tree alter the value before it gets passed to the next function. Haskell's implicit parameters feature is semantically equivalent to the reader monad, but without having to wrap a computation in a monadic action. And because they are statically typed, forgetting to set an implicit parameter or setting implicit parameters that don't exist results in a compile-time type error.

A downside of using these implicit parameters is that since they are statically typed, every function in a function calling sequence needs to have the implicit parameter constraint. Otherwise it is no longer possible to propagate a value from the top of the function call tree down to the bottom. This makes sense given the implementation, but it can cause confusing situations in actual usage. Similarly to dynamically scoped bindings in other languages, implicit parameters are only bound for a term when that term is *used* directly within the body of an implicit parameter let-binding, regardless of where it was *defined*. This results in a subtle but important interaction with the way implicit parameter values are propagated through a function calling sequence when each level may want to modify the value before passing it on to the next function. This is because helper functions defined in the where-clause of a function do not automatically have any implicit parameters constraints added to their type signature. As a result of this, unless the implicit parameter was explicitly added to that helper function, the function will always inherit the implicit parameter's value from the surrounding function, even if an implicit parameter let-binding was used to bind a new value to the implicit parameter before calling the helper function. This is a somewhat confusing but very important detail that will become significant when examining an approach for capturing source location information later in this thesis.

2.2 Call stacks

To capture source location information for an embedded program, there needs to be a way to represent the flow of the function calls throughout a program. Call stacks are commonly used to keep track of how a program calls function at runtime. These call stacks are snapshots of the program's function call graph or call tree. This is usually implemented by keeping track of a stack of *stack frames*, which describe the state of the active subroutines that have not yet been returned from. These are commonly used by interpreters and runtimes to provide backtraces during runtime exceptions, by profilers to trace a program's execution within nested function calls, and by debuggers to allow stepping in to and out of functions. GHC offers multiple ways to interact with these call stacks from within Haskell itself. These are used by libraries to show backtraces during fatal errors, but the programmer can also use them directly to access information about the current function call and the execution path leading up to it. The two main ways call stack information can be accessed in a Haskell program are *GHC Call Stacks* [3] and *RTS Execution Stacks* [4].

2.2.1 GHC Call Stacks

GHC Call Stacks are Haskell's main call stack interface. They are accessible through the `GHC.Stack` module in the `base` package. The compiler creates these call stacks at compile time for every function has the special `HasCallStack` constraint, which in turn uses an implicit parameter called `?callStack`. Each time a function with the `HasCallStack` constraint gets called, the called function is pushed onto the call stack stored in the implicit parameter. Those call stacks can then be read directly from that implicit parameter (or more commonly, using the `callStack` function) whenever they are needed. For example, the `error` function, which prints a fatal error and then exits the program, uses these call stacks to include context information when printing the error message. Similarly, the Hedgehog [5] automated property testing framework uses these call stacks to show where in the test's source code assertions are failing. The contents of these call stacks are determined at compile time based on the `HasCallStack` constraints. Therefore, using them does not carry a significant runtime performance cost which makes them very suitable for a lightweight source mapping implementation. However, this compile time approach comes with some caveats, as it requires every function in a calling sequence to be annotated with the `HasCallStack` constraint. Otherwise the call stack will end at the bottommost function in the call tree that does not have this constraint.

There is an important takeaway here which will become relevant later. While these GHC Call Stacks are backed by an implicit parameter, this is mostly an implementation detail. They end up behaving very differently from ordinary implicit parameters. When calling a `HasCallStack` function, the implicit parameter's value is automatically filled in by the compiler. As a result, they behave as if they have a default value, and both the use and the absence of the constraint is completely transparent to the programmer.

Because of this, it can be difficult to assert that every function in a call tree actually has been annotated with the `HasCallStack` constraint, and that no information has been lost. A solution for this will be presented later in this thesis.

2.2.2 RTS Execution Stacks

The alternative approach to call stacks in GHC makes use of debug symbols embedded in the compiled Haskell program to create call stacks at runtime. This feature is called *RTS Execution Stacks* [4]. The main downside of this feature, and the reason why these execution stacks are generally less applicable than GHC Call Stacks, is that support for this feature requires a relatively obscure non-default configuration option to be enabled when compiling GHC. Additionally, these execution stacks may not always contain the relevant stack frames when compiling Haskell with optimizations. As an example, GHC may try to inline functions or optimize away tail calls, both of which also get rid of the stack frames for those functions. Another issue is that, because they are a relatively obscure feature, in their current form the functions used to access these execution stacks contain a lot of noise that would need to be filtered out first. This makes extracting relevant information from these stacks difficult. All in all this means that RTS Execution Stacks cannot be relied upon for general use, but with a bit of effort they can still be used as a fallback whenever GHC Call Stacks are not an option.

2.3 A deeply embedded language

Before it is possible to take a look at how a deeply embedded language can be annotated, there first needs to be a language to reason about and experiment with. The language described in this section is a higher-order typed expression language with user-definable functions and support for recursion. The most important detail to keep in mind here is that programs written in deeply embedded languages simply evaluate to abstract syntax trees when executing the host language program. These ASTs can then be passed to some function provided by the language's implementation where they are optimized and either evaluated directly or compiled to native code in a separate runtime compilation step. As a result, a deeply embedded language needs to define the structure of those ASTs and provide ways to interact with them. Because writing a program's AST by hand is not going to make the language very usable, the following sections will first introduce the language's abstract syntax, followed by the concept of *smart constructors* to help the user build these ASTs. Those hide the language's actual AST types and internals from the user. To keep thing simple, the language reuses Haskell's type system and type classes instead of creating new ones that might be more suitable for the language's domain. Reusing Haskell's type system makes compiling embedded programs to something that can run as an efficient native binary outside of Haskell more difficult, but it also makes understanding and working with the language simpler, and it tends to be the canonical

```haskell
data Exp a where
  -- Lifts a constant term into the embedded language
  Const      :: a -> Exp a
  -- References the value from a bound variable
  Var        :: Ident t -> Exp t -- Unsafe, requires separate type checking
  -- Binds a value to a variable identifier before executing a body expression
  Let        :: Ident t -> Exp t -> Exp b -> Exp b
  -- Groups two values into a pair
  Pair       :: Exp a -> Exp b -> Exp (a, b)
  -- Fetches the left value from a pair
  PrjL       :: Exp (a, b) -> Exp a
  -- Fetches the right value from a pair
  PrjR       :: Exp (a, b) -> Exp b
  -- Calls an embedded function with an argument
  Apply      :: Exp (a -> r) -> Exp a -> Exp r
  -- Defines an anonymous embedded function with the argument bound to an identifier
  Lambda     :: Ident a -> Exp r -> Exp (a -> r)
  -- References a built-in primitive function
  Builtin    :: BuiltinFn (a -> r) -> Exp (a -> r)
  -- Evaluates either the first or the second arm based on a boolean condition
  IfThenElse :: Exp Bool -> Exp a -> Exp a -> Exp a

-- A typed variable identifier to allow static type checking
data Ident t where
  Ident :: Typeable t => String -> Ident t

-- Built-in primitive functions where 'sig' is the function's signature
data BuiltinFn sig where
  BuiltinAdd  :: Num a => BuiltinFn ((a, a) -> a)
  BuiltinMul  :: Num a => BuiltinFn ((a, a) -> a)
  BuiltinAbs  :: Num a => BuiltinFn (a      -> a)
  BuiltinSign :: Num a => BuiltinFn (a      -> a)
  BuiltinNeg  :: Num a => BuiltinFn (a      -> a)
  BuiltinLte  :: Ord a => BuiltinFn ((a, a) -> Bool)
```

Listing 2: A higher-order expression level with user defineable functions and lexically scoped let-bindings.

way to handle types for languages embedded in Haskell. For now the focus will be on the language's semantics rather than the low-level implementation details.

2.3.1 Abstract syntax

Listing 2 shows the abstract syntax of the deeply embedded example language as a type-indexed generalized algebraic data type [6]. The language contains terms that can be either constants, anonymous functions, bound variables, or tuples of those. As the value of a variable may not be statically known at compile-time, the language has the `PrjL` and `PrjR` projection constructors to extract the left and right element of a pair. Values can be bound to a variable using lexically scoped let-bindings, and the values of variables bound within the current scope can be referenced using the `Var` constructor. All variables are statically typed, and when compiling an embedded program a type checker would need to check whether all variables used within the program exist in the current scope and are of the correct type.

The language is higher-order, which means that it includes support for functions that can take other functions as an argument. Deeply embedded languages usually implement this in one of two ways: either the abstract syntax is *first-order* and it contains constructs for binding variables along with other constructs for using those bindings, or it is *higher-order*, where there are no explicit labelled binding sites and each use of a variable refers directly to the binding. Keep in mind that the terms first-order and higher-order here refer only to the way variable bindings are represented, this is unrelated to whether or not the language supports higher-order functions. The presented language uses a first-order abstract syntax (*FOAS*). Using a higher-order abstract syntax (*HOAS*) instead has a couple subtle implications for the implementation of an annotation system. These will be discussed later when examining the implementation of an annotation system in such an embedded language.

Being first-order, the language's anonymous functions can either be used directly or they can be bound to identifiers using let-bindings. Both the anonymous functions and functions bound to variables can be called to produce a result. Because of the way bindings work in this language, it also inherently supports recursive functions and currying. Finally, the language comes with a selection of built-in primitive functions as well as a conditional to allow the language to do something useful.

As an example, a function that calculates the factorial of some number using recursion might look similar to the program shown in Listing 3:

2.3.2 Smart constructors

Because writing the AST from the previous example by hand is both tedious and error-prone, the language needs to come with some form syntactic sugar to make it easier and more ergonomic to write programs in the embedded language. This syntactic sugar is provided in the form of *smart constructors*, which are simply functions that construct AST nodes. The resulting code not only looks nicer, but these smart constructors can also assert invariants for the embedded program, provide syntactic sugar for complex

```haskell
factorial :: (Num a, Ord a, Typeable a) => Exp a -> Exp a
factorial n =
  Let factorial' (Lambda n'                          -- let factorial' = \n' ->
        (IfThenElse (Apply (Builtin BuiltinLte) --       if n' <= 1
                           (Pair (Var n') (Const 1)))
            (Const 1)                           --         then 1
          (Apply (Builtin BuiltinMul)           --         else
                 (Pair (Var n')                 --           n'
                       (Apply (Var factorial')  --             * factorial' (n' - 1)
                              (Apply (Builtin BuiltinAdd)
                                     (Pair (Var n') (Const (-1)))))))))
    (Apply (Var factorial') n)                  -- in  factorial' n
  where
    factorial' = Ident "factorial"
    n'         = Ident "n"
```

Listing 3: A recursive factorial implementation, written as a raw AST without any syntactic sugar. The Haskell equivalent is shown in the comments on the right.

language constructs, and express concepts that are not directly part of the language. The rest of this section will discuss three distinct forms of smart constructors: regular functions, type class implementations, and pattern synonyms. Each form will end up having its own role when adding annotations to the language later.

2.3.2.1 Function based smart constructors

The first type of smart constructor, implemented through ordinary functions, can be seen in Listing 4. In their simplest form, these smart constructors directly map their arguments to AST nodes. Smart constructors can also be used to add syntactic sugar for higher level language constructs. For instance, the `letfn` function from the listing defines a smart constructor that binds a potentially recursive function to a variable, which can then be used to call that function inside of the let binding's body.

2.3.2.2 Type class instance smart constructors

Next are smart constructors that repurpose common type classes from Haskell's standard library. An example for the `Num` type class can be seen in Listing 5. This approach lets the embedded language integrate better with the host language, since any functions that generalize over this type class will now also work with the embedded language. In this particular case, implementing the `fromInteger` function also makes it possible to create numeric constants in the embedded languages by simply using literal constants in the

```haskell
cond :: Exp Bool -> Exp a -> Exp a -> Exp a
cond = IfThenElse

-- Alternatively, the language could define its own Ord-like type class
(<=) :: Ord a => Exp a -> Exp a -> Exp Bool
a <= b = Apply (Builtin BuiltinLte) (Pair a b)

letvar :: Typeable a => Exp a -> (Exp a -> Exp r) -> Exp r
letvar lhs body = Let ident lhs (body (Var ident))
  where
    ident = Ident (...)

letfn
  :: forall a r b
   . (Typeable a, Typeable r)
  => ((Exp a -> Exp r) -> Exp a -> Exp r)
  -> ((Exp a -> Exp r) -> Exp b)
  -> Exp b
letfn mkFn body = Let fnIdent (Lambda argIdent (mkFn callFn (Var argIdent)))
                      (body callFn)
  where
    fnIdent = Ident (...)
    argIdent = Ident (...)
    callFn :: Exp a -> Exp r
    callFn = Apply (Var fnIdent)
```

Listing 4: Smart constructors as functions that directly map to AST nodes.

```haskell
instance Num a => Num (Exp a) where
  x + y       = Apply (Builtin BuiltinAdd) (Pair x y)
  x * y       = Apply (Builtin BuiltinMul) (Pair x y)
  abs         = Apply (Builtin BuiltinAbs)
  signum      = Apply (Builtin BuiltinSign)
  fromInteger = Const . fromInteger
  negate      = Apply (Builtin BuiltinNeg)
```

Listing 5: Smart constructors as instances of existing type classes.

source code. Haskell implicitly uses that function to lift those values into embedded terms. This approach to smart constructors needs to be treated separately from the other smart constructors that were defined using free functions, as there is no control over the functions' signatures in this situation. This becomes relevant later when looking at gathering source locations.

2.3.2.3 Pattern synonym smart constructors

```haskell
pattern T2 :: Exp a -> Exp b -> Exp (a, b)
pattern T2 x y <- (\p -> (PrjL p, PrjR p) -> (x, y))
  where T2 x y = Pair x y
```

Listing 6: Smart constructors as pattern synonyms. This uses the *view patterns* language extension to destructure a pair into two terms.

The last type of smart constructor the language uses is implemented through *explicitly bidirectional pattern synonyms* and can be found in Listing 6. This form of smart constructors was introduced in the work of McDonell et al. [7] and it makes it possible to ergonomically work with embedded tuples and other compound data structures. This pattern synonym for the listing implements this only for pairs. It is also possible to create similar pattern synonyms for tuple of any size by nesting pairs within pairs. These tuple pattern synonyms not only make it possible to build tuples, but they can also be used on the left-hand side of a binding as a pattern to destructure tuples. Doing so creates *projections* for each element in the tuple that makes it possible to access those elements. Because unused code does not end up in the program's final AST, this approach has no runtime performance overhead.

With these concepts in place, it is now possible to write programs in this embedded expression language that look somewhat similar to a regular functional program, but with syntactic constructs and keywords replaced by functions. Listing 7 contains the exact same factorial function from Listing 3, but this time using the syntactic sugar that was defined in this section.

3 Annotations

Adding annotations to the language defined in Section 2.3 is done in two stages. This section discusses the structure of the annotations and how annotating parts of a program would work, while the next section covers several ways these annotations can be used.

3.1 Storing annotations

The end goal this thesis is working towards is to be able to extend an implementation of an existing deeply embedded language by annotating it with source locations and other information that may be useful to the compiler. To achieve this, this section compares a couple different ways annotations can be represented and stored in a deeply

```haskell
-- The embedded version
factorial :: (Num a, Ord a, Typeable a) => Exp a -> Exp a
factorial n = letfn (\factorial' n' ->
                       cond (n' <= 1) 1 (n' * factorial' (n' - 1)))
                 (\factorial' -> factorial' n)
```

```haskell
-- And a plain Haskell factorial function following the same structure
factorial :: (Num a, Ord a) => a -> a
factorial n =
  let factorial' n' = if n <= 1
                        then 1
                        else n' * factorial' (n' - 1)
   in factorial' n
```

Listing 7: The recursive factorial implementation from Listing 3 implemented using the smart constructors defined above, as well as a plain Haskell equivalent of the same program.

embedded language. An important consideration here is that it may be desirable for the implementation to require as few changes to the language's existing implementation as possible. Or alternatively put, an implementation may want to try and minimize the cognitive overhead and the maintenance burden added by the annotation system. In addition to that, existing programs written in the language should not require any modifications to keep working. The language defined in the previous sections is used as the running example, and later on these ideas are transferred to apply to an already existing and much more complex language.

There are several ways these annotations can be added to the language. The *Trees that Grow* paper by Najd and Jones explores this concept within the context of the GHC Haskell compiler itself [8]. As the goal is to extend an existing language with as few changes as possible, the simplest way to achieve this would be to add the annotation data directly to the AST constructors. This would work by defining an annotation data type as a record, and then simply adding fields for that type to each AST constructor. That way only the use sites for these constructors need to be modified, and once the initial work is done it becomes possible to add any additional information to the AST as needed without having to modify anything other than the annotation data type. An example of this can be found in Listing 8.

While this simple approach aligns well with the goal this thesis, there are a number of other options that may be worth considering. One of those is to use a mutually recursive data types for the AST and the annotations. This avoids the need to add a field to every constructor, but the drawback is that it makes interacting with the AST types in the compiler less straightforward. To implement this, a type parameter that refers to the expression type needs to be added to the annotation type. Similarly, the expression type

```haskell
data Ann = Ann { ... }

data Exp a where
  Const :: Ann -> a -> Exp a
  Var   :: Ann -> Ident t -> Exp t
  Let   :: Ann -> Ident t -> Exp t -> Exp b -> Exp b
  ...
```

Listing 8: Part of the abstract syntax that was defined in Section 2.3.1, but with annotations fields added. Each constructor now has a field for storing an `Ann` value.

needs to be modified to use a new higher-kinded type parameter for its nested expressions instead of directly reusing the expression's own type there. This type parameter can then be used to refer to the expression type wrapped by an annotation type. As these concepts are very abstract, Listing 9 contains an example of how this would work in practice. In summary, this change turned the expression type from an inductive type to a recursive type. Doing so avoids having to add annotation fields to every constructor, as the annotations and AST node constructors are always interleaved. Another benefit of recursive data types is that they can be trivially parameterized over different types, with the downside being that it makes working with these types more complicated. Most concretely, where the AST previously only required a single `Exp` constructor per nesting level, a single level now consists of an `AnnotatedExp` constructor, an `Exp` constructor, an `ExpAnn` constructor, and an `Ann` constructor.

```haskell
data Ann exp = Ann exp ...

data Exp (exp :: * -> *) a where
  Const :: a -> Exp exp a
  Var   :: Ident t -> Exp exp t
  Let   :: Ident t -> exp t -> exp b -> Exp exp b
  ...

newtype ExpAnn a = ExpAnn (Ann (Exp AnnotatedExp a))
newtype AnnotatedExp a = AnnotatedExp (Exp ExpAnn a)
```

Listing 9: The annotated abstract syntax from Listing 8, but using recursive data types. In this version the `Exp` data type takes a new higher kinded type parameter that is used wherever `Exp` was previously used as part of a constructor. An example of this is shown in the new `Let` constructor.

3.2 Annotating AST nodes

The previous section introduced data types and constructor fields for storing annotations. With these in place, the next step is to fill those fields with annotation data. This is where the smart constructors from Section 2.3.2 come in. Conceptually, the idea is very simple: add the annotation fields, and then modify the smart constructors to store fresh annotation objects in those fields whenever an AST node is created. That way the annotations can be added to the language without requiring any changes to user written code. Later sections will expand on this idea to be able to capture more information in the annotations. For now however, the only change required to the code is the introduction of a `mkAnn` function which creates fresh annotations. This function then needs to be called from the smart constructors, as shown in Listing 10.

```haskell
mkAnn :: Ann
mkAnn = Ann { ... }

cond :: Exp Bool -> Exp a -> Exp a -> Exp a
cond = IfThenElse mkAnn

(<=) :: Ord a => Exp a -> Exp a -> Exp Bool
a <= b = Apply mkAnn (Builtin BuiltinLte) (Pair a b)
```

Listing 10: Some of language's smart constructors from Section 2.3.2, but with fresh annotations added. The implementations for the other smart constructors follow the same pattern.

There is one last part of the puzzle here: adding ways to modify these annotations. In non-embedded languages this is done a variety of ways including attributes, decorators, and pragmas. But since it is usually not possible to add additional syntax to the host language, decorators are the most suitable approach to allow the programmer to annotate AST nodes with additional information. Decorators are simply functions that modify individual objects or functions without modifying all instances of those. In this case, the decorator would simply modify the annotation object stored in the AST node and leave the rest of the AST intact. These decorators can be generalized to any type that can be annotated by using a type class like the one from Listing 11. Decorator functions that modify part of an AST can then build upon the `modifyAnn` function from the listing to be able to work with any AST type. Notice how there is not only a type class implementation for the `Exp` ASTs, but also an implementation for functions returning annotateable types. The ability to decorate functions in addition to concrete ASTs creates a new class of combinators for building modified versions of existing functions and smart constructors with different behavior. When a more complex language has multiple AST types, then this type class can also generalize over those different types and over functions that construct them.

```haskell
class HasAnnotations a where
  modifyAnn :: (Ann -> Ann) -> a -> a
  getAnn    :: a -> Maybe Ann

instance HasAnnotations (Exp a) where
  modifyAnn f (Const ann x)    = Const (f ann) x
  modifyAnn f (Var ann ident) = Var (f ann) ident
  ...

instance HasAnnotations r => HasAnnotations (a -> r) where
  modifyAnn f f' x = modifyAnn f (f' x)
  -- It is not possible to get the annotation without evaluating
  -- the function first
  getAnn _ = Nothing
```

Listing 11: Type classes for modifying annotations for `Exp` ASTs and functions returning those ASTs.

3.3 Annotating AST subtrees

A special case of modifying the annotations stored in an AST is when one wishes to modify an entire subtree at once. There are several situations where this might be desirable. For instance, the annotations may store an expression level configuration flag that needs to be adjusted for all expressions in an entire subsection of the program. Or the annotations may store context information that needs to be changed as part of a compiler pass, for instance when the compiler rewrites file paths.

There are two main ways to go about this: the AST can either be rewritten directly in-place such that every node is replaced by an updated version, or the modification can be stored in a new AST constructor and processed later when the program gets compiled or interpreted. At a first glance both approaches may look similar, but there are some important differences under the hood. First of all, the evaluation order for the modifications would be different. Consider an AST with several levels of nested expressions, each interspered with subtree modifications. When the AST is rewritten in-place as part of a decorator function, then these modifications are evaluated from the bottom of the tree to the top. Modifications closer to the root of the tree will overwrite any changes previously made in the subtree. By contrast, storing the modification as a new type of AST node that wraps around the subtree for later allows the modifications to be performed from top to bottom. This makes it possible to change a setting stored for a subtree of the program's AST using a decorator, and then later override the same setting again with a different value for a smaller portion of that subtree.

There is also another important difference: rewriting the AST in-place changes every node in that AST, while simply wrapping the existing AST in a new node does not cause its children to be modified. This may sound like stating the obvious, but the distinction becomes important when implementing this system as part of a real deeply embedded language. The significance of this will be explained later when taking a look at how implementing an annotation system would work with a real, already existing embedded language. For now it is enough to realize that rewriting terms like this will cause terms that were once shared using a let-binding now no longer refer to the same variable as both uses have been modified independently.

For these reasons, the second approach of using a new constructor and delaying the modifications until the language gets compiled or interpreted is usually preferable over directly rewriting the AST. The extensions to the language needed to make this work can be found in Listing 12, which introduces a type class that's similar to the `HasAnnotations` from the previous listing but that modifies entire subtrees instead of single AST nodes. Since with this appraoch no actual work has to be done before the program is compiled, the implementation simply wraps AST nodes in a new constructor containing annotation data that should then later be applied to the subtree. An alternative implementation choice would be to define a new data type that only describes the changes that need to be made to an annotation instead of repurposing the actual annotation data type. To keep things simple, the implementation given here will not do this, but with a more complex annotation type this might be worth considering.

```haskell
data Exp a where
  AnnSubtree :: Ann -> Exp a -> Exp a
  ...

class TraverseAnnotations a where
  annotateSubtree :: Ann -> a -> a

instance TraverseAnnotations (Exp t) where
  annotateSubtree ann e = AnnSubtree ann e

instance TraverseAnnotations r => TraverseAnnotations (a -> r) where
  annotateSubtree ann f' x = annotateSubtree ann (f' x)
```

Listing 12: Type classes for modifying annotations for entire subtrees of `Exp` ASTs and functions returning those ASTs.

4 Using annotations

With the basis for an annotation system in place, it is now time to look at how this system can be used to improve the language. The following sections will examine how the annotation system can be used to automatically annotate the AST with source location information, and how the annotations can be repurposed to store local compiler configuration and optimization flags that can be set on the expression level using decorators. The source location information will help improve the language's usability by enabling better debugging options and diagnostics, while the optimization flags can help unlock more performance by allowing the programmer to make informed decisions for the language's optimizer.

4.1 Source mapping

As mentioned earlier, one major problem affecting many deeply embedded languages is that they can be difficult to debug and work with when things are not working as they should. Most of these problems stem from j0j the disconnect between the embedded program as written by the user and the compiled or interpreted version of that program as it is being executed by the language's runtime. This also means that whenever problems occur during the compilation of the program, such as references made to undefined variables, then those problems can only be communicated to the user using generic warnings and error messages. Ideally, those diagnostics would instead refer to the exact place in the source code where the issue occurred, similar to how most regular non-embedded languages work. Additionally, when the program does compile correctly and the compiled program is run, it would be useful to have more insight into the program's execution. For instance, when a program runs slower than expected then the first thing one should do is to run the program under a profiler. The problem is, if a profiler cannot inform you where exactly in the program most of the execution time happens, then there is still no information to go on. A third situation where the disconnect between source code and a compiled binary becomes obvious is when debugging a malfunctioning program. If the embedded language contains print statements then you will be able to get relatively far in debugging misbehaving programs, but there are times when stepping through the execution of a program is the only way to figure out what is going wrong. And without any form of debug symbols providing source information, that can be a difficult exercise.

To address the above problems and other related obstacles, the language's AST needs to contain information about the original program, and it needs to be possible to map parts of the compiled or interpreted program back to the original source code using that information. In the context of a language deeply embedded in Haskell, the simplest way to get this information would be to try and capture that information from within the language's smart constructors. To achieve this, the most suitable method would be to build upon the GHC Call Stacks feature discussed in Section 2.2.1. As this cannot always be used, the sections after this one will discuss alternative approaches for when this is not an option, or when it may not work as expected.

Because GHC Call Stacks end up forming the base for this source mapping system, it also makes sense to store the source location information in the same format. Or rather, to store sets of these call stacks, as an optimized program may contain AST nodes that are the result of combining AST nodes of different origins. While most uses of this source mapping information can only make use of a single call stack, storing this as a set has two main advantages. The obvious advantage for the user is that purpose made tooling would be able to show all source locations even if they are disjoint, but there is also a more technical advantage to this approach. By storing GHC's call stacks without modifying them, Haskell's sharing can make it so that multiple references to the same call stack can be shared in memory. They then only need to be merged into a smaller set of adjacent call stacks at the very end of the pipeline when they are used. The last point that needs some explanation is why the annotation would need to store entire call stacks rather than just the source locations in the user's code where the language's smart constructors were called from. Because these GHC Call Stacks require Haskell code to be annotated with the `HasCallStack` constraint, most captured call stacks will not end up containing any information beyond the calling location of the smart constructor. As a result, storing entire call stacks is not likely to result in significant increases in memory usage. But at the same time, storing entire call stacks does allow the user to manually add more context information to the captured call stacks by adding the `HasCallStack` constraint to their own Haskell code. And finally it also becomes possible to define new decorators that inject new call stack information into AST nodes without touching any of the existing source location information. This idea will be explored later as an alternative for when it is not possible to automatically capture sufficiently granular source location information.

There is one more piece to the puzzle than just storing the call stacks provided by GHC Call Stacks. As just discussed, GHC Call Stacks only provides information for functions annotated with the `HasCallStack` constraint. While that constraint technically uses an *implicit parameter* under the hood, this implicit parameter is an implementation detail of the GHC compiler and does not require the caller of such a function to do anything to make it work. That means that a lot of precaution must be taken, as it is easy to forget to add the `HasCallStack` contraint to helper functions that are called from the smart constructor that may end up directly or indirectly creating annotations. This would cause the call stacks to be lost with no way to tell other than to manually inspect the call stacks captured in the embedded program. Another related problem is that these helper functions should not show up in the captured call stack, as they are implementation details that are not relevant to the end user. This last problem can be solved by *freezing* the call stacks. This means that GHC will stop appending additional frames to the call stack. But since the first problem still exists, forgetting a `HasCallStack` constraint on a function will also unfreeze the call stack again.

Both of these problems can be solved at the same time by introducing a new implicit parameter-based constraint to act as a substitute for `HasCallStack`. This involves a novel approach to encode a certain degree of integrity in the type system to enforce that a calling function *must* provide a frozen call stack to the callee. The new constraint would also imply the existing `HasCallStack` so that all of the existing call stack infrastructure can be reused, but the use of a new implicit parameter ensures that the caller of a function annotated with the constraint either needs to have the same constraint, or that the caller must set that implicit parameter before calling the function. This constraint is defined in Listing 13 as the `SourceMapped` constraint alias. The single most important piece of the puzzle here is the data type of the new implicit parameter. By defining that type as a new data type with single constructor and then not exposing that constructor to other modules, it becomes possible to enforce that any function calling a `SourceMapped` function must go through one of the predefined ways to instantiate a source mapped context. In the listing this takes the form of the new `sourceMap` function, which freezes the current call stack, sets up the `SourceMapped` constraint, and then evaluates the provided term within the source mapped context. The function is also idempotent, so nothing will happen if `sourceMap` is called again within a source mapped context. Finally, the `mkAnn` function that creates an annotation was changed to also require a source mapped context so it can then grab the call stack and store it in the annotation. This approach ends up enforcing the use of frozen call stacks, while significantly reducing the risk of running into the pitfalls mentioned in the previous paragraph.

Diving deeper however, there are a couple subtleties to this appraoch that are worth exploring in more detail. While the `SourceMapped` constraint does enforce that a call to the function has a valid call stack, the `sourceMap` function itself that sets up the source mapped context cannot statically enforce that the caller has a valid call stack. In essence this thus becomes a partial function that diverges and prints an internal compiler error when it is called without a valid call stack. That way the invariant that source mapped functions always receive a valid call stack is still upheld.

The second caveat is a bit more subtle. Recall that values for implicit parameters are bound for a term based on the location in the program where that term is used, and not where it has been defined. Were the source mapping approach to only use `HasCallStack` without any additional implicit parameters, then frozen call stacks might have still become lost in a helper function defined in a smart constructor's where-clause as a result of the edge case described at the end of Section 2.1, even if every function in the call chain appeared to have the constraint. Luckily because this `SourceMapped` constraint can only be satisfied using the provided `sourceMap` function, this in turn makes it impossible to forget to set it for helper functions defined in a function's where-clause, as doing so results in a compile time error. If `HasCallStack` was used on its own instead then this would have become another potential vector for making mistakes.

However, while a dedicated implicit parameter protects against most potential mistakes, there is still one obvious way where this can still go wrong. If a smart constructor is implemented in terms of another smart constructor, then this approach cannot assert that the call stacks have already been frozen at the first smart constructor. This is because the second smart constructor cannot have the `SourceMapped` constraint, since that would also prevent it from being called from user code. This interacts with the previous caveat in a noteworthy way: if an outer smart constructor calls an inner smart constructor within its body without wrapping the call in `sourceMap`, then the use site of the inner smart constructor inside of the outer smart constructor will be captured in

```haskell
data Ann = Ann { locations :: HashSet CallStack, ... }

-- The constructor is not exported, forcing users of the 'SourceMapped'
-- constraint to use the 'sourceMap' function
data OpaqueType = NotExported
type SourceMapped = (?requiresSourceMapping :: OpaqueType, HasCallStack)

sourceMap :: HasCallStack => (SourceMapped => a) -> a
sourceMap a =
  let ?requiresSourceMapping = NotExported
      ?callStack             = freezeCallStack (popCallStack ?callStack)
   in if isEmptyStack ?callStack then {- internal compiler error -} else a

-- This now captures the call stack from the 'SourceMapped' constraint
mkAnn :: SourceMapped => Ann
mkAnn = Ann { locations = capture ?callStack, .. }
  where
    capture (FreezeCallStack EmptyCallStack) = HashSet.empty
    capture (FreezeCallStack stack@(FreezeCallStack _)) = capture stack
    capture (FreezeCallStack stack) = HashSet.singleton stack
    capture _ = error "Annotations can only be created from frozen call stacks"
```

. . . and in another module. . .

```haskell
cond :: HasCallStack => Exp Bool -> Exp a -> Exp a -> Exp a
cond = sourceMap $ IfThenElse mkAnn

(<=) :: (HasCallStack, Ord a) => Exp a -> Exp a -> Exp Bool
a <= b = sourceMap $ Apply mkAnn (Builtin BuiltinLte) (Pair a b)
```

Listing 13: The annotation system extended with source location information. The `SourceMapped` constraint ensures that a call stack must have been captured before `mkAnn` can be called.

the call stack instead. One possible way to somewhat mitigate this issue would be to follow a convention where every top-level smart constructor `foo` is defined in terms of a source mapped `foo'` function, and to only ever call the source mapped versions of the smart constructor functions in the language's library code.

4.1.1 Pattern synonyms

Capturing call stacks inside of pattern synonym based smart constructors works similarly to capturing call stacks inside of regular functions, but with a couple additional details that need to be taken into account.

Recall from Section 2.3.2.3 that *explicitly bidirectional pattern synonym* essentially just desugar to a pair of functions: one function to construct a term, and another function to destructure a term. When capturing call stacks for these pattern synonyms this also ends up behaving exactly like one would expect. The more practical real-world considerations here are dealing with pattern synonyms that are exclusively used as an alias inside of *simply-bidirectional pattern synonyms*, which are pattern synonyms in the form of `pattern Foo a b = Bar a b`, when using these pattern synonyms as part of a top-level smart constructor, and when targeting older GHC versions. Any of these situations has a subtle but important impact on the capturing of source location information from within a pattern synonym.

The first problem occurs defining a pattern synonym in terms of another pattern synonym. This will be referred to henceforth as nested pattern synonyms. Handling these nested pattern synonyms simply involves keeping track of the nesting depth and then popping that many additional stack frames from the call stack when creating a source mapping context. However, there is no way to programmatically determine how many layers to peel off. As a result, if an outer nested pattern synonym is used as part of yet another pattern synonym, then the captured call stack will end up containing the use site of outer pattern synonym. This is similar to the situation where smart constructor functions call others smart constructors. And even when the nesting depth is correct, all of the patterns still need to explicitly have the `HasCallStack` constraint added to them for this to work. If this is not the case, and the expected nesting depth has not yet been reached, then this will result in a runtime error similar to what would happen when capturing an empty call stack.

Another important consideration is the evaluation order of pattern synonyms. When a top-level pattern matches on its arguments and then uses those arguments in the function's body that has been wrapped in a `sourceMap` call, then the pattern will have already been evaluated outside of the source mapped context. Because the `SourceMapped` constraint cannot protect against this, this ends up capturing the wrong call stack in the terms produced by the pattern. To mitigate this, the function `foo (Bar baz) = sourceMap baz` must be rewritten in an eta-reduced form as `foo = sourceMap $ \(Bar baz) -> baz`. This ensures that the terms produced by the pattern synonym's pattern uses the correct frozen call stack.

The last important consideration is supporting older GHC version. In GHC versions before 9.2[1], GHC does not capture the use site of the pattern synonym in its call stacks. While there is no way to work around this problem, freezing empty call stacks instead

[1] https://gitlab.haskell.org/ghc/ghc/-/issues/19289

will make sure that no incorrect source location information makes its way into the AST and the mkAnn function will still know that source mapping was still taken care of.

For these reasons, the source mapping implementation for pattern synonyms ends up looking like the one in Listing 14. The parts to pay attention to here are the HasCallStack constraints on the pattern synonym and on the matching function. The compiler treats both the expression construction and pattern matching parts of the pattern synonym as regular function calls, and they also interact with call stacks as such. It is also possible to use a lambda inside of the pattern synonym instead of using a separate helper function. Having this in a separate function simply makes the example a easier to read.

```haskell
sourceMapPattern :: HasCallStack => Int -> (SourceMapped => a) -> a
sourceMapPattern nestingDepth a =
#if MIN_VERSION_GLASGOW_HASKELL(9,2,0,0)
  let ?requiresSourceMapping = NotExported
      ?callStack = freezeCallStack (iterate popCallStack ?callStack !! (nestingDepth +
   in if isEmptyStack ?callStack then {- internal compiler error -} else a
#else
  let ?requiresSourceMapping = NotExported
      ?callStack = case ?callStack of
        x@(FreezeCallStack _) -> x
        _                     -> freezeCallStack emptyCallStack
   in a
#endif
```

...and in another module...

```haskell
pattern T2 :: HasCallStack => Exp a -> Exp b -> Exp (a, b)
pattern T2 x y <- (sourceMapPattern 0 unliftT2 -> (x, y))
  where T2 x y = sourceMapPattern 0 $ Pair mkAnn x y

unliftT2 :: SourceMapped => Exp (a, b) -> (Exp a, Exp b)
unliftT2 p = (PrjL mkAnn p, PrjR mkAnn p)
```

Listing 14: The previous section's source mapping mechanism, extended to handle pattern synonym-based smart constructors.

4.1.2 Runtime backtraces

There is one last situation where call stacks are not yet being captured. Recall from Section 2.3.2.2 that the language also implements the standard Num type class to overload Haskell's usual numerical functions. Since the functions in these type classes usually do not have the HasCallStack constraint, there is no way to have GHC generate call stacks for those functions as of writing. A potential workaround for this would be to capture runtime backtraces using the RTS Execution Stacks from Section 2.2.2 instead.

The idea would be to capture backtraces at runtime based on stack frames and any debug symbols that have been embedded in the compiled Haskell program. These could then be filtered to include only user code outside of the language's library, converted to the same format as GHC Call Stacks, and finally stored as a frozen call stack so the exact same source mapping features from the previous sections can be reused without any other changes. In practice, however, this will not work reliably enough for two reasons. First of all, as mentioned earlier, this feature is not likely to be available. It requires GHC to be compiled with the `--enable-dwarf-unwind` configuration option, and it also requires programs to be compiled with the `-g` option. At the moment, no standard GHC distribution has this option enabled, so the user would need to compile their own GHC or track down a version that does enable the feature to be able to use these RTS Execution Stacks. More importantly, the execution stacks are currently not in a format that is easy to work with programmatically without making a lot of assumptions. The execution stacks will include noise output from compiler internals and the RTS itself, making it unclear where the back trace for the current function starts. On top of that, optimizations like tail call optimization and inlining may cause entire functions to be omitted from the backtrace. Still, with some additional work on GHC itself this could be a viable option to capture more source location information in embedded programs.

For the time being, the language will simply freeze empty call stacks in the situations where it is not possible to capture call stacks using GHC Call Stacks. This tells the `mkAnn` function that the call stacks are intentionally left empty.

4.1.3 Explicit context information

As mentioned in the previous section on gathering runtime backtraces, sometimes it is not possible to get a usable call stack from within a smart constructor. In most cases this will not be an issue. For example, with the current language only primitive numerical operations would not collect any call stacks. But these situations may still create holes in profiling output. And in some cases, the data from the gathered call stacks may be too coarse to easily interpret. In those scenarios, it would be nice to be able to manually inject additional context information in the collected call stacks. With the current source location implementation and the `TraverseAnnotations` type class from Listing 12, this can be done using a simple decorator. Listing 15 shows this decorator. As previously explained, these kinds of decorators will not perform any work upfront, and they will instead cause the new data to be merged with the subtree's existing annotations later on in the compilation process. The implementation as shown will inject a valid call stack into the subtree, but with a custom function set name on the bottommost stack frame. As a result, this tackles both the problem of potentially missing source locations as well as the potentially ambiguous call stacks because of generic function names.

```haskell
context :: (HasCallStack, TraverseAnnotations a) => String -> a -> a
context label =
  annotateSubtree $ Ann { locations = insert modifiedStack src, ... }
  where
    Ann { locations, ... } = mkDummyAnn -- Creates a blank annotation

    modifiedStack = case ?callStack of
      ((_, loc) : rest) -> fromCallSiteList ((label, loc) : rest)
      stack             -> stack
```

Listing 15: A decorator that captures the current call stack with a different function label and injects it in all AST in a program's subtree.

4.2 Optimization flags

The main goal of the annotation system was to solve the disconnect between an embedded program and a compiled binary for that program. With that annotation system already in place, it can also be trivially repurposed for other uses by simply extending the annotation data type. This section will look at several ways the annotation system can be used to add user-specifyable local configuration options to the AST that can influence the compiler's optimization pipeline.

Programs are usually written in a way that makes them understandable for humans. But more often than not, there are equivalent ways to formulate a program so that it produces the same result but in a way that's faster to compute. Ideally, compilers would be able to always figure out the optimal way to transform a program into a faster version of that program. But while compilers like the *Souper* [9] *superoptmizer* for LLVM do exist, superoptimization is generally nondeterministic and is thus not suitable for use in regular compilers where compilation speed is also important. Because of that, most compilers optimize programs based on heuristics, and when the try non-deterministic optimizations techniques there is a set limit on how much time the compiler may spend optimizing code. For most use cases, this is good enough. There are, however, situations where the programmer may be able to apply their domain knowledge or their pre-existing knowledge about the program's runtime behavior to make more informed decisions than the compiler would be able to make on its own. Most programming languages expose functionality for this exact purpose through compiler flags, pragmas, or attributes. Since extending the host language with these features is usually not possible, for embedded languages the simplest equivalent way to get the same functionality would be through the same decorator-based approach used in the previous sections. This will involve adding fields for the supported optimization flags and options to the annotation data type, creating decorators that set those options, and finally modifying the compilation pipeline to utilize them.

There are a number of common optimization techniques that may be useful to expose directly to the programmer. The rest of this section will discuss several optimization techniques that may be relevant for compiled deeply embedded languages.

4.2.1 Loop optimizations

A common group of optimization techniques that's widely applicable to most programs revolves around reducing the time it takes to execute a loop-based algorithm. These optimizations usually try to either reduce the number of conditional jumps required or they reorder the way data is accessed in an attempt to make better use of the associative data caches used by modern processors.

Arguably the most common loop optimization is *unrolling* [10]. Here multiple successive iterations are inlined, and the number of total iterations are reduced. If the body of the loop is simple, then this reduces the overhead of an iteration of the loop by only having a single conditional branch for the loop's termination condition after multiple iterations. Furthermore, modern processors have superscalar pipelines that can process multiple instructions at the same time, and not having branches in between those instructions reduces the need to rely on speculative execution. The downsides of this approach are that the loop usually has to be split into a main unrolled part and a tail loop to handle the remaining iterations, and that the amount of generated code can increase dramatically. This may cause some of the performance gains from unrolling the loop to be offset by worse instruction cache utilization.

A second class of loop optimizations involves rearranging loops in an attempt to achieve more optimal access patterns. A common example here is *tiling*, which is sometimes also referred to as *loop blocking* or *strip mining*. The idea is to iterate over the loop in smaller interleaved chunks, which essentially results in having multiple simultaneous read heads. This can allow a larger part of the array to remain in the CPU's cache, which increases cache locality and thus decreases the program's runtime. Other similar optimizations involve interchanging inner and outer loops, skewing loops to reduce intra-iteration dependencies, and more aggressively distributing vectorizable and non-vectorizable parts of a loop into separate loops.

The difficulties that arise with exposing these kinds of optimizations is that depending on the language a plain on/off toggle or an unroll count option may not be sufficiently granular. Imagine a language with special constructs for looping over two dimensional arrays. In that case one may wish to unroll the inner loop, as that loop involves rapid subsequent conditional jumps, but in such a way that keeps the outer loop intact as the amortized overhead from the conditional jumps may be smaller than the costs associated with generating and executing more code. In those cases the loop optimizations may need to be stored in a different way, or it may be necessary to introduce additional constructor-dependent annotation constructs to store dedicated optimization annotations depending on the constructor.

4.2.2 Expression-level optimizations

Another class of optimizations that may be worthwhile to implement are more general optimizations that can be applied to any expression. There are so many options here that it would be impossible to list all of them, so this section will focus on two of them.

One expression-level optimization that is widely applicable to most languages is a group of optimizations that is usually referred to as *unsafe floating point math optimizations*. These are usually exposed through a -ffast-math compiler flag, or through several more granular flags to enable specific unsafe optimizations. With these optimizations enabled, the compiler is allowed to optimize floating point heavy code in ways that would cause it to longer produce the same result or that wouldn't be legal according to the language specifications, but that would execute in significantly less time. While free performance like that is great, there are many situations where these optimizations can lead to subtle or less subtle problems. These problems range from algorithms producing different outputs, algorithms becoming unstable due to uncheckable NaN or infinity values, or subtle logic errors when floating point math is used inside of conditionals. For those reasons, it may be useful to be able to control this behavior within a local scope. With the deferred subtree annotation approach from Section 3.3 it would become possible to enable or disable these optimizations within local scopes of the program with the ability to override this behavior in smaller sub-scopes as needed. This allows the programmer to still have the benefits of faster numerical code when it makes sense while also maintaining correctness elsewhere in the program.

A second optimization is one that is particularly important for languages that are intended for high-performance computing. This optimization interacts with a loop optimization that has not yet been discussed called *loop fusion*. Fusing loops together means that the bodies of two or more loops are merged into a single loop, and as such the data only needs to be iterated over once, and the intermediate results don't need to be written to memory. For simple cases, like multiple sequential loops, compilers should be able to figure this out on their own. But when the results produced by one loop are reused in two separate, completely independent loops, then fusing the three loops may not always be possible. In that case the programmer may choose to get rid of the first loop and to recompute the values that would be produced by that loop directly inside the two other loops. But with deeply embedded languages this may not always work this way.

As discussed earlier, deeply embedded languages that use a *higher-order abstract syntax* do not use explicit variable bindings in their abstract syntax. Terms will always appear as if they were inlined when traversing the AST, even if they are referring to the same term in the host program. In those situations, the compiler may try to reintroduce *sharing* when compiling the program. This can be done by transforming the abstract syntax into another form where shared common subexpressions are bound to scoped variables. But as mentioned in the previous paragraph, as a result it may also prevent loop fusion from happening.

By adding a decorator to the language that prevents a term from being shared, or alternatively, to force it to be inlined, it can becomes easier for the compiler to fuse these intra-dependent loops. Forcing computations to be inlined rather than written to memory may speed up a program if the computation is trivial, as reading and writing data from and to main memory can add significant latency penalties.

4.2.3 High-level code generation instructions

A third group of optimizations are those that influence the entire compilation process. These settings can also be specified directly when invoking the compiler, but storing these settings in AST annotations can have a number of benefits that might make it worthwhile. First of all, moving these settings to the program instead of having to pass them to a compilation function can make specifying these options feel more natural. Instead of passing arguments to the compiler, the program can be decorated with the required optimization flags. This approach also makes it possible to have multiple versions of a program with different optimization options simply by changing which optimization flags the program is decorated with. Secondly and more importantly, the compiler may not compile the program as a whole. The compiler may end up dividing the program into sections that are each treated individually. Common examples of such sections are translation units or separate executable kernels. Storing the optimization flags as part of AST annotations makes it possible to have different compiler settings for each of those subprograms, and it turns runtime configuration in the host program into compile time configuration for the embedded program.

An example of a setting that would be useful to be able to specify on a (sub)program basis is the optimization level and the presence of debug symbols. Similarly, being able to specify which instruction sets the compiler may use when generating code can be useful if a compiled version of the deeply embedded program is embedded in the host program. That would allow programs to either be tuned for the specific machine they are running on, or they may target a more general set of instructions that allows the distributed program to run on a wider range of target machines. Lastly, domain specific embedded languages may benefit from tuning options specific to that domain. For instance, a language that compiles to GPU shaders may benefit from a limit on the number of registers that are used by each thread. Limiting this number to a divisor of the total number of available registers available to the compute unit may increase device occupancy at the cost of per-instance throughput.

5 Case study: Accelerate

As usual, the devil is in the details. Up until now this thesis has described a fairly generic framework for annotating a deeply embedded languages along with a source mapping annotation using GHC's features and an enumeration of other uses for this annotation system. This section will take a look at how these ideas end up working out in the real world. To evaluate this, the annotation system has been implemented in the Accelerate data-parallel array language [11]. Accelerate is a deeply embedded language written in Haskell that is interesting for a number of reasons. First of all, it allows writing highly vectorized code for working with multidimensional data arrays that can end up being magnitudes faster than the equivalent program written using plain Haskell. This is done by compiling the AST of the embedded program either to machine code intended for a CPU or to shaders intended for a GPU, all without having to change the program itself. Accelerate's language contains several second-order combinators that behave similarly to the typical list and array combinators one might use in a regular Haskell program such as maps and folds, as well as more specific operations for working with multidimensional arrays like transpositions and stencil operations. These operations, dubbed *collective operations*, are turned into a more generalized form by Accelerate's compiler that can be compiled into a kernel and then executed by Accelerate's runtime system.

Accelerate's internal architecture also adds a couple additional nuances to the implementation of the annotation system as it has been described so far. For instance, Accelerate uses multiple different AST types. Several compilation passes then transform the program written by the user through all of these different representations. This starts with a higher-order abstract syntax (*HOAS*) that gets transformed into a first-order abstract syntax (*FOAS*) when the program is compiled, which after a some optimization passes transformations gets transformed into another backend-specific representation by the compiler's backend [12]. During these transformations AST nodes are merged, new AST nodes are created, and semantics change. The rest of this section will go through this compilation pipeline to examine what needs to be done for annotations stored in the original HOAS to be propagated through to the final stages of compilation. After that, the implementation will be evaluated both in terms of performance and in the usability improvements it brings.

Links to the modified versions of Accelerate and its compiler along with instructions on how to migrate Accelerate libraries to the annotated version of Accelerate and an example of a migrated library can be found in Appendix A.

5.1 Sharing recovery

On the surface, writing a program in Accelerate works similarly to how one would use the example language from Section 2.3. In both cases the programmer ends up constructing an embedded program by chaining together smart constructors. In Accelerate's case, this may end up looking similar to a program that operates on lists written using Haskell's standard library. But this is also immediately where the example language and Accelerate diverge. Where the example language's abstract syntax was first-order and used explicit variable bindings, Accelerate's abstract syntax is higher-order and does not require the programmer to manually introduce let-bindings for variables. As discussed earlier, the downside to this approach is that naively evaluating a program written in such a language may involve recomputing the same terms many times, as they appear in the AST as if they were inlined. To avoid the performance cost associated with unnecessarily recomputing expressions like this, Accelerate uses a technique called *sharing recovery* in an attempt to recover the sharing of the terms in the original host program. In the process, Accelerate converts the HOAS representation of the program into a first-order form with explicit let-bindings.

This is where one of the complications hinted at in Section 3.3 comes in. There are several different ways to perform this sharing recovery. The naive method, which simply finds all common subexpressions throughout the program, has an exponential time complexity and is thus not suitable for most programs. Because of that, Accelerate and most similar deeply embedded languages written in Haskell use an approach proposed by Gill [13] that reuses the *stable names* Haskell's runtime system assigns to variables. These stable names identify a unique object that may have been used in multiple places after it was bound to an identifier. Two terms with the same stable name are aliases, and can be thought of as pointers to the same underlying object. Finding duplicate occurrences in the AST by keeping track of each AST node's stable name is thus much cheaper than doing direct comparisons. However, problems start arising when those AST nodes are modified in any way. A transformation on the HOAS that would go over every AST node without making any changes to the actual nodes themselves will end up producing the exact same AST. But in the process, all stable name information is also lost, since applying the identify function to a single term twice will end up producing two equivalent results but with different stable names. For this reason, annotating AST subtrees in Accelerate has been implemented by storing the modification as part of the tree, and the deferred modification is then applied at the very end of the sharing recovery process.

Actually applying these subtree annotations is relatively straightforward. The implementation involves a new implicit parameter that keeps track of the changes that need to be made to the current subtree's annotations and a function that merges those changes with the annotations already stored in the subtree's AST nodes. Then after recovering all sharing information, the implicit parameter keeps track of the current annotation state while recursively descending the AST, and the merging function merges the current subtree annotation state with the subtree annotations that were already stored in the subtree. This approach is simple, and the implicit parameters make it possible to implement all of this without much boilerplate. As mentioned earlier, this approach also makes it possible to disable and subsequently re-enable optimization flags such as the fast-math control flag for part of a subtree.

Aside from propagating annotations to the subtrees, another part of the annotation system that was implemented as part of the sharing recovery is forcing inlining of terms. If the option to always inline a term is set on an AST node, then any occurrence of that node is treated as if the node's stable name has not been seen before. Because sharing recovery is a transformation from a HOAS to a FOAS, this works oppositely of how inlining would work in a traditional programming language. Instead of forcing a term to be inlined, setting this flag means that any sharing of the term will not be recovered.

And because a HOAS semantically already contains inlined copies of every term, this means that the inlined version is preserved and carries over to the FOAS. Otherwise it would have been bound by a let-binding.

5.2 Code transformations

After sharing recovery, the new AST goes through several more transformations. The first set of transformations is a simplification step that mainly involves rewriting expressions into equivalent expressions that are cheaper to evaluate, pruning unnecessary branches, and constant folding. As a consequence, multiple AST nodes may be merged into one, and new AST nodes may appear as the result of a compiler optimization. After this transformation some of the program's AST nodes may have been constructed from multiple sites in the original source code that may also not be adjacent to each other. This is one of the reasons why the source locations in the annotations are stored as a set of call stacks rather than keeping track of a single call stack. When AST nodes are merged, the annotations are treated as a join-semilattice. Less formally this means that the sets of call stacks are merged, and the node will contain all optimization flags that were present in the original AST nodes.

The more interesting transformation pass is *array fusion* [14]. Here multiple subsequent collective array operations are merged into a single operation. This reduces the number of times an array has to be iterated over, and more importantly it also avoids having to write the intermediate arrays to memory. Especially when targeting GPU hardware, writing the results of a computation to memory and immediately reading it again in the next operation can result in a lot of performance overhead. The interesting part here with respect to the annotation system is deciding what should be done with the annotations when multiple operations are fused together. When the optimization flags are the same for every fused operation, then a simple join of the annotations stored in the original operations will suffice. But when the sets of optimization flags differ, then there are a couple different options. One option would be to throw a fatal compiler error and to refuse to compile the program as is. This would prevent situations where for instance loops are being unrolled that the user doesn't want to be unrolled, but it can also lead to a bad user experience since this happens late in the compilation pipeline and the user may not even be aware of the whole array fusion process. A second option would be to refuse to fuse array operations with incompatible optimization flags. This ends up costing performance, and even if a warning is printed it may still lead to confusion as the program's execution time may be several factors higher if fusion does not take place. A third option would be to only honor the optimization flags of the first operation that the other array operations get fused into, while printing a compiler warning to indicate that some optimization flags are being ignored. Finally, a fourth option would be to always join the optimization flags as they merged into the fused array operation. The current implementation takes this last approach because Accelerate currently does not have a robust compiler diagnostics system. As future research however, the third approach may be worthwhile to experiment with. Compiler diagnostics could be a useful addition to a language like Accelerate, and the annotation system could also play a role there by allowing warnings to be configured or silence within a local scope.

5.3 Code generation

The final part of the compilation pipeline revolves around doing something useful with the optimized AST. As mentioned before, Accelerate is backend agnostic, meaning that it is possible to implement any number of compiler backends without needing modifications to Accelerate itself. Currently, there is an interpreter backend and there are two LLVM-based backends that compile to native code. Since the interpreter is mostly useful as a reference implementation, the focus here will be on the two LLVM backends. One of the LLVM backends compiles to native machine code, while the other backend targets NVIDIA GPUs through CUDA's PTX interface. PTX is a low level assembly-style instruction set architecture that allows writing compute kernels for CUDA capable GPUs. Both backends also share a larger common code base that takes care of most of the non-target specific abstractions.

The first matter to discuss is getting the annotations stored in the optimized AST into the backend's code generation. The LLVM backends can conceptually be divided into three parts. Collective array operations are compiled into predefined *kernels*, while the expressions used within those array operations get turned into equivalent LLVM IR expressions that can be embedded into those kernels. Lastly, the backend may also evaluate small parts of the AST directly.

LLVM IR is a relatively simple and easy to reason about low level language that can be compiled to machine code targetting a variety of architectures. Within Accelerate's LLVM backends, the LLVM IR is represented by another internal strongly typed deeply embedded language which it can later turn into actual IR. Like the Accelerate language itself, this IR language also comes with a number of smart constructors and functions to make it possible to write low level programs with imperative semantics. There are also some situations where Accelerate's LLVM backend does not generate any code from the optimized AST, evaluating expressions directly instead. This is done for constructs like array-level conditionals, where the result of an expression determines which other collective operations are run next. Luckily, while they are conceptually very different, the implementation of these three uses for the optimized AST follow the same general structure. Additionally, at this point the annotations stored within the AST remain constant. This makes propagating the annotations through to every part of the backend compiler another good candidate for implicit parameters. This way the only change needed to the code is to add the implicit parameter to each function that either directly or indirectly needs to use the annotations. The alternative would be to explicitly pass the annotations around as an explicit function parameter, or to use the `Reader` monad as discussed earlier [2]. Both of those alternatives would end up adding boilerplate to every function that needs to pass through the annotations.

Currently, the annotations are used in two places in the LLVM backends: to generate debugging symbols and profiling information, and to influence both Accelerate's and LLVM's optimization behavior. The upcoming sections will discuss both of these uses separately.

5.4 Source locations

As discussed earlier, the main use cases for the source location information that has now been embedded in the ASTs are improved compiler diagnostics, better profiling support, and source-level debugging. Using the new implicit parameter these source locations can be accessed anywhere in the backend where they are needed. But before the source locations can be used, there is still one more step that needs to happen.

Source locations are stored in the annotations as sets of call stacks. This can be an empty set for AST nodes without any source information, a set containing a single call stack obtained through the source mapping smart constructors, or a set containing multiple potentially disjoint call stacks for nodes that have been modified or merged by Accelerate's optimizer. In the last case, adjacent call stacks must be merged into one larger region before they can be used to provide source information to external tooling. Furthermore, most existing tools can only handle a single call stack. If there are still multiple disjoint call stacks after merging them, then a heuristic can be used to select the call stack that is most likely to be relevant based on the number of call stacks that were merged into that call stack.

The algorithm used for merging call stacks is straightforward. First, all call stacks are sorted based on the file name, starting line number, and starting column number of the bottommost element in the call stack. Next, the algorithm iterates over the list of call stacks, and adjacent call stacks are merged into a single stack. Determining adjacency works based on a simple heuristic. When region A and region B are sorted after each other and region B is either contained within region A or if it starts in the line after region A's end, then they are considered adjacent. Finally, the actual merging works by prepending the function name from region B to region A's function name separated by a comma, while extending region A's source location to cover both regions. This can all be done in $O(n \log n)$ time where n is the number of call stacks in the set.

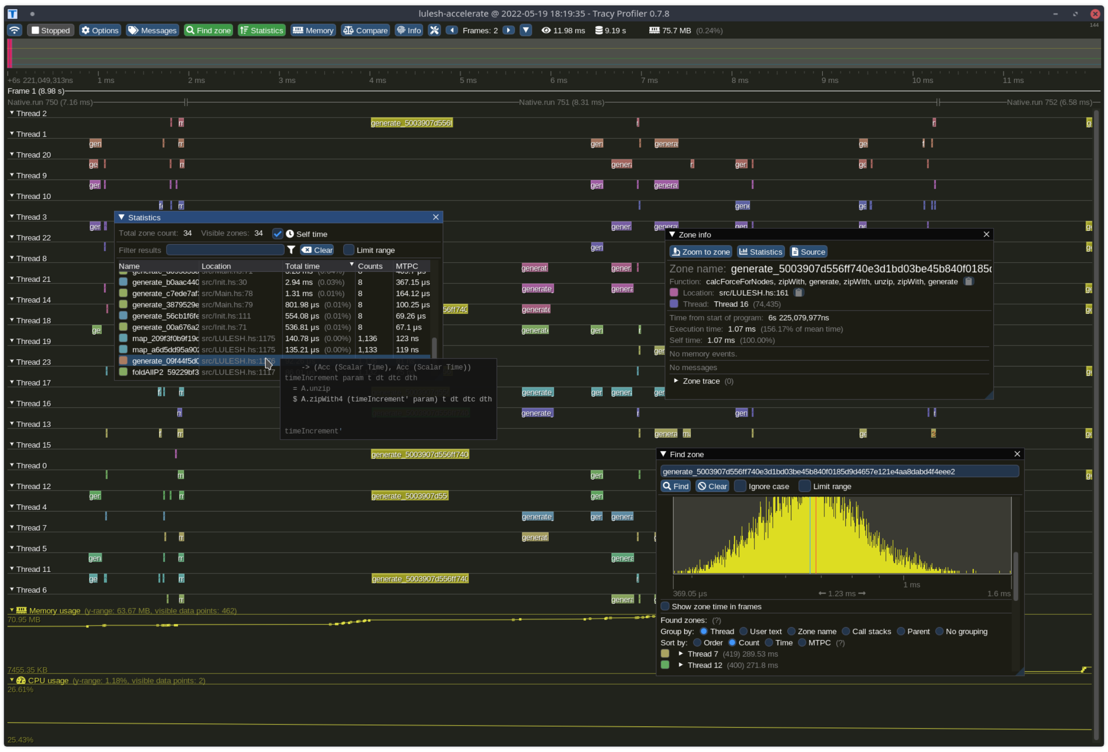
5.4.1 Frame profiling

One of Accelerate's current biggest obstacles is that it can be difficult to get good insight into why a program is running slower than expected. From the perspective of the user, an embedded program turns into a black as soon as the backend's run function compiles and executes the program. During the program's execution, profilers are a good way to get some level of insight into where most of the time is spent. NVIDIA's Nsight Systems [15] profiler can be used with any CUDA application and can thus also be used with Accelerate's GPU backend. Similarly, Accelerate's CPU backend contains support for the *Tracy* [16] frame profiler. Both of these tools visualize which kernels are executed when, for how long, and on which thread during a single run of the embedded program. While tools like this were originally designed for profiling frame timings for video games, they are also well suited for profiling any other repetitive process. However, while they do show the general high level program flow in terms of what kernels run where and when, without additional instrumentation they cannot offer any insight into what those kernels actually are and where they are coming from. Luckily, thanks to the annotation system this problem can now be solved.

As the annotation data for the expression that is currently being compiled is already accessible throughout the entire LLVM backend through the implicit parameter introduced in the last section, adding additional instrumentation to the existing Tracy support becomes trivial. On its own, Tracy does not do much. It requires the program to connect to a profiling server, and then the program needs to inform Tracy about the program's regions of interest by calling a Tracy API function. Before Accelerate used to only provide generic names and context information to these functions based on nondescript hashes. But with the source mapping system in place it is now possible to use the actual call stacks from the kernel function calls.

Figure 1 shows the result of running the LULESH-Accelerate [17] application under Tracy with the annotation system in place. This is a port of the LULESH [18] mini-application, which is a highly simplified hydrodynamics simulation. Tracy's output shows several `Native.run` calls along the horizontal axis. These correspond to a single call of the iterative algorithm. The vertical lanes show which computation kernels are being executed by each of the CPU's 24 threads at a given time point, as well as the application's overall CPU usage and memory consumption.

Previously, running this simulation under Tracy would also show these kernels being executed, but without any context information beyond the automatically generated kernel names. These are based on the kernel's main entry point function and a hash of the optimized AST for that kernel. While this may not pose any problems for simple programs, with more complex programs like this LULESH implementation it can be difficult to tell which kernel belongs to which part of the application as there is no direct link between them. However, with the source mapping system in place the output now contains a lot more context information that makes tracing the execution through the original program much simpler. Notice how every kernel now shows file names and line numbers corresponding to the location of the kernel's definition in the original source code. Hovering over these locations now also opens a dialog showing an excerpt from that source code. Also pay attention to the '`Function:`' field in the zone info dialog. That field now shows the name stored in the top most stack frame from the *merged call stacks* discussed earlier. This is useful for two reasons: first, it gives insight into how many array operations and more specifically which array operations were fused into the kernel. The more function names listed here, the better a job Accelerate's array fusion algorithm did optimizing out unnecessary array operations. Secondly, this also



Statistics

Total zone count: 34 Visible zones: 34 ✓ Self time

Filter results | Clear | Limit range

Name	Location	Total time	Counts	MTPC
generate_b0aac440	src/Init.hs:30	2.94 ms (0.03%)	8	367.15 µs
generate_c7ede7af	src/Main.hs:78	1.31 ms (0.01%)	8	164.12 µs
generate_3879529e	src/Main.hs:79	801.98 µs (0.01%)	8	100.25 µs
generate_56cb1f6fe	src/Init.hs:111	554.08 µs (0.01%)	8	69.26 µs
generate_00a676a2	src/Init.hs:71	536.81 µs (0.01%)	8	67.1 µs
map_209f3f0b9f19c	src/LULESH.hs:1175	140.78 µs (0.00%)	1,136	123 ns
map_a6d5dd95a902	src/LULESH.hs:1175	135.21 µs (0.00%)	1,133	119 ns
generate_09f44f5d0	src/LULESH.hs:1176			
foldAllP2 59229bf3	src/LULESH.hs:1117			

```
    -> (Acc (Scalar Time), Acc (Scalar Time))
timeIncrement param t dt dtc dth
  = A.unzip
  $ A.zipWith4 (timeIncrement' param) t dt dtc dth


timeIncrement'
```

Zone info

Zoom to zone | Statistics | Source

Zone name: generate_5003907d556ff740e3d1bd03be45b840f0185

Function: calcForceForNodes, zipWith, generate, zipWith, unzip, zipWith, generate

Location: src/LULESH.hs:161

Thread: Thread 16 (74,435)

Time from start of program: 6s 225,079,977ns

Execution time: 1.07 ms (156.17% of mean time)

Self time: 1.07 ms (100.00%)

No memory events.

No messages

Zone trace (0)

Find zone

generate_5003907d556ff740e3d1bd03be45b840f0185d9d4657e121e4aa8dabd4f4eee2

Find | Clear | Ignore case | Limit range

369.05 µs ← 1.23 ms → 1.6 ms

Show zone time in frames

Found zones: (?)

Group by: [✓ Thread] User text Zone name Call stacks Parent No grouping

Sort by: Order [✓ Count] Time MTPC (?)

Thread 7 (419) 289.53 ms

Thread 12 (400) 271.8 ms

Figure 1: The LULESH-Accelerate [17] application running under the Tracy frame profiler. This screenshot demonstrates the additional context information made possible through the new annotation system.

shows the explicit context information from Section 4.1.3 that was added to this part of the AST using a decorator function. Scattering information like this throughout the program makes it simpler to get an at a glance overview of what happens where in the program, and it also adds source information to functions that might otherwise not have any due to limitations of the source mapping system.

5.4.2 Debug symbols

While specialized profiling tools like Tracy can be invaluable for getting insight into the execution of a specific algorithm, their main weakness is that they require the program to be manually instrumented with calls into an accompanying profiling library. The more widespread approach to profiling and debugging applications thus revolves around *debug symbols* that are either embedded in the application's binary or stored alongside it in an accompanying metadata file. These debug symbols provide structural information about a region of the program's binary that can be used to map back to parts of the original program. This can range from descriptions of high level constructs like files, namespaces, and subprograms, all the way down to function parameter descriptions, jump labels, and individual expressions. The LLVM compiler's metadata system is able to generate these debug symbols for a program by annotating parts of the LLVM IR with special metadata tags. LLVM can then generate debug information in a variety of supported formats, such as the commonly used DWARF format [19]. Particularly interesting for Accelerate is the ability to output debug information when targeting CUDA through PTX. Sadly, as of writing support for outputting debug information in this format is incomplete. But if Accelerate already has full debug metadata support by the time LLVM gets better support for this feature then Accelerate should be able to benefit from better GPU debugging and profiling support using NVIDIA's Nsight Systems without requiring any changes to Accelerate itself. Because even though Nsight Systems does act like a frame profiler as mentioned in the last section, it doesn't use explicit instrumentation like Tracy does and instead relies on regular debug symbols.

The implementation of these debug symbols in Accelerate extends the backend with several methods to annotate the LLVM IR AST with the aforementioned debug information metadata. This is done by adding additional smart constructors to the backend's internal deeply embedded IR language. These smart constructors can generate local as well as global metadata definitions within the program. Whenever Accelerate's LLVM CPU backend now creates a kernel function, the backend also generates the associated debug metadata for that function. And when LLVM compiles the IR, this metadata gets turned into DWARF debug symbols. In theory this should be sufficient for sampling profilers like *perf* [20] or debuggers like *gdb* [21] to get information about the current kernel that's being executed at a given point in time. Sadly, the details on how to implement DWARF debug symbol support in LLVM are sparsely documented, and currently this feature does not always work correctly. Still, with more work it should be possible to get debugging and profiling working with these DWARF-based tools on the kernel-level. And since the expressions contained within a kernel also have source mapping information attached to them, it would also be possible to extend the implementation to be able to support full expression-level stepping debugging.

5.5 Optimization flags

While the source mapping was the main focus for this thesis, the secondary goal was to find other uses for the annotation system. Section 4.2 explored the idea of adding decorators to a deeply embedded language that would change the compiler's optimization behavior either for a single expression, for a subtree of the program, or for an entire compilation unit by storing that information inside the annotations. Most of the listed optimizations have been implemented in Accelerate and its LLVM backends. As most of these optimizations directly influence the way code is being generated, most of the implementation work was done within Accelerate's LLVM backends. The exception here is the forcibly inlining of variables, as explained in more detail a couple sections ago. Before discussing the other optimizations, there is one more important detail to the forcibly inlining of terms that has not yet been discussed.

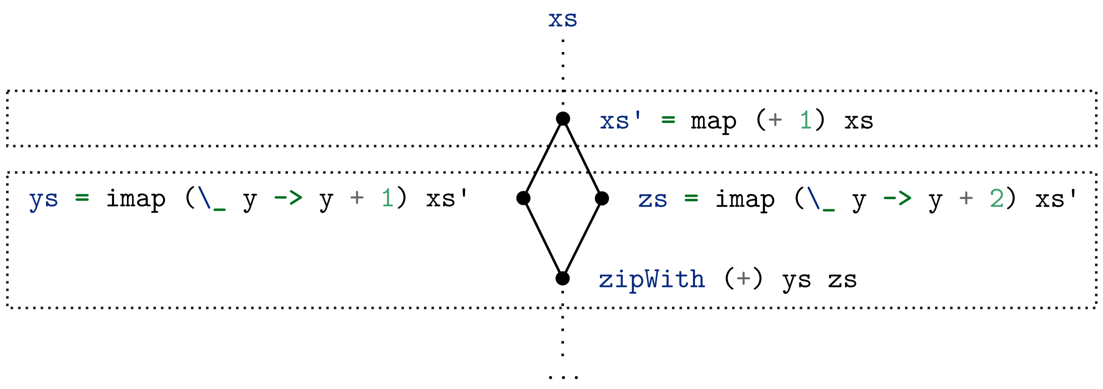


Figure 2: An array program written in the Accelerate data-parallel array language. The dotted boxes indicate a single pass over the data after fusion has taken place. In this example, the result of `map (+ 1) xs` being shared prevents this group of operations from fusing into a single kernel.

5.5.1 Inlining

As alluded to when first discussing the optimization, other than getting rid of potentially expensive memory operations, forcing terms to be recomputed can have another advantage within the context of Accelerate. One limitation of Accelerate's current array fusion system is that it can only fuse an array operation that produces an array into at most one other array operation. Because of that, array programs that fuse well tend to have a funnel-like structure where produced arrays are either used to produce one array, or the array is combined with other arrays to produce a single array. That in turn means that when an output array is used as the input for two other array computations, then array fusion is not possible with Accelerate's current fusion model. An example of a situation where this is a problem is shown in Figure 2. Here there is an input array called `xs` that has each of elements incremented by one. That incremented version is then used in two other mapping operations, which are then summed together to produce a result. With Accelerate's current array fusion model the bottom three array operations that are surrounded by the dotted box can be fused together, while the simple `map (+1)` operation at the top of the figure needs to be fully computed and written to memory before the fused kernel can be executed. This is wasteful, as the time it takes to read and write this memory as well as the time it takes to synchronize all participating threads within Accelerate's scheduler greatly exceeds the time it would take to perform a single addition. By simply changing that expression to `forceInline (map (+1) xs)`, the addition is now instead calculated as part of the two `imap` calls, which in turn allows the entire calculation from the figure to be fused into a single operation.

5.5.2 Loop unrolling

The second optimization that has been implemented is the unrolling of loops. As Accelerate's LLVM backends use an internal deeply embedded language for constructing LLVM IR, these implementation is localized to two primitive loop operations. The first operation iterates over an array with optional indices and a certain step size, while the second operation accumulates a value as part of the loop. Most other loop-based constructs supported by the backend are implemented in terms of one of these two functions. And because the annotation data for the current expression that is being converted to LLVM IR code was already made accessible through an implicit parameter, implementing this did not require any changes outside of these two functions. As a result, it would be possible to implement similar optimization flags for performing other algorithmic optimizations in the future with minimal unrelated code changes. While the unrolling works as expected, one limitation of the current loop unrolling implementation is that it will apply to all loops produced by an AST node. Annotating a simple mapping function with `unrollIters 16` will cause the loop to be unrolled into chunks of sixteen elements as one might expect, but doing the same thing with a two dimensional stencil operation will cause both the inner and the outer loop to be unrolled. Coming up with a flexible method to have more control over the unrolling behavior of nested and compound loop operations would be an interesting subject for future research.

It is also worth noting that LLVM by itself already supports many common loop transformations [22]. By simply adding metadata annotations to the LLVM IR source code, LLVM's compiler should be able to automatically perform those transformations to any program. Sadly, as of writing the loop unrolling annotations don't do anything. And Clang, one of the most prominent compilers for the C family of programming languages and the main consumer of LLVM, also doesn't use this metadata. Still, when this does start to work in a future revision of LLVM, then using this metadata to do the optimizations instead of implementing them by hand before generating the IR would both open up new optimization possibilities while simultaneously removing complexity from Accelerate's LLVM backend.

5.5.3 Other optimizations

The last two optimization flags that were implemented are support for controlling the `-ffast-math` behavior on a program subtree level, and the ability to limit the number of registers that can be used by a thread in a CUDA kernel. In the LLVM IR language, which `-ffast-math` semantics an expression should use is set directly as an optional argument to the expression itself. This makes it possible to allow these unsafe optimizations that don't conform to the specifcation on specific operations, like a single addition and a single multiplication, without affecting anything else. And because every primitive scalar level expression in the annotated version of Accelerate's AST contains an annotation, it is possible to map this idea one-to-one to Accelerate programs. To recap, the idea is that the programmer can disable the on-by-default `-ffast-math` behavior for an expression or an entire kernel by decorating it with a `withoutFastMath` combinator. It is also possible to re-activate these optimizations for a smaller part of that expression or kernel by using a similar `withFastMath` decorator. During sharing recovery these optimization flags are then written to the annotations stored in all child nodes of the annotated subtree. As a result, when generating LLVM IR code based on an expression, whether or not that expression should use `-ffast-math` semantics can be queried by simply inspecting the current expression's annotation that gets passed around using an implicit parameter.

The optimization flag for controlling the number of registers that may be used by a CUDA kernel works by simply passing the value stored in the kernel's annotations to the PTX assembler that turns PTX code into the GPU's equivalent of machine code. While this is not particularly exciting, this way of using annotations does open the door to other similar methods of programmatically controlling compiler the compiler's overall behavior. One optimization flag that has not yet been implemented but that would be very useful within the context of Accelerate is the ability to control the architecture that kernels are compiled for. Currently the CPU backend uses all of the instruction set features available on the CPU that's compiled the program, and the CUDA backend uses the latest architecture supported by the connected GPU. Accelerate has an interesting feature called `runQ` that compiles the embedded program at Haskell compile time and then embeds the binary for the compiled embedded program into the compiled Haskell binary. When executing the host program, the embedded program can then immediately run without needing to be compiled first. With a couple of tweaks it would be possible to use the annotation system to create kernels that target specific and even multiple architectures. Embedding those into a compiled host program would make it possible to run the embedded programs on a wider range of computers without requiring the full LLVM toolchain to be installed on those computers.

5.6 Results

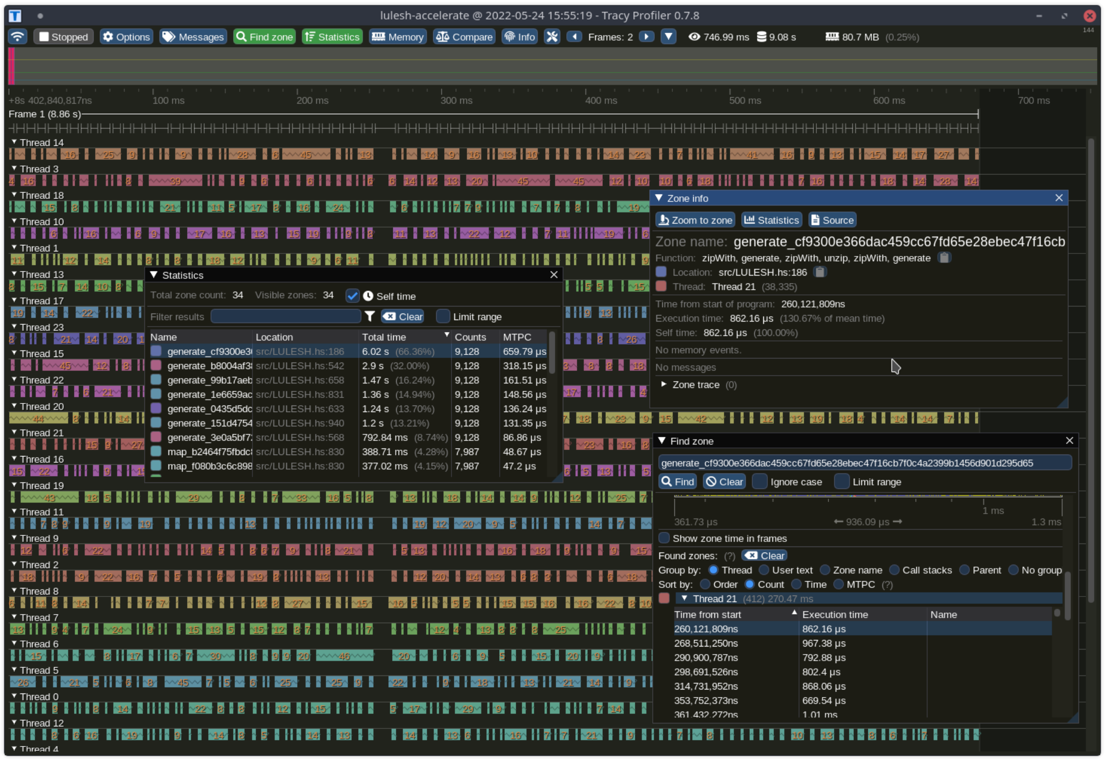
Previous sections already looked at what the addition of source information can do for Accelerate programs in terms of improving the general development experience. Now it is time to look at some of the more practical uses of having access to the optimization flags from the previous section. As part of the optimization process, Tracy is used to find the hotspots and to analyze how well Accelerate is able to fuse array operations.

5.6.1 Compensated summation

The first example we will take a look at was the direct motivation for implementing expression level `-ffast-math` configuration in Accelerate. Computers usually store decimal numbers in a floating point format. With a couple exceptions for special values and numbers that are too small to be represented this way, this comes down to storing a number in the form of $s \cdot m \cdot 2^e$, where s indicates the sign with either a -1 or a 1 value, m is a number in the half-open range $[1, 2)$, and e is a negative or positive exponent. This allows a wide ranger for numbers to be represented even with a limited number of bits assigned to m and e. But because m spans the $[1, 2)$ range with only a finite number of bits, not every value between one and two can be represented. The difference between the expected result of a floating point addition or multiplication and the value stored by the computer is called the *roundoff error*. This is the result of the actual value being quantized to the closest matching value that can be represented using s, m, and e. While in most cases these roundoff errors tend to be so small that they can be safely ignored without causing any harm, there are numerous situations where the accumulated roundoff errors from repeated floating point operations add up to an amount that is actually significant. As an example, the numbers 0.1, 100.0 and 1000.0 can all be perfectly represented by both single and double precision IEE-754 floating point numbers, which is the floating point format used by all current consumer CPUs. However, starting with a value of 0.0 and then adding 0.1 a thousand times to that in a loop results in a total value of 99.99905 for single and 99.9999999999986 for double precision rather than the expected value of 100.0. A similar situation occurs when summing very large or very small floating point numbers. The four single-precision floating point numbers 1.0, 1.0, 2.0^{100} and -2.0^{100} should sum to 2.0. But depending on the order of the numbers, adding them up results in a value of either 0.0 or 1.0, even though all of these numbers as well as the expected result can be perfectly represented by a single precision floating point number. Because of these reasons, people have come up with algorithms that try to compensate for these rounding errors. This class of algorithms is aptfully named *compensated summation* [23].

The `accelerate-blas` package implements several of these compensated summation algorithms for Accelerate [2]. These algorithms keep track of the accumulated error from previous additions and then compensate for that error in the next addition. While the exact details on how this works are not important, it is relevant to understand that for this to work, these algorithms need to perform additions and subtractions in a specific order to be able to compute the error. However, as previously noted, for performance reasons Accelerate always enables all of LLVM's unsafe floating point math optimizations. And as a result, the compiler ends up reordering the aforementioned additions and subtractions, breaking the compensated summation algorithms. As a workaround, the `accelerate-blas` package currently directly imports LLVM's addition and subtraction intrinsics through the foreign function interface and uses those in place of Accelerate's regular addition and subtraction operators. Doing so requires deep knowledge of Accelerate's internals and the implementation also needs to be modified to use these new FFI function calls. With the new annotation based optimization flags however, the original unmodified version of the algorithm works as expected when wrapping it with the `withoutFastMath` decorator.

[2] `https://web.archive.org/web/20220523154459/https://github.com/tmcdonell/accelerate-blas/blob/master/src/Data/Array/Accelerate/Numeric/Sum.hs`


lulesh-accelerate @ 2022-05-24 15:55:19 - Tracy Profiler 0.7.8

Frames: 2 · 746.99 ms · 9.08 s · 80.7 MB (0.25%)

Frame 1 (8.86 s)

Statistics

Total zone count: 34 Visible zones: 34 Self time

Name	Location	Total time	Counts	MTPC
generate_cf9300e3	src/LULESH.hs:186	6.02 s (66.36%)	9,128	659.79 μs
generate_b8004af3	src/LULESH.hs:542	2.9 s (32.00%)	9,128	318.15 μs
generate_99b17aeb	src/LULESH.hs:658	1.47 s (16.24%)	9,128	161.51 μs
generate_1e6659ac	src/LULESH.hs:831	1.36 s (14.94%)	9,128	148.56 μs
generate_0435d5dc	src/LULESH.hs:633	1.24 s (13.70%)	9,128	136.24 μs
generate_151d4754	src/LULESH.hs:940	1.2 s (13.21%)	9,128	131.35 μs
generate_3e0a5bf7	src/LULESH.hs:568	792.84 ms (8.74%)	9,128	86.86 μs
map_b2464f75fbdc	src/LULESH.hs:830	388.71 ms (4.28%)	7,987	48.67 μs
map_f080b3c6c898	src/LULESH.hs:830	377.02 ms (4.15%)	7,987	47.2 μs

Zone info

Zone name: generate_cf9300e366dac459cc67fd65e28ebec47f16cb
Function: zipWith, generate, zipWith, unzip, zipWith, generate
Location: src/LULESH.hs:186
Thread: Thread 21 (38,335)
Time from start of program: 260,121,809ns
Execution time: 862.16 μs (130.67% of mean time)
Self time: 862.16 μs (100.00%)
No memory events.
No messages
Zone trace (0)

Find zone

generate_cf9300e366dac459cc67fd65e28ebec47f16cb7f0c4a2399b1456d901d295d65

361.73 μs · ← 936.09 μs → · 1.3 ms

Group by: [✓ Thread] User text Zone name Call stacks Parent No group
Sort by: Order [✓ Count] Time MTPC

Thread 21 (412) 270.47 ms

Time from start	Execution time	Name
260,121,809ns	862.16 μs	
268,511,250ns	967.38 μs	
290,900,787ns	792.88 μs	
298,691,526ns	802.4 μs	
314,731,952ns	868.06 μs	
353,752,373ns	669.54 μs	
361,432,272ns	1.01 ms	

Figure 3: Tracy showing hotspots in the LULESH-Accelerate [17] application. The highlighted `generate` kernel been fused with five other operations.

5.6.2 LULESH

Now it is time to take a look at how trivial usage of the loop unrolling decorators affects the performance of the aforementioned Accelerate implementation of the LULESH [18] hydrodynamics simulation. More specifically, this section goes through the basic workflow of optimizing a program based on profiling data. For brevity's sake the rest of this section will link to specific lines in the LULESH implementation rather than citing them directly.

The first step when optimizing any application is to find where the hotspots are. Luckily, thanks to the new annotation system, Tracy was able to tell us exactly where most of the time was being spent. A screenshot of Tracy's output is shown in Figure 3. Over 65% of an average run's time is spent in the function that calculates forces for each element in the simulation[3]. This function shows up in Tracy as `zipWith, generate, zipWith, unzip, zipWith, generate`. That indicates to us that the compiler has been able to fuse six different collective array operations into a single operation. Since this is a hot

[3] `https://github.com/robbert-vdh/lulesh-accelerate/blob/2618ebf7881d00a2052b9c4aee78c47c985484bb/src/LULESH.hs#L165-L215`

loop, it may seem like a good candidate for manual unrolling.

Before making any changes, the unmodified program takes an average of 8.65 seconds to run on an AMD Ryzen 9 5900x 12 core processor with SMT with a standard deviation of 0.12 seconds over the coarse of 30 runs. We'll prepend `unrollIters 8` to one of the Accelerate function calls in the function linked above and... the runtime increased to an average wall runtime of 9.60 seconds with a standard deviation of 0.07 seconds also over 30 runs. What happened? As briefly discussed when first mentioning loop unrolling in Section 4.2, one potential drawback of loop unrolling is that the number of times the loop is unrolled also causes the amount of generated code to be multiplied by that amount, plus one more time for the tail loop. As a potential consequence, the CPU may not be able to fit the entire loop in its level-1 instruction caches anymore and it thus may need to fetch it from either lower cache levels or from main memory. Running both the regular and the eight times unrolled versions of the program under the Cachegrind tool that's part of the Valgrind [24] dynamic binary analysis framework confirms this. Cachegrind is a tool that can precisely indicate what the CPU's caches and branch predictor are doing at the instruction or source line level by simulating those parts of the CPU. Because Cachegrind needs to simulate parts of a CPU, programs tend to run 20 to 100 times slower under Cachgrind than they normally would. In the case of the LULESH application, granular analysis is not needed a general overview of the program's overal cache usage is sufficient to get the idea. Running both versions of the program under Cachegrind for one minute shows very different cache usage statistics. As expected, the data cache usage for both versions is almost identical, but the instruction cache miss rate increases from 0.30% to 4.18% for the first level instruction caches and from 0.01% to 0.07% for the last level instruction caches when comparing the original and the unrolled versions. And to make matters worse, these numbers also include the identical shared setup phase of the program that takes up the first 20 seconds of the 60 seconds spent profiling. It is thus no surprise that the unrolled version is slower.

However, things change when switching over to Accelerate's GPU backend. When running both the original program and the eight times unrolled version on an NVIDIA RTX 2080 SUPER GPU, the unrolled version now ends up being significantly faster. With that GPU the average runtimes and standard deviations over 30 runs are 3.13 and 0.04 seconds for the original version, with 3.07 and 0.04 for the unrolled version. While this is only a 1.8% speedup, it is free performance that only required a tiny change to the code. With different unrolling amounts the speedup may be smaller or larger and the specific GPU used also plays a role, so optimizing a program like this still involves some degree of experimentation. The reason why only eight times unrolling was tested here is because both the CPU and GPU versions of the embedded LULESH program are compiled at Haskell compile time, and Accelerate's LLVM backends currently does not handle large amounts of LLVM IR well. Particularly on the CPU backend unrolling more than a couple times would increase compile times exponentially. So to make the comparison fairer, the same settings were tested for both backends.

6 Discussion and future work

Having an annotation system for a language like Accelerate opens up the door for both better development experiences and more opportunities to hand optimize embedded programs. The current implementation enables new profiling based optimization workflows whereas previously this process would mostly be driven by trial and error. But more importantly, it also forms a framework that can be built upon. Right now the captured source information is only used to add kernel-level context information for the Tracy frame profiler, but the optimized ASTs contain line-level source locations that can be used to enable step debugging and give more detailed insights into the runtime of a program. Similarly, the optimization options backed by the annotation system mostly serve as a proof of concept. For instance, the loop unrolling optimization could be turned into a more general loop optimization construct that would allow the programmer to better finetune Accelerate's multidimensional, sequential and fused loops.

Another aspect that has not yet been explored in the context of deeply embedded languages is using the annotation and source location systems to improve the experience of the language's compilation process. Source locations can be used to provide detailed warnings during compilation, for instance when the compiler wants to fuse together array operations with incompatible optimization flags or when detecting common mistakes. These warnings can also be suppressed by introducing new annotation based decorators. Aside from diagnostics, annotations could also be used to generate multiple variants of an embedded program that are optimized to run on different target processor features in order to increase the portability of a compiled embedded program that has been embedded inside of an executable binary. In conclusion, the annotation system enables new workflows while also forming the foundations for a new area of exploration.

A Source code

The implementation for the annotation system as implemented in Accelerate can be found in the GitHub repositories for the forked `accelerate`[4] and the `accelerate-llvm`[5] packages. The `Data.Array.Accelerate.Annotations`[6] module in the `accelerate` fork contains a complete guide on how to migrate Accelerate libraries to the new annotated version of Accelerate. Take a look at the `linear-accelerate`[7] fork for an example on how to do this.

[4] https://github.com/robbert-vdh/accelerate/tree/feature/force-inline

[5] https://github.com/robbert-vdh/accelerate-llvm/tree/feature/tracy-annotations

[6] https://github.com/robbert-vdh/accelerate/blob/feature/force-inline/src/Data/Array/Accelerate/Annotations.hs

[7] https://github.com/robbert-vdh/linear-accelerate/tree/feature/annotations

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