Abstract

In previous work, we proposed a new approach to the problem of implementing compilers in a modular manner, by combining earlier work on the development of modular interpreters using monad transformers with the `a la carte approach to modular syntax. In this article, we refine and extend our existing framework in a number of directions. In particular, we show how generalised algebraic datatypes can be used to support a more modular approach to typing individual language features, we increase the expressive power of the framework by considering mutable state, variable binding, and the issue of noncommutative effects, and we show how the Zinc Abstract Machine can be adapted to provide a modular universal target machine for our modular compilers.

Categories and Subject Descriptors D.1.1 [Programming Techniques]: Applicative (Functional) Programming

Keywords compilation, modularity, monads, catamorphisms

1. Introduction

When describing a compiler, the term modularity is frequently used to refer to the decoupling of the various stages of the manipulation of a source program, such as lexing, parsing and code generation. However, there is an alternative axis of modularity on which comparatively little work has been done, namely according to the features supported by the source language, of which there are two varieties. First of all, there are effectful features, such as exception handling and mutable state, and secondly those relating to control flow, such as conditional expressions and recursion schemes.

In previous work [10], we have begun addressing the implementation of programming languages in a modular manner by building upon the work of Liang et al. [26], who showed how to construct interpreters in a modular manner using monad transformers. Combining this approach with the à la carte technique of Swierstra [39] permits a framework in which the syntax of individual language features are described as functions, the semantics of features are given as algebras, and fold operators (or catamorphisms) are used to combine the individual components. The result is a modular approach in which the addition of new language features only requires describing these new features and how they interact with the existing features, rather than modifying any existing definitions.

In this article, we refine and extend the modular compiler framework that was developed in our previous work [10]. More specifically, this article makes the following contributions:

- We show how the use of generalised algebraic datatypes (GADTs) [37] to model particular signature functors and value domains permits certain forms of type constraints to be captured in a clean and modular manner.
- We extend the framework with support for both mutable state and variable binding (via the lambda calculus), improving the potential expressive power of a modular source language.
- We consider the issue of effects that do not commute (such as exceptions and state), which potentially require programs to be compiled in different manners depending on the ordering of the effects, and present two approaches to addressing this issue.
- We define a modular variant of the Zinc Abstract Machine [13] as a suitable universal target machine upon which to execute the resulting code from our modular compilers.

This article is aimed at functional programmers with a basic knowledge of interpreters, compilers, monads and monad transformers, but we do not assume specialist knowledge of modular interpreters or the à la carte technique. As in our previous work, we use Haskell throughout as both a semantic meta-language and an implementation language. The Haskell code associated with the article is available from the authors’ web pages.

The rest of the article is structured as follows. In sections 3 and 4 we briefly recap the theory behind the à la carte technique, and describe the structure of generalised algebraic datatypes and show how they can be used when compiling into a modular target language in section 5. Sections 6 and 7 respectively describe the way in which the lambda calculus and state are introduced to this framework, with the latter leading into a discussion on noncommutative effects in section 8. Two approaches to the issue of compiling programs with respect to varying semantics are proposed in section 9, and we describe the construction of a virtual machine capable of executing the resulting code in section 10. The article concludes with a brief survey of related work in the field of modular compilation, and presents a number of directions for future research.

2. Motivation

When using a programming language compiler, we trust that source programs are translated to target programs in the correct manner. More formally, the correctness of a compiler can be captured by stating that the result of compiling an expression in the source language and then executing the resulting code is equivalent to evaluating the expression with respect to its semantic interpretation, as illustrated by the following commuting diagram:
Given the close interplay between the various datatypes and functions involved in such a statement of correctness, the compiler designer should seek to ensure that any changes made to the source language Expr are minimally disruptive. However, even comparatively minor extensions can require disproportionate modifications. By way of a simple example, consider the source language Expr comprising integers and addition, and a function eval which maps expressions to their integer values:

```
data Expr = Val Int | Add Expr Expr

eval :: Expr -> Int

eval (Val n) = n

eval (Add x y) = eval x + eval y
```

Now suppose that we wish to extend our expression language to support a simple form of exception handling, by adding throw and catch constructors to the Expr datatype:

```
data Expr = Val Int | Add Expr Expr | Throw | Catch Expr Expr
```

In order to accurately reflect the new semantics expected of Expr, we must alter the eval function to take account of the fact the evaluation may either succeed or fail:

```
eval :: Expr -> Maybe Int

eval (Val n) = Just n

eval (Add x y) = case eval x of
  Nothing -> Nothing
  Just n -> case eval y of
    Nothing -> Nothing
    Just m -> Just (n+m)
eval (Throw) = Nothing

eval (Catch x h) = case eval x of
  Nothing -> eval h
  Just n -> Just n
```

As we can see, the required changes to the definition of eval are substantial. These changes could be described in a more compact manner using monadic do-notation, however for expository purposes here we prefer the expanded version above.

In particular, the return type has been changed from Int to Maybe Int to accommodate potential failure, the existing cases for Val and Add are rewritten to reflect the new semantics, which in the case of addition now requires a cumbersome nested case analysis, and new two cases are added to handle the new Throw and Catch constructors. This is far from ideal.

This is a common issue, which Wadler terms the "expression problem" [39]. Multiple solutions have been proposed, and in the next section we describe one such technique for solving the expression problem in Haskell: *datatypes à la carte*.

3. Modular Syntax Revisited

In the previous section, we saw an example of how extending a datatype Expr with new constructors can result in many other definitions needing to be changed as a consequence. The *à la carte* technique [39] allows us to build datatypes such as Expr and functions over them in a modular manner. To illustrate the technique, consider the underlying signatures for both the arithmetic and exception handling features of our simple language:

```
data Arith e = Val Int | Add e e

data Except e = Throw | Catch e e
```

These signatures are trivially functors in Haskell:

```
instance Functor Arith where
  fmap f (Val n) = Val n
  fmap f (Add x y) = Add (f x) (f y)

instance Functor Except where
  fmap _ (Throw) = Throw
  fmap f (Catch x h) = Catch (f x) (f h)
```

For any functor f, its induced recursive datatype, Fix f, is defined as the least fixpoint of f, implemented as follows:

```
data Fix f = In (f (Fix f))
```

Now that we have tied the ‘recursive knot’ of a signature, Fix Arith is a language equivalent to the original Expr datatype which allowed integer values and addition. In turn, Fix Except is a degenerate language in which the only operations that are permitted are the throwing and catching of exceptions. What we require now is a manner in which signatures can be combined. This can be achieved using the notion of *coproduct*:

```
data (f :+: g) e = Inl (f e) | Inr (g e)
```

By taking the coproduct of multiple signatures and taking the least fixpoint of the result, we can now define the syntax of expression in a modular manner by combining the two sublanguages:

```
type Expr = Fix (Arith :+: Except)
```

The above type synonym describes a datatype that is equivalent to the extended Expr of the previous section, but is obtained via the composition of algebraic descriptions of each constituent feature rather than by augmenting an existing datatype. Building values of such types is expedited through the use of *smart constructors* [39] to insert the appropriate fixpoint and coproduct tags, but the details are straightforward and are omitted here.

Now that we can define syntax – and datatypes – in a modular fashion, we consider how to write modular functions.

4. Modular Semantics Revisited

In this section we construct a modular function eval that interprets expressions that are constructed using the *à la carte* technique, by exploiting the functorial nature of signatures, the monadic nature of semantics, and the class system of Haskell.

For suitable functors f, we are able to define a generic fold operator – or *catamorphism* [28] – to act as the foundation upon which functions are defined over Fix f:

```
fold :: Functor f => (f a -> a) -> Fix f -> a
fold f = Fix (fold_t f)
```

The first argument of fold is an f-algebra, which provides the behaviour of each constructor associated with a given signature f. The goal now is to define a modular evaluation function using fold. Such a function will have type Fix f -> m Int for some functor f that captures the syntax of the language, and some monad m that captures the underlying effects of the language. To define functions of this type using fold, we introduce the notion of a *class of evaluation algebras*, defined as follows:
class (Functor f, Monad m) => Eval f m where
evalg :: f (m Int) -> m Int

For example, the evaluation algebra for the signature functor for
arithmeti can be defined in the following manner:

instance Monad m => Eval Arith m where
evalg (Val n) = return n
evalg (Add x y) = do n <- x
      m <- y
      return (n + m)

There are two important points to note about this definition.
First of all, it is parametric in an underlying monad m, as reflected in
the Monad m class constraint. Secondly, the semantics of arithmetic
is now defined in terms of the return and >>= operations of this
monad (as abbreviated by the use of the do notation.)

In a similar manner, the evaluation algebra for exceptions can be
defined as follows, in which the class constraint MonadPlus m
allows us to use the generic mzero and mplus operations to define the
semantics for the two exception handing operations:

instance MonadPlus m => Eval Except m where
  evAlg (Throw) = mzero
  evAlg (Catch x h) = x `mplus` h

Finally, an evaluation algebra can be defined for coproducts in
terms of the algebras for the two underlying signatures:

instance (Eval f m, Eval g m) => Eval (f :+: g) m where
  evAlg (Inl x) = evAlg x
  evAlg (Inr y) = evAlg y

Using the above machinery, we can now define a modular inter-
preter eval by simply folding an evaluation algebra:

eval :: Eval f m => Fix f -> m Int
  eval = fold evalg

This evaluation function can be used with any modularly con-
structed datatype Fix f, provided that algebra instances have been
defined for each component signature, and the associated monad
satisfies all of the required constraints. In this way, we recover the
behaviour of the original, nonmodular, evaluation functions. For
example, consider the following two values (where val, add and
throw are smart constructors that insert the appropriate tags):

three :: Fix Arith
three = val 1 'add' val 2

error :: Fix (Arith :+: Except)
error = val 42 'add' throw

The meaning of these values is then given by our modular eval-
uation function eval. Note that the choice of the underlying monad
can be varied, provided that it satisfies the necessary constraints, as
shown in the two interpretations of three below:

> eval three :: Identity Int
3

> eval three :: Maybe Int
Just 3

> eval error :: Maybe Int
Nothing

We now consider how the à la carte technique can be used to
implement a modular compiler which translates between modular
source and target languages. As we shall see in the next section,
there are a number of difficulties that arise.

5. Introducing GADTs

In our original presentation of a modular compilation frame-
work [10], the compilation function that we constructed using the
techniques described in the previous sections had type:

comp :: Expr -> Code -> Code

The function comp takes an expression in the source language,
Expr, and with the help of an accumulator of type Code, returns a
program in the target language Code that can then be executed on
a stack-based virtual machine. The use of the accumulator both
avoids the need for the (++) operation in the implementation of
the compiler, and as we shall see, plays an important part of our
implementation of exceptions. In the case of our example language
supporting arithmetic and exceptions, the target language Code can
be defined in a modular manner as follows:

type Code = Fix (ARITH :+: EXCEPT :+: NIL)

data ARITH e = PUSH Int e |
      | ADD e |

  data EXCEPT e = THROW e |
      | MARK Code e |
      | UNMARK e |

  data NIL e = NIL

The PUSH and ADD operations simply push a number onto the
stack and add the topmost two numbers, respectively. In turn,
THROW indicates that an exception has been raised and initiates
a search for handler code on the stack, MARK pushes the supplied
handler code onto the stack to be run in the event of an exception,
and UNMARK pops such a handler when it is no longer in scope.
Finally, NIL provides a base case for the code type.

However, there is a problematic issue regarding the MARK con-
structor. While the target language Code is constructed in a mod-
ular manner, it is explicitly defined as a type synonym, fixing the
features of the target language. The use of this fixed type in the
MARK constructor breaks precisely the modularity in the definition
of the target language that we are attempting to obtain. Note that
we cannot simply replace the use of Code in MARK by the parame-
ter type e, because the type for the compilation function comp de-
termines that e will ultimately be instantiated to Code -> Code,
whereas what is required here is simply Code. Similarly, due to the
modular nature of individual language signatures, were the argu-
ment to MARK to be polymorphic in the form (Fix f) a compile-time
error would occur because there is no way to give f its correct type,
as nothing is known about its component signatures. A potential
solution is to extend the EXCEPT signature as follows:

data EXCEPT f e = THROW e |
      | MARK (Fix f) e |
      | UNMARK e |

However, making the underlying functor f into a parameter in
this manner essentially means that every language that uses EXCEPT
now needs to explicitly refer to the overall functor f that captures
all the desired language features, which again breaks modularity.
One approach to resolving this would be to impose a class con-
straint on the signature itself, in the following manner:

data Functor f =>
  EXCEPT e = THROW e
Haskell 2010, being widely considered a misfeature. Our solution is to define those signatures which contain problematic constructors such as MARK as generalised algebraic datatypes (GADTs), which permits individual constructors to be typed explicitly, and with their own class constraints. For example, consider the GADT representation of the nonmodular variant of Expr as described in section 2:

```haskell
data Expr e where
    Val :: Int -> Expr Int
    Add :: Expr Int -> Expr Int -> Expr Int
    Throw :: Expr e
    Catch :: Expr e -> Expr e -> Expr e
```

Note that whilst this representation enables a level of type-safety which was previously unavailable (consider the Add constructor, which dictates that only subexpressions which represent an Int can be added together), we primarily utilise GADTs to leverage existential types into our framework. By describing constructors as methods associated with a type, we can now impose constraints on individual constructors without affecting the datatype as a whole. Using this idea, the signature functor for exception handling in the target language can be redefined as follows:

```haskell
data EXCEPT e where
    THROW :: e -> EXCEPTION e
    MARK :: Functor f => Fix f -> e -> EXCEPTION e
    UNMARK :: e -> EXCEPTION e
```

As a result, we have made two significant improvements over the original definition. Firstly, by abstracting over the syntax of the target language we have avoided the need to refer to an explicitly defined type synonym that must be edited whenever the source language is changed. And secondly, we have placed a constraint on the argument f without including it in the top-level definition of EXCEPTION and without constraining other constructors similarly.

This now suggests that our modular compiler need not target a particular language, but rather any language which meets the appropriate constraints. Key to these constraints is the notion of a subtype relation, captured at the type-level as follows:

```haskell
class (Functor f, Functor g) => f :<: g where
    inj :: f a -> g a
```

The inj method embeds a value within the subtype functor f into a value within the supertype functor g. For example, the subtyping relation Arith :<: (Arith :+: Except) states that the signature functor for arithmetic is a component of the signature functor for both arithmetic and exception handling combined. The modular counterpart of the compilation function comp will have type Fix f -> Fix g -> Fix g, for signature functors f and g that characterise the syntax of the source and target languages respectively. In order to supply an initial value for the accumulator (the second argument), we require that NIL (which represents the empty code fragment) is a subtype of g. Putting this all together, we define the class of compilation algebras as follows:

```haskell
class (Functor f, NIL :<: g) => f :: g where
    compAlg :: f (Fix g -> Fix g) -> Fix g -> Fix g
```

We can now instantiate the compilation algebras for Arith and Except, using the subtype relation to constrain the target functor g to any language which supports the required signatures (where push, add, throw, etc. are smart constructors):

```haskell
instance (ARITH :<: g) => Comp Arith g where
    compAlg (Val n) = push n
    compAlg (Add x y) = x . y . add

instance (EXCEPT :<: g) => Comp Except g where
    compAlg (Throw) = throw
    compAlg (Catch x h) = \c -> mark (h c) (x (unmark c))
```

In the above, expressions are compiled in the expected manner, framed in our modular setting. In particular, values are compiled by pushing the associated integer onto the stack, and addition is compiled by compiling the two subexpressions and adding the resulting two values on top of the stack. In turn, a thrown exception is compiled directly into a corresponding throw instruction in the machine, while catch blocks are compiled by marking the stack with the compiled handler code, producing code for the body of the block, and finally unmarking the stack by removing the topmost handler. Note that the use of continuation-passing style in the compilation algebras is key to compiling the Catch operation [10], and also means that concatenation of code is achieved simply by function composition as shown in the case for addition. The resulting modular compilation function is obtained by folding the compilation algebra over an empty accumulator nilC as follows:

```haskell
nilC :: (NIL :<: g) => Fix g
nilC = In (inj NULL)

comp :: Comp f g => Fix f -> Fix g -> Fix g
comp x = fold compAlg x nilC
```

In conclusion, we have successfully refactored our modular compiler to produce code for a modular target language. Having seen how GADTs can be used in one aspect of our modular framework, in the next section we extend the modular source language with support for variable binding, and discuss the role that GADTs play in defining its modular semantics.

6. Introducing the Lambda Calculus

The ability to abstract over variables in the body of a function is a near-universal feature in programming languages, and in this section we will introduce variable binding into our modular framework using the untyped lambda calculus of Church [7]. Although variables in lambda terms are often given names in the same way that we would name other variables, there are many alternative ways to model bindings, including such approaches as higher order abstract syntax (HOAS) [32] and de Bruijn indices [11], amongst others. The HOAS approach uses the binders of the metalanguage (in our case, Haskell) to describe the binding structure of the language being implemented, via a datatype such as the following:

```haskell
data LamHOAS e = App e e | Lam (e -> e)
```

Unfortunately, LamHOAS is not an instance of the Functor typeclass, as the type parameter e appears in both a contravariant and covariant position within the Lam constructor [2]. As a result, the a la carte technique cannot be used for this datatype, as this is incompatible with the type of fold. For this reason, in this article we use a de Bruijn indexed encoding of the lambda calculus. Using this technique, the syntax of lambda terms can be defined as follows:

```haskell
data Lambda e = Index Int | Abs e | Apply e e
```

1 The GHC pragma allowing this, `XDatatypeContexts`, was removed in Haskell 2010, being widely considered a misfeature.
The number associated with an Index constructor represents a variable, and refers to the number of binders in scope before its binding site. In turn, Abs indicates the presence of a binder and Apply represents the substitution of lambda terms, and is passed both a function body and its argument as subexpressions.

However, by choosing not to use the HOAS approach, a problematic issue arises regarding the Apply constructor. When defining a modular semantics for the Lambda signature, the carrier of the evaluation algebra determines that both of its subexpressions will be typed as m Int. The following attempt at defining the evaluation algebra illustrates the underlying problem:

```
instance Monad m => Eval Lambda m where
evAlg (Apply f x) = f >>= \f' -> ...
```

The definition of Apply cannot be completed in a sensible way, because the semantic domain is not expressive enough. In particular, the result of binding the function body \( f \) has the primitive type \( \text{Int} \) which accepts no arguments. Moreover, binding the result of \( f \) breaks the abstraction that a function body represents.

Our solution to this issue is to extend the semantic domain with support for closures. To do this, we redefine Value as a GADT:

```
data Value m where
     Num :: Int -> Value m
     Clos :: Monad m => [Value m] -> m (Value m) -> Value m
```

In the above, the Num constructor represents an integer value, and the Clos constructor takes as arguments a list of values (which acts as an environment) and a computation which represents an unevaled function body. There are two points to note about this definition. Firstly, we would not be capable of representing closures in this way without the Monad m constraint [42], and secondly, this constraint is imposed on the Value parameter \( m \), rather than a parameter specific to a single constructor.

To make use of closures when giving a semantics to the lambda calculus, we define a class CBVMonad of operations associated with the call-by-value evaluation scheme, which reduces arguments to values prior to their use:

```
class Monad m => CBVMonad m where
    evAlg (Apply f x) = f >>= \f' -> ... 
```

We have presented two separate evaluation algebras, both defined over a signature Lambda. However, despite the differing contexts, Haskell does not permit the two algebras to coexist in the same source file, stating that they are overlapping instances. One possible solution is to define two source signatures LambdaCBV and LambdaCBN which contain appropriately tagged constructors to avoid naming conflicts. An alternative involves parameterising the evaluation algebra class with a tag that can be pattern-matched upon, and we will see more of this idea when describing a solution to the issue of noncommutative effects.

Having successfully implemented variable binding modelled using the lambda calculus in a modular manner, a natural progression is to consider how to introduce the notion of persistent, updatable state to our modular compilation framework.

### 7. Introducing Mutable State

In programming languages, a widely used feature is mutable state variables that can change value over time. In this section, we extend the expressive power of a modular source language by introducing the notion of mutable state. As proof of concept we consider a single integer variable, and the syntax associated with such an updatable value is given by the following signature:

```
mutable x
```

We can now write terms in our modular source language that use variable binding. For example, consider the following example (where apply, abs etc. are the appropriate smart constructors):

```
e :: Fix (Lambda :+: Arith)
e = apply (abs (ind 0)) (add (val 1) (val 2))
```

```
> eval e :: [Value Identity] [Num 3]
```

The source language used in this example is capable of using both variable binding and arithmetic. The expression \( e \) represents the lambda term \( (\lambda x. x)(1 + 2) \), and evaluating \( e \) with respect to the list monad, which can readily be made into an instance of the class CBVMonad, returns the singleton value Num 3.

We are also capable of defining multiple evaluation schemes for terms in the lambda calculus. A common alternative is call-by-name, which does not evaluate arguments before applying them to a function body. The difference between this scheme and the call-by-value scheme which we have just implemented is that environments now contain computations, not values. Another class CBNMonad is needed to reflect this change, defined as follows:

```
class Monad m => CBNMonad m where
    evAlg (Apply f x) = do e <- e
                           (Clos e t) <- f
                           with (t:ctx t) t
```

Constraining by this class allows a call-by-name semantics to be defined for the lambda calculus as follows:

```
instance CBNMonad m => Eval Lambda m where
    evAlg (Apply f x) = do (Clos ctx t) <- f
                           with (x:ctx t) t
```

The above definition is similar to that for call-by-value evaluation, the main difference being that the substitution of terms does not bind the argument \( x \) to a value before using it.

We have presented two separate evaluation algebras, both defined over a signature Lambda. However, despite the differing contexts, Haskell does not permit the two algebras to coexist in the same source file, stating that they are overlapping instances. One possible solution is to define two source signatures LambdaCBV and LambdaCBN which contain appropriately tagged constructors to avoid naming conflicts. An alternative involves parameterising the evaluation algebra class with a tag that can be pattern-matched upon, and we will see more of this idea when describing a solution to the issue of noncommutative effects.

Having successfully implemented variable binding modelled using the lambda calculus in a modular manner, a natural progression is to consider how to introduce the notion of persistent, updatable state to our modular compilation framework.
data State e = Get | Set Int e

In the above, the Get operation represents the current state, and the Set operation takes an integer and an expression that treats this new value as the current state. As with each new feature, we define a class StateMonad of associated operations:

class Monad m => StateMonad m where
  update :: (Int -> Int) -> m Int

The update operation takes a function (Int -> Int) and uses it to alter the state variable. By passing different functions to update, it can be used to define an evaluation algebra for the State signature in the following manner:

instance StateMonad m => Eval State m where
  evAlg (Get) = do n <- update id
                   return (Num n)
  evAlg (Set v c) = update (\_ -> v) >> c

When evaluating a Get constructor, the update operation is passed the identity function id, which leaves the state value unchanged. This value is then bound to v and embedded into the Value domain. In turn, when evaluating a Set constructor, update is passed an anonymous function overwriting the state value to v before evaluating the subexpression c.

We can now write terms in our modular source language that utilise an integer state variable. To illustrate, consider the following two terms x and y, built from languages supporting both arithmetic and state, and state and exception handling, respectively:

x :: Fix (Arith :+: State)
x = set 1 (add get (val 2))

y :: Fix (State :+: Except)
y = set 1 (catch throw get)

Informally, the expression x first sets the state to value 1, then adds the current state to the number 2. In turn, expression y first sets the state to value 1, then immediately throws an exception that is then handled by returning the value of the current state.

In our modular compilation framework, we evaluate a modular expression with respect to a monad that has been constructed by applying the appropriate monad transformers to a base monad, for which purposes we often use the identity monad Identity. The underlying machinery associated with the monad transformer class allows access to the operations associated with each constituent feature (such as throw, update, env etc.) at the top level, with all of the necessary lifting handled automatically.

Every monad transformer comes equipped with an accessor function – such as runS and runE – which allows access to the underlying representation. By first evaluating an expression and then applying the accessor functions of all monad transformers comprising the monad within which the expression was evaluated, we obtain a final value, as illustrated in the following:

newtype StateT s m a =
  S { runS :: s -> m (a, s) }

newtype ErrorT m a =
  E { runE :: m (Maybe a) }

> let a = eval x
  :: StateT Int Identity (Value ())
> runS 0 (runId a)
Num 3

> let b = eval y
  :: ErrorT (StateT Int Identity) (Value ())
> runS (runS 0 (runId b))
Just (Num 1)

In both of the above evaluations, we see that modular expressions involving state are given a semantics by applying the StateT state monad transformer at some point when building the monad, and similarly that the ErrorT exception monad transformer is applied when handling exceptions modularly.

However, an issue arises when considering the order in which certain monad transformers are applied, namely that of non-commutative effects. To illustrate, consider the following:

instance MonadT (StateT s) where
  lift m = S $ \s -> do x <- m
                        return (x, s)

instance MonadT m => MonadT (StateT s m) where
  lift x = S $ \s -> return (x, s)
    (S g) >>= f = S $ \s -> do (x, t) <- g s
                             runS (f x) t

instance Monad m
    => StateMonad (StateT Int m) where
  update f = S $ \a -> (f, a)

The above instantiations and instance declarations of the StateT monad transformer appear at first glance to be no different to that of any other transformer associated with a particular feature. However, in the next section we shall see that defining StateT in this manner leads to non-commutativity concerns.

8. Noncommutative Effects

In the previous section we saw two examples of how monad transformers are used to access the operations needed to define evaluation algebras. However, in some cases separate features can interact in multiple ways, and this is reflected when applying the associated monad transformers in different orders [26]. Consider the following expression demo, constructed from a modular source language which supports arithmetic, mutable state and exception handling:

demo :: Fix (Arith :+: Except :+: State)
demo = set 0
    (catch
      (add (set 1 get) (throw))
      get)

The demo example must be evaluated within a monad that supports both exceptions and state, and therefore must contain both of the relevant monad transformers. It is less obvious, however, that switching the order in which these two transformers are applied has an observable effect on the resulting semantic domain. Assuming that no other features are present, and using Identity as the base monad, the types resulting from the two possible orderings are:

> type LocalM a =
  StateT Int (ErrorT Identity) a
  = Int -> ErrorT Identity (a, Int)
  = Int -> Identity (Maybe (a, Int))
  = Int -> Maybe (a, Int)

> type GlobalM a =
  ErrorT (StateT Int Identity) a
  = StateT Int Identity (Maybe a)
  = Int -> Identity (Maybe a, Int)
  = Int -> (Maybe a, Int)
In particular, when applied to a parameter a, the underlying representation of the LocalM monad takes an Int and either successfully returns a pair (a, Int), or an exception in the form of Nothing. In turn, the GlobalM monad also takes an Int but always returns a pair, where the first element can return Nothing.

More specifically, when handling an exception the ‘local state’ monad restores the state to its most recent value prior to entering the catch-block that threw the exception, while the ‘global state’ monad treats any updates to the state value as irreversible. Specifically, demo produces the value Num 1 when evaluated with respect to GlobalM, and the value Num 0 with respect to LocalM.

These are both sensible results, and depend on how we wish to order the underlying effects. The natural progression at this point is to address the issue of compiling expressions with multiple interpretations, such as demo, in a modular manner. Our modular compiler will currently compile demo to the following code sequence (written using Haskell list notation for simplicity):

> comp demo []
[SET 0, MARK [GET]
[SET 1, GET, THROW, ADD, UNMARK]]

The above code is associated with the global approach to state, as the SET operation within the catch-block cannot be reversed when the THROW instruction is encountered. To model the behaviour associated with the local approach to state, two additional operations are required as seen in the following:

> comp demo []
[SET 0, MARK [RESTORE, GET]
[SAVE, SET 1, GET, THROW, ADD, UNMARK]]

The SAVE operation records the current value of the state on the stack, and in turn the RESTORE operation restores the state to its previous value before handler code is executed.

Both of the above results are valid, corresponding to compiling demo with respect to a particular ordering of effects. However, a modular compiler is only capable of generating one such program in any particular session, as the compilation algebra class is only parameterised by the source and target signatures, with no extra information available concerning the intended semantics.

Clearly, there is a need for a more flexible compilation algebra that is aware of the context of an argument expression. To do this, we must allow the compilation algebra to examine the monad in which the expression is evaluated, as the semantics are defined by the order in which certain monad transformers are applied.

9. Monadic Parameterisation

In this section, we propose two techniques for directing the modular compilation of an expression by inspecting its underlying semantic monad. As we have seen, in our framework we make use of monads that have been constructed by applying a sequence of transformers to a base monad. Taking advantage of the fact that monad transformers are defined as newtypes, we can inspect their constructors at the type level, giving rise to our first technique:

Technique 1. Type-Level Monadic Parameterisation

class (Functor f, Functor g, Monad m)
=> Comp f g m where
  comAlg :: f (m () -> Fix g) -> m () -> Fix g

In the above, the compilation algebra class is parameterised by a monad. The algebra carrier then includes a monadic computation as an argument, however this computation is parameterised by the void type () to indicate that the monad is not explicitly used in the compilation process, but rather used as a context reference.

In this manner, multiple instances of a compilation algebra can be defined for a single source signature by pattern-matching upon constructors associated with monad transformers. This allows for expressions such as demo (defined in the previous section) to be compiled using different schemes for different orderings of effects. For example, the compilation schemes for the two different orderings of exceptions and state can now be defined as follows:

-- global compilation scheme
instance (EXCEPT :<: g, Monad m) =>
  Comp Except g (ErrorT (StateT s m)) where
  comAlg (Throw) = \_ -> throw
  comAlg (Catch x h) = \m c -> mark (h m c)
      (x m (unmark c))

-- local compilation scheme
instance (EXCEPT :<: g, Monad m) =>
  Comp Except g (StateT s (ErrorT Identity)) where
  comAlg (Throw) = \_ -> throw
  comAlg (Catch x h) = \m c -> mark (h m c)
      (save (x m (restore $ unmark c)))

An advantage of this technique is that we only need to match on constructors associated with monad transformers that cause semantics to differ. For example, consider the monad transformer ReaderT, a commutative transformer that can be applied in any order. If ReaderT were to appear between ErrorT and StateT in the above, we could abstract over this transformer using a generic variable t of type MonadT, allowing the programmer to focus on the task of defining algebras only for non-commutative orderings.

Conversely, however, the monadic computation that appears in the carrier of the algebra allows for effectful operations to be manifested by calling its associated methods. The user must be careful to not use any monadic operations when defining a compilation algebra for a particular signature, as we define compilation to be an effect-free mapping between modular source and target languages. Further, this computation cannot be removed from the carrier, as it must be threaded through to subexpressions.

To address this concern, we require a way to provide the compilation algebra with information concerning the ordering of monad transformers, without explicitly passing around the resulting monad. Our solution to this issue is to use GADTs to reify a monad, representing it as a sequence of constructors. We capture this notion with the datatype MTList, defined as follows:

data MTList m where
  Err :: MTList m
  -> MTList (ErrorT m)
  Sta :: ST -> MTList m
  -> MTList (StateT ST m)
  Id :: MTList Identity

Using the auxiliary datatype ST of state types to reify monad transformer parameters, an instance of MTList m represents the monad m by applying the appropriate constructors to Id. To illustrate, the two monads LocalM and GlobalM that are defined in the previous section can be reified as follows:

local :: MTList (StateT s (ErrorT Identity))
local = Sta IntT (Err Id)

global :: MTList (ErrorT (StateT s Identity))
global = Err (Sta IntT Id)
There are two points to be made concerning the above. Firstly, we use the variable a to abstract over the parameter type of the state monad transformer, highlighting that it is the structure of the underlying representation that we are concerned with as opposed to the types involved in its definition. Secondly, the ordering of the monad transformers can now be examined at the function level by using pattern matching on the data constructors Sta and Err.

We can now replace the monadic computation m () in the carrier of the compilation algebra with its reified representation MTList m. In doing this, we eliminate the concern that effectful operations may ‘leak’ into the compilation process by removing the possibility of invoking any monadic operations. This leads to the definition of our second technique.

Technique 2. Function-Level Monadic Reification

```haskell
class (Functor f, Functor g)
  => Comp f g where
  comAlg :: f (MTList m -> Fix g -> Fix g)
          -> MTList m -> Fix g -> Fix g
```

By performing case analysis on the MTList argument, we can now define multiple compilation schemes within a single compilation algebra instance, as seen in the following:

```haskell
instance (EXCEPT ::<: g) =>
  Comp Except g where
  comAlg (Throw) = \_ -> throw
  comAlg (Catch x h) = \_ \_ -> case m of
    (Err (Sta s t)) -> mark (h m c)
    (x m (unmark c))
    (Sta s (Err t)) -> mark (h m c)
    (save (x m (restore $ unmark c)))
```

Particularly important in the above is that the compilation algebra is no longer parameterised by a monad m, highlighting the fact that a modular compiler is informed by a monad, rather than defined in terms of one. However, the use of MTList to reify the context comes at the cost of being unable to abstract over commutative monad transformers, as is possible when using the first technique. The user can still abstract over a base monad using this second technique, but must explicitly define any intermediate monad transformers that are applied between a conflicting pair.

Both of these techniques are potential solutions for the issue of modular compilation in the presence of noncommutative effects. We highlight that modular compiler instances exist for all of the other features described up to this point, and can be found in the associated code on the authors’ websites. In the next section, we discuss how the techniques and extensions we have discussed up to this point apply to executing modular code produced by a modular compiler on a modular virtual machine.

10. A Modular Virtual Machine

In previous work [10], we defined a modular virtual machine in terms of an execution algebra targeting a stateful computation, parameterised by a modular datatype Stack, as follows:

```haskell
type StackT m a = StateT Stack m a
```

```haskell
class (Monad m, Functor f) => Exec f m where
  exAlg :: f (StackT m ()) -> StackT m ()
```

By defining Stack as a modular datatype, functions over Stack were defined as folds, resulting in a significant amount of boilerplate in order to retrieve and place values onto the modular stack. To illustrate, consider the quite convoluted – code from [10] defining the modular implementation of stack-based addition:

```haskell
add :: Monad m => StackTrans m ()
add = do x <- pop; y <- pop
    case (extract x, extract y) of
      (VAL n _, VAL m _) -> push (n + m)
```

This code is more complicated than it should be, as a result of the auxiliary functions that need to be defined in order to manipulate the modular stack. The extract operation in the above explicitly describes the type of value which we can pop off of the top of the stack. Such stack-based operations are conceptually simple, however the need to write them in terms of folds and monads leaves much to be desired. In order to eliminate this overhead, in our library for building modular compilers we use the Haskell list type to represent stacks, rather than a modular datatype of modular values.

There are two reasons for this change. Firstly, the stack may be considered to be an auxiliary structure, as the real work is done by the compilation and execution algebras upon the source and target languages, which we have already defined in a modular manner. As we have demonstrated in previous work, it is possible to construct a modular stack, however the amount of boilerplate required when extending a (module) source language with new features quickly becomes prohibitive, and detracts from the main task at hand. Secondly, while the two representations of the stack are equivalent, we feel that it is easier for a user to conceptualise stack-based operations in terms of lists. This, combined with access to the Haskell library functions, improves the overall extensibility of our modular compilation framework.

More important than the choice of representation type for the stack is the fact that the stack transformer StackT is defined by applying the state transformer to a base monad m. In our original presentation, the intent of the monad m was to help execute modular code in the presence of noncommutative effects. However, as we have seen in the previous section, by inspecting the ordering of monad transformers when compiling a modular source expression, we have solved this issue, and the monad is no longer required.

In section 6, we defined a semantics for lambda terms modelled using de Brujin indices according to the call-by-value evaluation scheme. To execute such terms on a modular virtual machine, we define a compilation algebra mapping between lambda terms and the instruction set LAMBDA associated with the Zinc Abstract Machine, a call-by-value variant of the Krivine machine [8]. This compilation scheme is captured by the following function C:

```haskell
C[\lambda x. t] = [CLS (C[\lambda x. t] + [RE]]
C[f x] = C[x] + C[f] + [APP]
```

The instructions of the Zinc Abstract Machine (IND, CLS, RET and APP) usually operate upon a pair of lists, the first representing an environment and second representing a stack. In this article we also make use of a third list, representing a state stack.

The resulting modular execution algebra class and associated data structures are defined in the following manner:

```haskell
class Functor f => Execute f where
  exAlg :: f ZAM -> ZAM
```

```haskell
data VALUE e where
  NUM :: Int -> VALUE e
  REC :: Int -> VALUE e
  HAN :: Execute f => Fix f -> VALUE e
  CTN :: ZAM -> [VALUE e] -> VALUE e
  CLO :: Execute f => Fix f -> [VALUE e] -> VALUE e
```

```haskell
newtype StateT s a = StateT s a
```

```haskell
= S { runSM :: s -> (a, s) }
```
In the above, the type `VALUE` of data that is manipulated on the stack is defined as a GADT. In turn, the state transformer `Stack` is defined in a similar manner to the state monad transformer, but without the presence of a monad.

Further, the type synonym `ZAM` is defined as a stateful computation parameterised by a triple of lists defined over `VALUE`, and finally we define the execution algebra class, targeting the modular representation of the Zinc Abstract Machine.

In previous work [10], the instructions that are executed upon a virtual machine have had intuitive operational transitions. For example, the `PUSH` operation pushes a value onto the top of a stack. However, the operational transitions associated with the instructions of the Zinc Abstract Machine are more complicated, and are defined by a transition relation on `(Code,Env,Stack)` triples:

<table>
<thead>
<tr>
<th>Code</th>
<th>Env</th>
<th>Stack</th>
</tr>
</thead>
<tbody>
<tr>
<td>IND n</td>
<td>e</td>
<td>s → e</td>
</tr>
<tr>
<td>CLS k</td>
<td>e</td>
<td>s → e</td>
</tr>
<tr>
<td>APP</td>
<td>e</td>
<td>v: [d, t] → d</td>
</tr>
<tr>
<td>RET</td>
<td>c</td>
<td>v: f: s → d</td>
</tr>
</tbody>
</table>

In the above, the notation `[c, e]` is shorthand for a closure consisting of a code fragment `c` and associated environment `e`. Instantiating the resulting execution algebra is straightforward:

```haskell
instance Execute LAMBDA where
  exAlg (IND i c) = S $ \(e, s, stk) -> runSM c (e, (e !! i):stk)
  exAlg (CLS k c) = S $ \(e, s, stk) -> runSM c (e, (CLS k e):stk)
  exAlg (APP c)   = S $ \(e, s, stk) -> case stk of (v:(CTN c' e'):stk') -> runSM c' (e', s, v:stk')
  exAlg (RET c)   = S $ \(e, s, stk) -> case stk of (v:(CLS k e':stk') -> runSM k' (v:e', s, ((CTN c e):stk'))

instance Execute ARITH where
  exAlg (PUSH n c) = S $ \(e, s, stk) -> runSM c (e, s, (NUM n):stk)
  exAlg (ADD c)    = S $ \(e, s, stk) -> case stk of (m:(NUM m):stk') -> runSM c (e, s, (NUM (n + m)):stk')
```

It is worth noting that had the representation of the stack types not been changed to lists, the amount of boilerplate required to make use of the state stack. These 'new' instances of the execution algebra are defined in the following manner:

```haskell
instance Execute EXCEPT where
  exAlg (THROW c) = S $ \(e, s, stk) -> case dropWhile (not . isHAN) stk of ((HAN h):stk') -> runSM h' (e, s, stk')
  exAlg (MARK h c) = S $ \(e, s, stk) -> runSM c (e, s, (HAN h):stk)
  exAlg (UNMARK c) = S $ \(e, s, stk) -> case dropWhile (not . isHAN) stk of (HAN ..):stk' -> runSM c (e, s, stk')
  exAlg (SAVE c)   = S $ \(e, s, stk) -> case s of (NUM n):.. -> runSM c (e, s, (REC n):stk)
  exAlg (RESTORE c) = S $ \(e, s, stk) -> case dropWhile (not . isREC) stk of (REC n):stk' -> runSM c (e, (NUM n):s, stk')

instance Execute STATE where
  exAlg (GET c)   = S $ \(e, s, stk) -> case s of (NUM n):.. -> runSM c (e, z, (NUM n):stk)
  exAlg (SET c)   = S $ \(e, s, stk) -> runSM c (e, (NUM n):s, stk)
```

Our modular virtual machine is now defined by folding the algebra over an initial triple of empty stacks:

```haskell
exec :: Execute f => Fix f -> ZAM
exec f = S $ \_ -> runSM (fold exAlg f) ([], [], [])
```

In section 6, we saw an example of the idea that lambda terms can be interpreted according to more than one evaluation scheme. In particular, we showed how both the call-by-value and call-by-name schemes are implemented. The Zinc Abstract Machine executes lambda terms in a manner corresponding to call-by-value, however its instruction set can be used in the definition of an alternative compilation scheme `K`, equivalent to the Krivine machine [8], which is defined recursively as follows:

```haskell
K[n]  = [IND n, APP]
K[\_] = [POP] ++ K[\]
K[x]  = [CLS[K[x]]] ++ K[x]
```

The only instruction that we have not yet specified the behaviour for is `POP`, and this simply moves the topmost value of the stack onto the environment stack. The resulting execution algebra that executes lambda terms with respect to call-by-name is defined over the instruction set `LAMDBA'` (to avoid constructor naming conflicts, as seen in section 6) in the following manner:

```haskell
instance Execute LAMDBA' where
  exAlg (IND' i c) = S $ \(e, s, stk) -> runSM c (e, s, (e !! i):stk)
  exAlg (CLS' k c) = S $ \(e, s, stk) -> runSM c (e, (CLS k e):stk)
  exAlg (APP' c)   = S $ \(e, s, stk) -> case stk of (v:(CTN c' e'):stk') -> runSM c' (e', s, v:stk')
  exAlg (RET c)    = S $ \(e, s, stk) -> case stk of (v:(CLS k e':stk') -> runSM k' (v:e', s, ((CTN c e):stk'))
```

11. Related Work
In this section we briefly review a range of previous work that is related to our approach to the implementation of compilers in a
modular manner. We consider a number of categories of related work, each described in a separate subsection.

**Modular Interpreters and Monad Transformers.** Whilst the work we present here is primarily focussed on the process of compilation between modular languages, a key source of inspiration has been the work of Liang, Hudak and Jones [26]. This article introduced the idea of defining the syntax of a language in a modular manner, but did not consider how this approach could be exploited further to define the meaning of programs using fold operators. It also discussed the issue of combining different effects by combining their underlying monads, giving rise to the notion of base monads, monad transformers, and noncommutative effects.

**Modular Compilers Based on Monad Transformers.** There is a long line of previous work on generating compilers starting from interpreters by using the technique of partial evaluation [21, 25, 44]. Harrison and Kamin showed how a modular compiler can be developed by applying partial evaluation to a modular interpreter that is structured using the monadic approach summarised above. In particular, by writing the interpreter in continuation-passing style [34] and then partially evaluating, we can obtain a compiler and associated virtual machine. However, the use of monad transformers in this work was primarily for the purpose of introducing intermediate data structures in the virtual machine, whereas in our setting these are used explicitly to model individual language features. Moreover, Harrison’s work did not consider the issue of noncommutative effects, or the role that types can play in structuring and informing the development of a modular compiler.

**Compilation as Metacomputation.** Extending their previous work, Harrison and Kamin identified that metacomputations (computation producing computations) naturally arise in the compilation process [16]. Metacomputations can be classified into two distinct varieties capturing different aspects of the behaviour of a program: static, such as code generation and optimisation, and dynamic, such as stack and state manipulation. This ‘staging’ of a program can readily be implemented using monads constructed using transformers. The notions of static and dynamic metacomputations correspond, respectively, with the compilation and execution algebras that we implement using the à la carte technique.

**Modular Compilers & Their Correctness Proofs.** Harrison’s PhD thesis consolidates the work of the two items above by describing the construction and verification of reusable compiler building blocks (RCBBs) for various features of a source language. Two distinct approaches to RCBBs are proposed, namely as metacomputations and monadic code generators. The approach to the latter is similar to that taken in this article, in the sense that for each individual feature a function compile :: Source -> m Target is defined. The correctness of a compiler can then be verified by proving relations between the standard and compilation semantics of any given RCBB. The thesis concludes by describing limitations on the combinations of RCBBs for a non-trivial language in much the same way as the usage of monad transformers in our own work dictates the need to consider multiple compilation schemes.

**Virtualising the Monad Stack.** The work of Schrijvers and Oliveira [36] considers two issues that surround the use of monad transformers, namely that lifting operations through transformer stacks is tedious, and extending transformer stacks often requires modifying existing code with additional lifting operations. The heart of the issue is that monad transformer stacks are concrete interfaces, and a mechanism for interacting with them in a more abstract manner is missing. By introducing the concepts of monad zippers and structural masks, Schrijvers and Oliveira show how such an abstract interface to transformer stacks can be developed. It would be interesting to consider to what extent this approach can be applied in the context of developing modular compilers, in particular how it may be used to provide an alternative approach to defining multiple algebras for an individual language feature.

**Monatron.** The standard monad transformer library in Haskell [15] suffers from the problem of requiring a quadratic number of instance declarations to lift monadic operations through each possible monad transformer. Jaskelioff’s Monatron library [18] solves this problem by lifting operations through monad transformers in a uniform manner, underpinned by ideas from Plotkin and Power’s algebraic theory of effects [17, 33], in which the operations used to manifest effects are themselves treated as primitives rather than being derived from the type of an underlying monad. As an application, Jaskelioff’s PhD thesis [19] describes how Monatron can be used to both define a modular interpreter for a multi-feature language and its modular operational semantics [20].

**Eff.** The recently developed language Eff of Bauer and Pretnar [3] presents another approach to handling non-commutative effects, based upon the algebraic theory of effects. Whereas in our approach we explicitly define a concrete monad and – where possible – a monad transformer for a particular computational effect, in the algebraic approach one specifies the operations that are required by a program and their desired properties (i.e. their type), rather than how they are actually implemented. In this manner, the issue of combining effects becomes one of combining mathematical specifications rather than their underlying concrete implementations. These ideas are realised in Eff via the notion of effect handlers, which describe how the operations of a given effect manifest themselves. Importantly, if a program contains – for example – operations corresponding to choice and exceptions, the handlers for both choice and exceptions need to be invoked in order to return a final (pure) result. However, the user is free to call the necessary handlers in any order they wish, allowing them to decide the intended semantics of a program, rather than leaving this to the Haskell type system. A Haskell library based upon the ideas that underlie the Eff language has been implemented by Visscher [41].

**Extensible Effects.** Since presentation of this paper at IFL 2013, another approach to addressing the issues with monad transformer stacks has been proposed, by Kiselyov, Sabry and Swords [23]. Building on the algebraic approach to effects, this approach makes a separation between the use of effects in a program and how these effects are actually realised in practice. In particular, this is achieved using the idea of a ‘bureaucracy’ of handlers, which provides a stratified view of an open union of handlers, such that if a particular handler is not able to deal with a particular effect, it passes control to a higher authority in the bureaucracy. In this approach, decisions about the order in which effects are combined are made by the user at the type level, without the need to change the program text itself if the order of effects is changed.

**Effect Handlers in Action.** The topic of effect handlers has received considerable attention in recent years. However, to date they have not yet seen widespread use in practice. To address this, recent work of Kammar, Lindley and Oury [22] seeks to popularise work on the handlers of algebraic effects by presenting a number of examples of handlers in use, and describe the implementation of a handler library in Haskell, OCaml, Standard ML and Racket. Further, this work provides some parallels with which to compare effect handlers, for those new to the concept: in particular, they can be seen as folds over free monads, a generalisation of exception handlers [4], or a safer way of invoking delimited continuations.

**Compositional Datatypes.** Recent research by Bahr and Hvitved [1, 2] extends the à la carte technique in a range of directions, including alternative recursion schemes, generic programming, datatypes with holes (zippers) and mutual recursion. Building upon this work, it was then shown how languages with variable binding can also be implemented in the same setting, using Chlipala’s parametric higher-order abstract syntax (PHOAS) [6]. However, the
use of PHOAS requires higher-order folds [27], which adds significantly to the overall complexity of the approach. Furthermore, Day and Bahr [9] developed a modular framework for implementing control structures (such as conditionals and loops), making use of both the à la carte framework and Oliveira and Cook’s purely functional representation of graphs [31], in this case control-flow graphs. It is worth noting that the issue of noncommutativity, a key issue in this paper, does not arise when considering modular control structures.

First-Class Syntax and Semantics. Viera’s doctoral thesis [40] presents a treatment of modular compilers using attribute grammars (described in the next section), instead of the à la carte technique used in this paper. More specifically, a single attribute grammar – defined via a GADT – describes a single aspect of a language, and these grammars are treated as first-class Haskell values. The semantics of a particular component is given via the attributes associated with its corresponding grammar, and the interfaces between multiple grammars are defined as types, ensuring that nonsensical combinations are rejected via Haskell’s type system. In contrast to the work presented in this paper, the notion of compiler discussed is at a much lower level than our own: considerable effort goes into defining modular parsers and token dictionaries, instead of focussing exclusively on the translation between languages. There are also a number of standalone attribute grammars with a particular focus on extensible language implementation, such as LISA [30], JastAdd [14] and Silver [45].

12. Conclusions

In this article we have described a number of extensions and improvements to a framework for defining compilers that are modular with respect to the various features provided by a source language. In particular, we have demonstrated how generalised algebraic datatypes allow for increased flexibility and type-safety when targeting modular languages, how variable binding and mutable state can be handled in our modular framework, and how the issue of noncommutative effects such as exceptions and state can be resolved by pattern matching on the structure of the underlying monad transformers either at the type or the value level.

However, this is by no means the end of the modular compilation story, and much remains to be done. We briefly outline a number of directions for further work below.

Additional Computational Features. We wish to investigate the extension of our framework with support for additional features, in particular other forms of control flow such as recursion and continuations. Furthermore, it would be interesting to see how one might exploit the algebraic theory of effects to give a principled understanding of how easy it may be to integrate a new feature based upon the new operations that it provides.

Attribute Grammars. The Utrecht University Attribute Grammar Compiler (UUAGC) [38] is a Haskell preprocessor which simplifies the construction of catamorphisms over tree-like structures, whilst taking advantage of inherited attributes that are passed down the tree, such as environments, and synthesised attributes that are passed up the tree, such as computed results. Moreover, the ability of the UUAGC to define the constructors of a datatype in multiple distinct locations and define its attributes and semantics separately provides an alternative approach to modular programming. We intend to examine the extent to which the UUAGC system is suitable for generating modular compilers.

Indexed Type Families. In Haskell, the indexed type family extension [5], which permits ad-hoc overloading of datatypes, may prove useful in explicitly declaring a link between the signature functors of a source and target language for a particular effect. For example, in section 5 we defined an evaluation algebra that mapped terms that are constructed from the source functor \( \text{Arith} \) into terms constructed from the target functor \( \text{ARITH} \). At present, we are capable of compiling into any target language, provided it supports the requisite signatures. Declaring a type family with a functional dependency in order to define a mapping from a source language \( \text{Fix}_f \) to a target language \( \text{Fix} \) (\( \text{Target}_f \)) would ensure that the target language is minimal, and remove the requirement that the user define the target language in advance.

Automatic Context Inference. A recent observation \(^2\) is that it may be possible to use the ordering of signature functors in the type of an expression to automatically infer the monadic context within which we wish to evaluate it. For example, from a term with signature \( (\text{Arith} : + : \text{Except} : + : \text{State}) \), we might infer that it is to be evaluated in a monad that is built up from the identity monad corresponding to \( \text{Arith} \) by first applying the exception monad transformer, and then applying the state monad transformer. Such an interpretation may be useful as the default behaviour, which the user could override if they wished.

Modular Correctness Proofs. If one constructs compilers in a modular manner, it is natural to ask if the proof of correctness of such a compiler can also be structured in a modular manner. A suitable starting point for such an exploration would be Acerbi’s encoding \(^3\) of our previous work on modular compilers in the Coq theorem prover, combined with the use of metatheory à la carte (MTC) [12], a recent Coq library designed for reasoning about modular definitions using Mendler-style catamorphisms [29].

Alternative Target Languages. At present, we compile into a stack-based target language. It would also be useful to consider how our framework can be adapted to other forms of target language, in particular register-based languages such as LLVM [24], which is used as the target language for many imperative language compilers, or logic-based languages such as System F [35], a variant of which is used as the target language for GHC.

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