A right-to-left type system for mutually-recursive value definitions

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Abstract. In call-by-value languages, some mutually-recursive value definitions can be safely evaluated to build recursive functions or cyclic data structures, but some definitions (let rec \( x = x + 1 \)) contain vicious circles and their evaluation fails at runtime. We propose a new static analysis to check the absence of such runtime failures.

We present a set of declarative inference rules, prove its soundness with respect to the reference source-level semantics of Nordlander et al. (2008), and show that it can be (right-to-left) directed into an algorithmic check in a surprisingly simple way.

Our implementation of this new check replaced the existing check used by the OCaml programming language, a fragile syntactic/grammatical criterion which let several subtle bugs slip through as the language kept evolving. We document some issues that arise when advanced features of a real-world functional language (exceptions in first-class modules, GADTs, etc.) interact with safety checking for recursive definitions.

Keywords: recursion · type systems · programming languages

1 Introduction

OCaml is a statically-typed functional language of the ML family. One of the features of the language is the let rec operator, which is usually used to define recursive functions. For example, the following code defines the factorial function:

```ocaml
let rec fac x =  
  if x = 0 then 1  
  else x * (fac (x - 1))
```

Beside functions, let rec can also be used to define recursive values, as in the following definition of an infinite list ones where every element is 1.

```ocaml
let rec ones = 1 :: ones
```

Note that this “infinite” list is actually cyclic: it uses a finite amount of memory, because it is composed of a single cons-cell referencing itself.

However, not all recursive definitions can be computed. For example, here is a definition that is justly rejected by the compiler:
let rec x = 1 + x

The variable \( x \), which is typed as an integer, is used in its own definition. Computing \( 1 + x \) requires the variable \( x \) to have a known value: this definition contains a vicious circle, and any runtime evaluation strategy would fail if it is accepted by the language.

Functional languages have different ways to deal with recursive values. Standard ML takes a simple approach, rejecting all recursive definitions that are not recursive function values. At the other extreme, the lazy language Haskell accepts every well-typed recursive definition, although some of them lead to infinite computation. In OCaml, safe cyclic-values definitions are accepted, and they are occasionally useful.

For a cute example, consider an interpreter for a small programming language with datatypes for ASTs and for values:

```
type ast = Fun of var * expr | ...  
type value = Closure of env * var * expr | ...
```

Our interpretation function `eval` takes an environment of type \((\text{var} \times \text{value})\) list and an ast and builds a value

```
let rec eval env = function  
| ...  
| Fun (x, t) -> Closure(env, x, t)
```

Now consider adding an ast constructor `FunRec of var * var * expr` for recursive functions; `FunRec ("f", "x", t)` represents the recursive function \((\text{fix} f. \lambda x. t)\), or `let rec f x = t in f`. Our OCaml interpreter can use value recursion to build a closure for these recursive functions, without changing the type of the Closure constructor: the recursive closure simply adds itself to the closure environment.

```
let rec eval env = function  
| ...  
| Fun (x, t) -> Closure(env, x, t)  
| FunRec (f, x, t) ->  
  let rec clo = Closure((f,clo)::env, x, t) in clo
```

Until recently, the static check used by OCaml to reject vicious recursive definitions relied on a syntactic/grammatical description. While we believe that the check as originally defined was correct, it proved fragile and difficult to maintain as the language evolved and new features interacted with recursive definitions. Over the year, several bugs were found where the check was unduly lenient. In conjunction with OCaml’s efficient compilation scheme for recursive definitions (Hirschowitz et al., 2009), this leniency resulted in memory safety violations, and led to segmentation faults.

Seeking to address these problems, we have designed and implemented a new recursive check for safety of recursive definitions, based on a novel static analysis, formulated as a simple type system. Our implementation was integrated into the main OCaml distribution in August 2018.
In the present document, we formally describe our analysis, presented using a core ML language restricted to the salient features for value recursion (§3). We present inference rules (§4), study the meta-theory of the analysis, and show that it is sound with respect to the operational semantics proposed by Nordlander et al. (2008) (§5). We also discuss the challenges caused by scaling the analysis to OCaml (§6), a full-fledged functional language, in particular the delicate interactions with non-uniform value representations (§6.2), with exceptions and first-class modules (§6.3), and with Generalized Algebraic Datatypes (GADTs) (§6.4).

Contributions We studied related work in search of an inference system that could be used, as-is or with minor modifications, for our analysis – possibly neglecting finer-grained details of the system that we do not need. We did not find any. Existing systems, detailed in our related work section (§7.1), have a finer-grained handling of functions (in particular ML functors), but coarser-grained handling of cyclic data, and most do not propose effective inference algorithms.

We claim the following contributions:

– We propose a new system of inference rules that captures the needs of OCaml (or F♯) recursive value definitions, previously described by ad-hoc syntactic restrictions (§4). We implemented a checker derived from these rules, scaled up to the full OCaml language and integrated in the OCaml implementation.
– We prove the soundness of our analysis with respect to a pre-existing source-level operational semantics: accepted recursive values evaluate without vicious-circle failures (§5).
– Our analysis is less fine-grained on functions than existing works (thanks to a less demanding problem domain), but in exchange it is noticeably simpler.
– The idea of right-to-left computational interpretation (from type to environment) helps bring the complexity down – a declarative presentation designed for a left-to-right reading would be more complex. It is novel in this design space and could inspire other inference rules designers.

2 Overview

2.1 Access modes

Our analysis is based on the classification of each use of a recursively-defined variable using “access modes” or “usage modes” \( m \). These modes represent the degree of access needed to the value bound to the variable during evaluation of the recursive definition.

For example, in the recursive function definition

\[
\text{let rec } f = \text{fun } x \to \ldots f \ldots
\]

the recursive reference to \( f \) in the right-hand-side does not need to be evaluated to define the function value \( \text{fun } x \to \ldots \); since its value will only be required
later, when the function is passed an argument. We say that, in this right-hand-side, the mode of use of the variable \( f \) is Delay.

In contrast, in the vicious definition

\[
\text{let rec } x = 1 + x
\]

evaluation of the right-hand side \( 1 + x \) involves access the value of \( x \); we call this usage mode a Dereference. Our static check will reject (mutually-)recursive definitions that access a recursive name under this mode.

Some patterns of access fall between the extremes of Delay and Dereference. For example, in the cyclic datatype construction

\[
\text{let rec } \text{ones} = 1 :: \text{ones}
\]

the recursively-bound variable \( \text{ones} \) appears on right-hand side without being placed inside a function abstraction. However, since it appears in a “guarded” position, directly beneath the value constructor \( :: \), evaluation only needs to access its address, not its value. We say that the mode of use of the variable \( \text{ones} \) is Guard.

Finally, a variable \( x \) may also appear in a position where its value is not inspected, neither is it guarded beneath a constructor, as in the expression \( x \), or

\[
\text{let } y = x \text{ in } y
\]

for example. In such cases we say that the value is “returned” directly and use the mode Return. As with Dereference, recursive definitions that access variables at the mode Return, such as the following

\[
\text{let rec } x = x
\]

would be under-determined and are rejected.

We also use a last Ignore mode to classify variables that are not used at all in a term.

### 2.2 A right-to-left inference system

The central contribution of our work is a simple system of inference rules for a judgment of the form \( \Gamma \vdash t : m \), where \( t \) is a program term, \( m \) is an access mode, and the environment \( \Gamma \) maps term variables to access modes – modes classify terms and variables, playing the role of types in usual type systems. The example judgment

\[
x : \text{Dereference}, y : \text{Delay} \vdash (x + 1, \text{lazy } y) : \text{Guard}
\]

can be read alternatively

**left-to-right:** If we know that \( x \) can safely be used in Dereference mode, and \( y \) can safely be used in Delay mode, then the pair \( (x + 1, \text{lazy } y) \) can safely be used under a value constructor (in a Guard-ed context).

**right-to-left:** If a context accesses the program fragment \( (x + 1, \text{lazy } y) \) under the mode Guard, then this means that the variable \( x \) is accessed at the mode Guard, and the variable \( y \) at the mode Delay.

This judgment uses access modes to classify not just variables, but also the constraints imposed to a subterm by its surrounding context. If a context \( C[\square] \) uses its subterm \( \square \) at the mode \( m \), then any derivation for \( C[t] : \text{Return} \) will contain a sub-derivation of the form \( t : m \).

In general, we can define a notion of mode composition: if we try to prove \( C[t] : m' \), then the subderivation will check \( t : m'[m] \), where \( m'[m] \) is the composition of the access-mode \( m \) under a surrounding usage mode \( m' \), and \text{Return} is neutral for composition.
Our judgment $\Gamma \vdash t : m$ can be directed into an algorithm following our right-to-left interpretation. Given a term $t$ and an mode $m$ as inputs, our algorithm computes the least demanding environment $\Gamma$ such that $\Gamma \vdash t : m$ holds.

For example, the inference rule for function abstractions in our system is as follows:

$$\frac{\Gamma, x : m_x \vdash t : m[\text{Delay}]}{\Gamma \vdash \lambda x. t : m}$$

The right-to-left reading of the rule is as follows. To compute the constraints $\Gamma$ on $\lambda x. t$ in a context of mode $m$, it suffices to check the function body $t$ under the weaker mode $m[\text{Delay}]$, and remove the function variable $x$ from the collected constraints – its mode does not matter. If $t$ is just a variable $y$ and $m$ is $\text{Return}$, we get the environment $y : \text{Delay}$ as a result.

Given a family of mutually-recursive definitions $\text{let rec } (x_i = t_i)^{i \in I}$, we run our algorithm on each $t_i$ at the mode $\text{Return}$, and obtain a family of environments $(\Gamma_i)^{i \in I}$ such that all the judgments $(\Gamma_i \vdash t_i : \text{Return})^{i \in I}$ hold. The definitions are rejected if one of the $\Gamma_i$ contains one of the mutually-defined names $x_j$ under the mode $\text{Dereference}$ or $\text{Return}$ rather than $\text{Guard}$ or $\text{Delay}$.

3 A core language of recursive definitions

Family notation We write $(\ldots)^{i \in I}$ for a family of objects parametrized over an index $i$ over finite set $I$. Furthermore, we assume that index sets are totally ordered, so that the elements of the family are traversed in a predetermined linear order; we write $(t_{i_1})_{i_1 \in I_1}, (t_{i_2})_{i_2 \in I_2}$ for the combined family over $I_1 \uplus I_2$, with the indices in $I_1$ ordered before the indices of $I_2$. When there is no need for precision, we often omit the index set, writing $(\ldots)^{i}$. Our syntax, judgments, and inference rules will often use families: for example, $\text{let rec } (x_i = t_i)^{i \in I}$ is a mutually-recursive definition of families $(t_i)^{i}$ of terms bound to corresponding variables $(x_i)^{i}$ – assumed distinct, we follow the Barendregt convention. Sometimes a family is used where a term is expected, and the interpretation should be clear: for example, when we say “$\Gamma_i \vdash t_i : m_i$ holds”, we implicitly use a conjunctive interpretation: each of the judgments in the family holds. Finally, we write $\emptyset$ for the empty family.

3.1 Syntax

Figure 1 introduces a minimal subset of ML containing the interesting ingredients of OCaml’s recursive values:

- a multi-ary $\text{let rec}$ binding $\text{let rec } (x_i = t_i)^{i}$ in $u$,
- functions ($\lambda$-abstractions) $\lambda x. t$ to write recursive occurrences whose evaluation is delayed. (OCaml has additional constructs for delaying computation, such as $\text{lazy}$ values and object literals.)
- datatype constructors $K(t_1, t_2, \ldots)$ to write (safe) cyclic data structures; these stand in both for user-defined constructors and for built-in types such as lists and tuples
Fig. 1. Core language syntax

- shallow pattern-matching \((\text{match } t \text{ with } (K_i (x_{i,j}) \rightarrow u_i)^j)\), to write code that inspects values, in particular code with vicious circles.

The following common ML constructs do not need to be primitive forms, as we can desugar them into our core language. In particular, the full inference rules for OCaml (and our check) exactly correspond to the rules (and check) derived from this desugaring.

- \(-n\)-ary tuples are a special case of constructor:
  \((t_1, t_2, \ldots, t_n)\) desugars into \(\text{Tuple}_n(t_i)^{\in [1;n]}\).
- Non-recursive \(\text{let}\) binding are recursive bindings with recursive mode \(\text{ignore}\):
  \(\text{let } x = t \text{ in } u\) desugars into \(\text{let rec } x = t \text{ in } u\).
- Conditionals are a special case of pattern-matching:
  \(\text{if } t \text{ then } u_1 \text{ else } u_2\) desugars into \(\text{match } t \text{ with } (\text{True} \rightarrow u_1 | \text{False} \rightarrow u_2)\).

Besides dispensing with many constructs whose essence is captured by our minimal set, we further simplify matters by using an untyped ML fragment: we do not need to talk about ML types to express our check, or to assume that the terms we are working with are well-typed.\(^4\) (Untyped algebraic datatypes might make you nervous, but they work just fine, and were invented in that setting. A \text{match} form gets stuck if the head constructor of the scrutinee is not matched (with the same arity) by any clause.) However, we do assume that our terms are well-scoped – note that, in \(\text{let rec } (x_i = v_i)^j\) \(u\), the \((x_i)^j\) are in scope of \(u\) but also of all the \(v_i\).

4 Our inference rules for recursive definitions

4.1 Access/usage modes

Figure 2 defines the usage/access modes that we introduced in Section 2.1, their order structure, and the mode composition operations.

The modes are as follows:

\(^4\) In more expressive settings, the structure of usage modes does depend on the structure of values, and checks need to be presented as a refinement of a ML type system. We discuss this in our related work section (§7.1). Our modes are a degenerate case, a refinement of uni-typed ML.
Modes $\exists m ::= \text{Ignore} \\
| \text{Delay} \\
| \text{Guard} \\
| \text{Return} \\
| \text{Dereference}

Mode order: $\text{Ignore} \prec \text{Delay} \prec \text{Guard} \prec \text{Return} \prec \text{Dereference}$

Mode composition:

<table>
<thead>
<tr>
<th>$m[m']$</th>
<th>Ignore</th>
<th>Delay</th>
<th>Guard</th>
<th>Return</th>
<th>Dereference</th>
</tr>
</thead>
<tbody>
<tr>
<td>Ignore</td>
<td>Ignore</td>
<td>Ignore</td>
<td>Ignore</td>
<td>Ignore</td>
<td>Ignore</td>
</tr>
<tr>
<td>Delay</td>
<td>Ignore</td>
<td>Delay</td>
<td>Delay</td>
<td>Delay</td>
<td>Dereference</td>
</tr>
<tr>
<td>Guard</td>
<td>Ignore</td>
<td>Delay</td>
<td>Guard</td>
<td>Guard</td>
<td>Dereference</td>
</tr>
<tr>
<td>Return</td>
<td>Ignore</td>
<td>Delay</td>
<td>Guard</td>
<td>Return</td>
<td>Dereference</td>
</tr>
<tr>
<td>Dereference</td>
<td>Ignore</td>
<td>Dereference</td>
<td>Dereference</td>
<td>Dereference</td>
<td>Dereference</td>
</tr>
</tbody>
</table>

\textbf{Fig. 2.} Access/usage modes and operations

\textbf{Remark 1 (Discarding).} The Guard mode is also used for subterms whose result is discarded by the evaluation of their context. For example, the hole $\Box$ is in a Guard context in (let $x = \Box$ in $u$), if $x$ is never used in $u$; even if the hole value is not needed, call-by-value reduction will first evaluate it and discard it. When these subterms participate in a cyclic definition, they cannot create a self-loop, so we consider them guarded.

Our ordering $m \prec m'$ places less demanding, more permissive modes that do not involve dereferencing variables (and so permit their use in recursive definitions), below more demanding, less permissive modes.
Each mode is closely associated with particular expression contexts. For example, \( t \□ \) is a Dereference context, since the function \( t \) may access its argument in arbitrary ways, while \( \lambda x. □ \) is a Delay context.

Mode composition corresponds to context composition, in the sense that if an expression context \( E[□] \) uses its hole at mode \( m \), and a second expression context \( E'[□] \) uses its hole at mode \( m' \), then the composition of contexts \( E[E'[□]] \) uses its hole at mode \( m|m'| \). Like context composition, mode composition is associative, but not commutative — for instance, Dereference[Delay] is Dereference, but Delay[Dereference] is Delay.

Continuing the example above, the context \( t (\lambda x. □) \), formed by composing \( t □ \) and \( \lambda x. □ \), is a Dereference context: the intuition is that the function \( t \) may pass an argument to its input and then access the result in arbitrary ways. In contrast, the context \( \lambda x. (t □) \), formed by composing \( \lambda x. □ \) and \( t □ \), is a Delay context: the contents of the hole will not be touched before the abstraction is applied.

Finally, Ignore is the absorbing element of mode composition \( (m| Ignore) = Ignore = Ignore[m] \), and Return is the identity element \( (Return[m] = m = m|Return) \).

4.2 Inference rules

Environment notations Our environments \( \Gamma \) associate variables \( x \) to modes \( m \). We write \( \Gamma_1,\Gamma_2 \) for the union of two environments with disjoint domains, and \( \Gamma_1 + \Gamma_2 \) for the merge of two overlapping environments, taking the maximum mode for each variable. We sometimes use family notation for environments, writing \( (\Gamma_i)_i \) to indicate the disjoint union of the members, and \( \sum (\Gamma_i)_i \) for the non-disjoint merge of a family of environments.

Inference rules Figure 3 presents the inference rules for access/usage modes. The rules are composed into several different judgments, even though our simple core language makes it possible to merge them. In the full system for OCaml the decomposition is necessary to make the system manageable.

Variable and subsumption rules The variable rule is as one would expect: the usage mode of \( x \) in an \( m \)-context is \( m \). In this presentation, we let the rest of the environment \( \Gamma \) be arbitrary; we could also have imposed that it map all variables to Ignore. Our directed/algorithmic check returns the “least demanding” environment \( \Gamma \) for all satisfiable judgments, so it uses Ignore in any case.

We have a subtyping/subsumption rule: if we want to check \( t \) under the mode Guard, it is always correct to make our life harder and try to check it under the stronger mode Dereference. Our algorithmic check will never use this rule; it is here for completeness. The direction of the comparison may seem unusual. We can coerce a \( \Gamma \vdash t : m \) into \( \Gamma \vdash t : m' \) when \( m \gg m' \) holds, while we would expect \( m \leq m' \). This comes from the fact that our right-to-left reading is opposite to the usual reading direction of type judgments, and influenced our order definition. When \( m \gg m' \) holds, \( m \) is more demanding than \( m' \), which means (in the usual subtyping sense) that it classifies fewer terms.
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Term judgment \( \Gamma \vdash t : m \)

\[
\frac{\Gamma, x : m \vdash x : m}{\Gamma \vdash t : m \quad m \succ m'} \quad \frac{\Gamma \vdash t : m}{\Gamma \vdash t : m'}
\]

\[
\frac{\Gamma, x : m \vdash t : m}{\Gamma, x : m \vdash \lambda x. t : m \quad \Gamma \vdash \lambda x. t : m}
\]

\[
\frac{\sum (\Gamma_i)^i \vdash \text{Guard} \to (\Gamma_i)^i}{\sum (\Gamma_i)^i \vdash \text{Guard} \to (\Gamma_i)^i} \quad \frac{\Gamma_i \vdash t : m[\text{Dereference}]}{\Gamma_i \vdash \text{match} t \text{ with } h : m}
\]

\[
\frac{\sum (\Gamma_i)^i \vdash \text{Return} \to (\Gamma_i)^i}{\sum (\Gamma_i)^i \vdash \text{Guard} \to (\Gamma_i)^i} \quad \frac{\Gamma_i \vdash t_i : m}{\Gamma_i \vdash \text{Return} \to (\Gamma_i)^i}
\]

\[
\frac{\sum (\Gamma_i)^i \vdash \text{Guard} \to (\Gamma_i)^i}{\sum (\Gamma_i)^i \vdash \text{Guard} \to (\Gamma_i)^i}
\]

Fig. 3. Mode inference rules

Simple rules We have seen the \( \lambda x. t \) rule already, in Section 2.2. Since \( \lambda \) delays evaluation, checking it a usage context \( m \) involves checking the body \( t \) under the weaker mode \( m[\text{Delay}] \). The necessary constraints \( \Gamma \) are returned, after removing the constraint over the function variable.\(^5\)

The application rule checks both the function and its argument in a Dereference context, and merges the two resulting environments, taking the maximum (most demanding) mode on each side; for example, a variable \( y \) is dereferenced by \( t u \) if it is dereferenced by either \( t \) or \( u \).

The constructor rule is similar to the application rule, except that the constructor parameters appear in Guard context, rather than Dereference.

Pattern-matching The rule for \( \text{match} t \text{ with } h \) relies on a different clause judgment \( \Gamma \vdash^{\text{cl}} h : m \), check each clause in turn and merging their environments. On a single clause \( K (\Gamma_i)^i \to u \), we check the right-hand-side expressions \( u \) in

\(^5\) In situations where it is desirable to have a richer mode structure to analyze function applications, as considered by some of the related work (Section 7.1), we could use the mode \( m_x \) in a richer return mode \( m_x \to m \).
the ambient mode \( m \), and remove the pattern-bound variables \((x_i)^t\) from the environment.\(^6\)

**Recursive definitions** The rule for mutually-recursive definitions let \( \textbf{rec} \ b \) in \( u \) is split into two parts with disjoint responsibilities.

First, the binding judgment \((x_i : \Gamma_i)^t \vdash \textbf{rec} \ b\) is independent from the ambient context and usage mode, it checks recursive bindings in isolation in the Return mode, and relates each name \( x_i \) introduced by the binding \( b \) to an environment \( \Gamma_i \) on the ambient free variables. The recursively-bound variables \((x_i)^t\) do not appear in this environment, but we check that their modes \((m_i)^t\) are Guard or less demanding, to ensure that these mutually-recursive definitions are valid. This is the check we mentioned at the end of Section 2.2.

Second, the let \( \textbf{rec} \ b \) in \( u \) rule of the term judgment takes these \( \Gamma_i \) and uses them under a composition \( m'_i[\Gamma_i] \), to account for the actual usage mode of the variables. \((m[\Gamma])\) is the pointwise lifting of composition for each mode in \( \Gamma \). The usage mode \( m'_i \) is a combination of the usage mode in the body \( u \) and Guard, used to indicate that our call-by-value language will compute the values now, even if they are not used in \( u \), or only under a delay – see our Remark 1.

### 4.3 Properties

The following technical results can be established by simple inductions on typing derivations, without any reference to an operational semantics.

**Lemma 1.** \( \Gamma \vdash t : \text{Ignore} \) is provable with only Ignore modes in \( \Gamma \).

**Lemma 2.** \( \Gamma \vdash t : \text{Delay} \) holds exactly when \( \Gamma \) maps all free variables of \( t \) to Delay or Ignore.

**Lemma 3.** \( \Gamma \vdash t : \text{Dereference} \) holds exactly when \( \Gamma \) maps all free variables of \( t \) to Dereference.

**Lemma 4 (Environment flow).** If \( \Gamma \vdash t : m \) contains a subderivation \( \Gamma' \vdash t' : m' \), then \( \Gamma \succeq \Gamma' \).

**Lemma 5 (Weakening).** If \( \Gamma \vdash t : m \) holds then \( \Gamma + \Gamma' \vdash t : m \) also holds.

(Weakening would not be admissible if our variable rule imposed Ignore on the rest of the context.)

**Lemma 6 (Substitution).** If \( \Gamma, x : m_u \vdash t : m \) and \( \Gamma' \vdash u : m_u \) hold, then \( \Gamma + \Gamma' \vdash t[u/x] : m \) holds.

**Lemma 7 (Subsumption elimination).** Any derivation in the system can be rewritten so that the subsumption rule is only applied with the variable rule as premise.

\(^6\) If we wanted a finer-grained analysis of usage of the sub-components of our data, we would use the sub-modes \((m_i)^t\) of the pattern variables to refine/enrich/annotate the datatype of pattern scrutinee.
Theorem 1 (Principal environments). Whenever both $\Gamma_1 \vdash t : m$ and $\Gamma_2 \vdash t : m$ hold, then $\min(\Gamma_1, \Gamma_2) \vdash t : m$ also holds.

Proof. The proof first performs subsumption elimination on both derivations, and then by simultaneous induction on the results. The elimination phase makes proof syntax-directed, which guarantees that (on non-variables) the same rule is always used on both sides in each derivation. ☐

This results tells us that whenever $\Gamma \vdash t : m$ holds, then it holds for a minimal environment $\Gamma$ – the minimum of all satisfying $\Gamma$.

Definition 1 (Minimal environment). $\Gamma$ is minimal for $t : m$ if $\Gamma \vdash t : m$ and, for any $\Gamma' \vdash t : m$ we have $\Gamma \preceq \Gamma'$.

In fact, we can give a precise characterization of “minimal” derivations, that uniquely determines the output of our right-to-left algorithm.

Definition 2 (Minimal derivation). A derivation is minimal if it does not use the subsumption rule, and, in the conclusion $\Gamma \vdash x : m$ of each variable rule, $\Gamma$ is minimal for $x : m$.

Definition 3 (Minimization). Given a derivation $D :: \Gamma \vdash t : m$, we define the minimal derivation $\text{minimal}(D)$ by:

- performing subsumption-elimination to get another derivation of $\Gamma \vdash t : m$,
- replacing the context of each variable rule by the minimal context for this variable, which gives a minimal derivation of $\Gamma_m \vdash t : m$ with $\Gamma_m \preceq \Gamma$.

(We had to check, by direct induction, that changing variable rules in this way preserves typability, and does not require introducing subsumptions.)

Lemma 8 (Determinism). If $D_1 :: \Gamma_1 \vdash t : m$ and $D_2 :: \Gamma_2 \vdash t : m$, then $\text{minimal}(D_1)$ and $\text{minimal}(D_2)$ are the same minimal derivation.

Corollary 1. The environment $\Gamma$ of a derivation $\Gamma \vdash t : m$ is minimal for $t : m$ if and only if $\Gamma \vdash t : m$ admits a minimal derivation.

Proof. If $\Gamma \vdash t : m$ has a minimal environment, then the context $\Gamma_m \preceq \Gamma$ obtained by minimization must itself be $\Gamma$.

Conversely, if a derivation $D_m :: \Gamma \vdash t : m$ is minimal, then all other derivations $\Gamma' \vdash t : m$ have $D_m$ as minimal derivation by determinism (Lemma 8), so $\Gamma \preceq \Gamma'$ holds. ☐

Theorem 2 (Localization). $\Gamma \vdash t : m'$ implies $\text{max}(\Gamma) \vdash t : \text{max}(m', m'')$.

Furthermore, if $\Gamma$ is minimal for $t : m$, then $\text{max}(\Gamma)$ is minimal for $t : \text{max}(m', m'')$.

Proof. The proof proceeds by direct induction on the derivation, and does not change its structure: each rule application in the source derivation becomes the same source derivation in the result. In particular, minimality of derivations is preserved, and thus by Corollary 1 minimality of environments is preserved.

Besides associativity of mode composition, many cases rely on the fact that external mode composition preserves the mode order structure: $m_1' \prec m_2'$ implies $m[m_1'] \prec m[m_2']$, and $\text{max}(m[m_1'], m[m_2'])$ is $m[\text{max}(m_1', m_2')]$. ☐


5 Meta-theory: soundness

5.1 Operational semantics

\begin{align*}
\text{Values} \ni v &::= \lambda x. t \mid K (w_i)^i \\
\text{WeakValues} \ni w &::= x, y, z \mid v \\
\text{ValueBindings} \ni B &::= (x_i = v_i)^i \\
\text{EvalCtx} \ni E &::= \square \mid E[F]
\end{align*}

\[
(\lambda x. t) v \rightarrow^\text{bd} t[v/x] \\
(x = v) \in \text{frame} \quad \text{let rec } B \text{ in } \square
\]

\[
\forall (K' (x'_i)^i \rightarrow u') \in h, \ K \neq K' \\
\text{match } K (w_i)^i \text{ with } (h \mid K (x_i)^i \rightarrow u \mid h') \rightarrow^\text{bd} u[(w_i/x_i)^i]
\]

\[
(x = v) \in (b \cup b') \\
(x = v) \in \text{frame} \quad \text{let rec } b, y = \square, b' \text{ in } u
\]

\[
(x = v) \in \text{frame} \quad \text{let rec } b, y = \square, b' \text{ in } u \\
(x = v) \in E[F] \\
(x = v) \in E[F]
\]

\[
E[t] \rightarrow E[t'] \\
E[x] \rightarrow E[v]
\]

Fig. 4. Operational semantics

Figure 4 and the explanations below recall the operational semantics of Nordlander, Carlsson, and Gill (2008). (Unless explicitly noted, the content and ideas come from this work.)

Weak values. As we have seen, constructors in recursive definitions can be used to construct cyclic values. For example, the definition

\[
\text{let rec } x = \text{Cons} (\text{One} (\emptyset), x)
\]

is normal for this reduction semantics. The occurrence of the variable \(x\) inside the Cons cell corresponds to a backreference, the cell’s address in a cyclic in-memory representation.

This key property is achieved by defining a class of weak values, noted \(w\), to be either (strict) values or variables. Weak values occur in the definition of the semantics wherever a cyclic reference can be passed without having to dereference.

Several previous works (see Related Work Section 7.1) defined semantics where \(\beta\)-redexes were of the form (\(\lambda x. t\) \(w\)), to allow yet-unevaluated recursive definitions to be passed as function arguments. OCaml does not allow this (a function call requires a fully-evaluated argument), so our redexes are the traditional (\(\lambda x. t\) \(v\)). This is a difference from Nordlander, Carlsson, and Gill (2008). On the other hand, we do allow cyclic datatype values by only asking weak values under data constructors: besides closures (\(\lambda x. t\)), the other value form is \(K (w_i)^i\).
**Bindings in evaluation contexts** An evaluation context $E$ is a nested list of evaluation frames $F$ under which evaluation may happen. We chose a very under-constrained semantics (for example, $t\ u$ may perform reductions on either $t$ or $u$), as OCaml has unspecified evaluation order for applications and constructors, but making it deterministic would not change much.

One common aspect of most operational semantics for letrec, including our reference, is that $\text{let rec } B \text{ in } \square$ can be part of evaluation contexts, where $B$ represents a (recursive) “value binding”, an island of recursive definitions that have all been reduced to values. This is different from traditional source-level operational semantics of $\text{let } x = v \text{ in } u$, which is reduced to $u[v/x]$ before going further. This cannot be done for letrec blocks, as the value $v$ may itself refer to its name $x$, so instead “value bindings” remain in the context, in the style of explicit substitution calculi.

**Head reduction** Head redexes, the sources of the head-reduction relation $t \rightarrow^\text{hd} t'$, come from applying a $\lambda$-abstraction or from pattern-matching on a head constructor. Following ML semantics, pattern-matching is ordered: a matching clause is taken only if it is the first matching clause in order.

**Reduction** Reduction $t \rightarrow t'$ may happen under any evaluation context, and is of either of two forms. The first is completely standard: any redex $H[v]$ can be reduced under an evaluation context $E$.

The second rule reduces a variable $x$ in in an evaluation context $E$ by binding lookup: it is replaced by the value of the recursing binding $B$ in the context $E$ which defines it. This uses the auxiliary definition $(x = v)^\text{ctx} \in E$ to perform this lookup.

The lookup rule has worrying consequences for our rewriting relation: it makes it non-deterministic and non-terminating. Indeed, consider a weak value of the form $K(x)$, used in a pattern-matching $\text{match } K(x) \text{ with } h$ for example. It is possible to reduce the pattern-matching immediately, or to first lookup the value of $x$ and then reduce. Furthermore, it could be the case that $x$ is precisely defined by a cyclic binding $x = K(x)$. Then the lookup rule would reduce to $\text{match } K(K(x)) \text{ with } h$, and we could keep looking indefinitely. This is discussed in detail in Nordlander, Carlsson, and Gill (2008), who prove that the reduction is in fact confluent modulo unfolding. (Allowing these irritating but innocuous behaviors is a large part of what makes their semantics simpler than previous presentations.)

### 5.2 Failures

In this section, we are interested in formally defining dynamic failures. When can we say that a term is “wrong”? — in particular, when is a valid implementation of the operational semantics allowed to crash? This aspect is not discussed in detail in Nordlander, Carlsson, and Gill (2008), so we had to make our own definitions; we found it surprisingly subtle.
The first obvious sort of failure is a type mismatch between a value constructor and a value destructor: trying to pass an argument to a non-function, or pattern-matching on a function instead of a head constructor. These failures would be ruled out by a simple type system.

The difficult part is defining failures that occur when trying to access a recursively-defined variable too early. In the lookup reduction rule for a term $E[x]$, we look for the value of $x$ in a binding of the context $E$. This value may not exist (yet), and that may or may not represent a runtime failure.

We assume that bound names are all distinct from each other, so there may not be several $v$ values. The only binders that we reduce under are `letrec`, so $x$ must come from one; however, it is possible that $x$ be part of a `letrec` block currently being evaluated, with an evaluation context of the form $E[\text{let rec } x = t, E'u]$ for example, and that $x$’s binding has not been reduced to a value yet.

However, not all such cases are failures, in presence of data constructors allowing to build cyclic values. For example the term `let rec x = Pair(x,t) in x` can be decomposed into $E[x]$ to isolate the occurrence of $x$ as the first member of the pair. This occurrence of $x$ is in reducible position, but there is no $v$ such that $(x = v) \in_\text{ctx} E$, unless $t$ is already a weak value.

To characterize failures during recursive evaluation, we propose to restrict ourselves to forcing contexts, that we note $E_f$, which must really access the value of their hole. A variable in a forcing context that cannot be looked up in the context is a dynamic failure: we are forcing the value of a variable that has not yet been evaluated. If a term contains such a variable in lookup position, we call it a vicious term.

This choice of name comes from a close correspondence with what we called a vicious circle: if there exists an evaluation order for a mutually-recursive block that does reach a vicious term, then there is no circle. If any evaluation order, at some point, hits a vicious term (it requires something that is not yet evaluated), we have a vicious circle.

\[
\text{Mismatch} \overset{\text{def}}{=} \{ E[H[v]] \mid H[v] \rightarrow^\text{hd} \} \quad \text{Vicious} \overset{\text{def}}{=} \{ E[x] \mid \nexists v, (x = v) \in_\text{ctx} E \}
\]

**Fig. 5. Failure terms**

We define these failure terms precisely in Figure 5.

Mismatches are characterized by head frames, context fragments that would form a $\beta$-redex if they were plugged a value of the correct type. A term of the form $H[v]$ that is stuck for head-reduction is a constructor-destructor mismatch.

The definition of forcing contexts $E_f$ takes into account the fact that recursive value bindings remain, floating around, in the evaluation context. A forcing
frame $F_1$ is a context fragment that forces evaluation of its variable; it would be tempting to say that a forcing context is necessarily of the form $E[F_1]$, but for example $F_1[\text{let rec } B \text{ in } \Box]$ must also be considered a forcing context.

Note that, due to the flexibility we gave to the evaluation order, mismatches and vicious terms need not be stuck: they may have other reducible positions in their evaluation context. In fact, a vicious term failing on a variable $x$ may reduce to a non-vicious term if the binding of $x$ is reduced to a value.

### 5.3 Soundness

**Lemma 9 (Forcing-dereference).** If $\Gamma, x : m \vdash E[x] : \text{Return}$ is derivable, then $m$ is Dereference.

*Proof.* This is immediate for forcing frames $F_1$, and proved by direct induction for forcing contexts $E_1$: evaluation contexts do not contain any Ignore or Delay frames, and all other modes are absorbed by Dereference during composition. \qed

**Theorem 3 (Vicious).** $\emptyset \vdash t : \text{Return}$ never holds for $t \in \text{Vicious}$.

*Proof.* Given $\vdash t : \text{Return}$, let us assume that $t$ is $E[x]$ with no value binding for $x$ in $E$, and show that $E$ is not a forcing context.

We implicitly assume that all terms are well-scoped, so the absence of value binding means that $E[x]$ is of the form

$$E[x] = E_{\text{out}}[\text{let rec } b, x = E_{\text{in}}[x], b' \text{ in } u]$$

Given our letrec typing rule (see Figure 3), the typing derivation for $t$ contains a sub-derivation for $t_{\text{rec}}$ of the form

$$\forall i, m_i \leq \text{Guard} \quad (\Gamma_i, (x_i : m_i)^i \vdash t_i : \text{Return}^i)$$

In particular, the premise for $E_{\text{in}}[x]$ is of the form $\Gamma, x : m \vdash E_{\text{in}}[x] : \text{Return}$ with $x \leq \text{Guard}$, and in particular $x \neq \text{Dereference}$. By the Forcing-dereference Lemma 9, $E_{\text{in}}$ cannot be a forcing context, and in consequence $E$ is not forcing either. \qed

**Theorem 4 (Subject reduction).** If $\Gamma \vdash t : m$ and $t \rightarrow t'$ then $\Gamma \vdash t' : m$.

*Proof.* We reason by inversion on the typing derivation of redexes, first for head-reduction $t \rightarrow_{\text{hd}} t'$ and then for reduction $t \rightarrow t'$.

**Head reduction** We only show the head-reduction case for functions; pattern-matching is very similar. We have:

$$\Gamma_i, x : m_x \vdash t : m | \text{Dereference} | \text{Delay}$$

$$\Gamma_i \vdash \lambda x. t : m | \text{Dereference} \quad \Gamma_i \vdash v : m | \text{Dereference}$$

$$\Gamma_i + \Gamma_i \vdash (\lambda x. t) v : m$$

By associativity, $m | \text{Dereference} | \text{Delay}$ is the same as $m | \text{Dereference}$. 

By subsumption, $\Gamma, x : m_x \vdash t : m[\text{Dereference}]$ implies $\Gamma, x : m_x \vdash t : m$.

To conclude by using the Substitution Lemma 6, we must reconcile the mode of the argument $v : m[\text{Dereference}]$ with the (apparently arbitrary) mode $x : m_x$ of the variable. We reason by an inelegant case distinction.

- If $m[\text{Dereference}]$ is Dereference, then by Lemma 3 either $m_x$ is Dereference (problem solved) or $x$ does not occur in $t$ (no need for the substitution lemma).
- If $m[\text{Dereference}]$ is not Dereference, then $m$ must be Ignore or Delay. If it is 
  Ignore, we can trivially prove our goal (Lemma 1). If it is Delay, then $m_x$
  itself can be weakened (by strengthening the derivation of $t$) to be below Delay (Lemma 2).

**Reduction under context** Reducing a head-redex under context preserves typeability by the argument above. Let us consider the lookup case.

$$\begin{align*}
(x = v) & \in E \\
E[x] & \rightarrow E[v]
\end{align*}$$

By inspecting the $(x = v) \in E$ derivation, we find a value binding $B$ within $E$ with $x = v$, and a derivation of the form

$$\begin{align*}
(x_i : \Gamma_i)^i & \vdash \text{rec } B \quad (m'_i)^i \overset{\text{def}}{=} (\max(m_i, \text{Guard}))^i \quad \Gamma_u, (x_i : m_i)^i \vdash u : m \\
\sum & (m'_i[\Gamma_i])^i + \Gamma_u \vdash \text{let rec } B \text{ in } u : m
\end{align*}$$

$$\forall i, m_i \leq \text{Guard} \quad (\Gamma_i, (x_i : m_i)^i \vdash v_i : \text{Return})^i$$

$$(x_i : \Gamma_i)^i \vdash \text{rec } (x_i = v_i)^i$$

Let $\Gamma_x$ and $m_x$ be the environments and mods of $x$ in the $\text{letrec}$ derivation (the $m_i$ and $\Gamma_i$ for $i$ such that $x = x_i$).

The occurrence of $x$ in the hole of $E[\square]$ is typed (eventually by a variable rule) at some mode $m[\square]$. The declaration-side mode $m_x$ was built by collecting the usage modes of all occurrences of $x$ in the $\text{letrec}$ body $u$, which in particular contains the hole of $E$, so we have $m[\square] \leq m_x$ (by the Environment Flow Lemma 4).

The binding derivation gives us a proof $\Gamma_x, \Gamma' \vdash v : \text{Return}$ that the binding $x = v$ was correct at its definition site ($\Gamma'$ has the mutually-recursive variables), independently of the ambient context. By the Localization Theorem 2, we can compose this within $m[\square]$ to get a derivation $m[\square][\Gamma_x, \Gamma'] \vdash v : m[\square]$, that we wish to substitute into the hole of $E$.

Plugging this derivation in the hole of $E$ requires weakening the derivation of $u$ (the part of $E[\square]$ that is after the declaration of $x$) to add the environment $m[\square][\Gamma_x, \Gamma']$. Weakening is always possible (Lemma 5), but it may change the environment of the derivation, while we need to preserve the environment of $E[x]$. This works, because the environment at the declaration site of $x$ is stricter than $m[\square][\Gamma_x, \Gamma']$, so it weakens into itself: indeed, it contains $m'_i[\Gamma_x]$ with $m'_i \overset{\text{def}}{=} \max(m_i, \text{Guard})$, so:
A right-to-left type system for mutually-recursive value definitions

\[ m \sqsubseteq m_i \sqsubseteq m_i', \text{ so } m \sqsubseteq [I_x] \leq m_i'[I_x] \]

In \( I' \), all modes are below Guard (this is the correctness condition in the derivation judgment), so they are below \( m_i' \).

\( \square \)

**Corollary 2.** Return-typed programs cannot go vicious.

## 6 Extension to a full language

### 6.1 The size discipline

The OCaml compilation scheme (there are other ways to perform recursive declarations) proceeds by reserving heap blocks for the recursively-defined values, and using the addresses of these heap blocks (which will eventually contain the values) as dummy values: it adds them to the environment and computes the values accordingly. If no vicious term exists, the addresses are never dereferenced during evaluation, and we get “correct” values as result. Those correct values are then moved into the dummy space, so that the original addresses contain the correct result.

For this strategy to work, we need to know how much space to allocate for each value. Not all OCaml types have a uniform size; variants (sum types) may contain constructors with different arities, resulting in different in-memory size.

After checking that mutually-recursive definitions are meaningful using the rules we described, the OCaml compiler checks that it can realize them, by trying to infer a static size for each value. It then accepts to compile each declaration if either:

- it has a static size, or
- it doesn’t have a statically-known size, but its usage mode of mutually-recursive definitions is always Ignore

(The second category corresponds to detecting some values that are actually non-recursive and lifting them out. Non-recursive values often occur in standard programming practice, when it is more consistent to declare a whole block as a single letrec but only some elements are recursive.)

This static-size test may depend on lower-level aspects of compilation, or at least value representation choices. For example,

\[
\text{if } p \text{ then } (\text{fun } x \to x) \text{ else } (\text{fun } x \to \text{not } x)
\]

has a static size (both branches have the same size), but

\[
\text{if } p \text{ then } (\text{fun } x \to x + 1) \text{ else } (\text{fun } x \to x + \text{offset})
\]

does not: the second function depends on a free variable \text{offset}, so it will be allocated in a closure with an extra field. (While \text{not} is also a free variable, it is a statically-resolvable reference to a global name.)
6.2 Dynamic representation checks: float arrays

OCaml uses a dynamic representation check for its polymorphic arrays: when the values given at array-creation time are floating-point numbers, OCaml chooses a specialized, unboxed representation for this array.

Inspecting the representation of elements during array creation means that although array construction looks like a guarding context, it is often in fact a dereference (unless the type of array elements is statically known to be non-\texttt{float}, in which case the inspection is elided). The following program must be rejected, for example:

\begin{verbatim}
let rec x = (let u = [\|y\|] in 10.)
    and y = 1.
\end{verbatim}

6.3 Exceptions and first-class modules

In OCaml, exception declarations are generative: if a functor body contains an exception declaration then invoking the functor twice will declare two exceptions with incompatible representations, so that catching one of them will not interact with raising the other.

Exception generativity is implemented by allocating a memory cell at functor-evaluation time (in the representation of the resulting module); and including the address of this memory cell as an argument of the exception payload. In particular, creating an exception value \texttt{M.Exit 42} may dereference the module \texttt{M} where \texttt{Exit} is declared.

Combined with another OCaml feature, first-class modules, this generativity can lead to surprising incorrect recursive declarations, by declaring a module with an exception and using the exception in the same recursive block. For instance, the following program is unsound and rejected by our analysis:

\begin{verbatim}
module type T = sig exception A of int end

let rec x = (let module M = (val m) in M.A 42)
    and (m : (module T)) = (module (struct exception A of int end) : T)
\end{verbatim}

6.4 GADTs

The original syntactic criterion for OCaml was implemented not directly on surface syntax, but on an intermediate representation quite late in the compiler pipeline (after typing, type-erasure, and some desugaring and simplifications). In particular, at the point where the check took place, exhaustive single-clause matches such as \texttt{match t with x -> ...} or \texttt{match t with () -> ...} had been transformed into direct substitutions.

This design choice led to programs of the following form being accepted:

\begin{verbatim}
type t = Foo
let rec x = (match x with Foo -> Foo)
\end{verbatim}
While this seems entirely innocuous, it becomes unsound with the addition of GADTs to the language:

```ocaml
type (_, _) eq = Refl : ('a, 'a) eq
let universal_cast (type a) (type b) : (a, b) eq =
    let rec (p : (a, b) eq) = match p with Refl -> Refl in
    p
```

For the GADT `eq`, matching against `Refl` is not a no-op: it brings a type equality into scope that increases the number of types that can be assigned to the program. It is therefore necessary to treat matches involving GADTs as inspections to ensure that a value of the appropriate type is actually available; without that change definitions such as `universal_cast` violate type safety.

### 6.5 Laziness

OCaml’s evaluation is eager by default, but it supports an explicit form of lazy evaluation: the programmer can write `lazy e` and `force e` to delay and force the evaluation of an expression.

The OCaml implementation performs a number of optimizations involving `lazy`. For example, when the argument of `lazy` is a trivial syntactic value (variable or constant), since eager and lazy evaluation usually behave equivalently, the compiler picks eager evaluation as an optimization to avoid thunk allocation.

However, for recursive definitions eager and lazy evaluation are not equivalent, and so the recursion check must treat `lazy trivialvalue` as if the `lazy` were not there. For example, the following recursive definition is disallowed, since the optimization described above nullifies the delaying effect of the `lazy`

```ocaml
let rec x = lazy y and y = ...
```

while the following definition is allowed by the check, since the argument to `lazy` is not sufficiently trivial to be subject to the optimization:

```ocaml
let rec x = lazy (y+0) and y = ...
```

Our typing rule for `lazy` takes this into account: “trivial” thunks are checked in mode `Return` rather than `Delay`.

### 7 Conclusion

We have presented a new static analysis for recursive value declarations, designed to solve a fragility issue in the OCaml language semantics and implementation. It is less expressive than previous works that analyse function calls in a fine-grained way; in return, it remains fairly simple, despite its ability to scale to a fully-fledged programming language, and the constraint of having a direct correspondence with a simple inference algorithm.

We believe that this static analysis may be of use for other functional programming languages, both typed and untyped. It seems likely that the techniques...
we have used in this work will apply to other systems — type parameter variance, type constructor roles, and so on. Our hope in carefully describing our system is that we will eventually see a pattern emerge for the design and structure of “things that look like type systems” in this way.

7.1 Related work

Degrees Boudol (2001) introduces the notion of “degree” $\alpha \in \{0, 1\}$ to statically analyze recursion in object-oriented programs (recursive objects, lambda-terms). Degrees refine a standard ML-style type system for programs, with a judgment of the form $\Gamma \vdash t : \tau$ where $\tau$ is a type and $\Gamma$ gives both a type and a degree for each variable. A context variable has degree 0 if it is required to evaluate the term (related to our Dereference), and 1 if it is not required (related to our Delay). Finally, function types are refined with a degree on their argument: a function of type $\tau^0 \rightarrow \tau'$ accesses its argument to return a result, while a $\tau^1 \rightarrow \tau'$ function does not use its argument right away, for example a curried function $\lambda x. \lambda y. (x, y)$ – whose argument is used under a delay in its body $\lambda y. (x, y)$. Boudol uses this reasoning to accept a definition such as

\[ \text{let rec obj = class\_constructor obj params,} \]

arising from object-oriented encodings, where $\text{class\_constructor}$ has a type $\tau^0 \rightarrow \ldots$.

Our system of mode is finer-grained than the binary degrees of Boudol; in particular, we need to distinguish Dereference and Guard to allow cyclic data structure constructions.

On the other hand, we do not reason about the use of function arguments at all, so our system is much more coarse-grained in this respect. In fact, refining our system to accept $\text{let rec obj = constr obj params}$ would be incorrect for our use-case in the OCaml compiler, whose compilation scheme would not allow passing yet-uninitialized data to a function.

In a general design aiming for maximal expressivity, access modes should refine ML types; in Boudol’s system, degrees are interlinked with the type structure in function types $\tau^\alpha \rightarrow \tau'$, but one could also consider pair types of the form $\tau_1^{\alpha_1} \times \tau_2^{\alpha_2}$, etc. In our simpler system, there are no interaction between value shapes (types) and access modes, so we can forget about types completely, a nice conceptual simplification. Our formalization will be done entirely in an untyped fragment of ML.

Compilation Hirschowitz, Leroy, and Wells (2003, 2009) discuss the space of compilation schemes for recursive value definitions, and prove the correctness of a compilation scheme similar to one used by the OCaml compiler, using in-place update to tie the knot after recursive bindings are evaluated. Their source language has $\text{letrec}$ bindings and a source-level operational semantics, based on floating bindings upwards in the term (similar to explicit substitutions or local thunk stores). Their target language can talk about uninitialized memory cells and their update, and a mutable-store operational semantics.
In the present work, we do not formalize a compilation scheme for recursive definitions, we only prove our static analysis correct with respect to a source-level operational semantics.

While they are presenting a lambda-calculus, these works were concerned with recursive modules and mixin modules in ML languages – as other related work to follow. Recursive modules are used when programming at large, where programmers are willing to introduce cyclic dependencies in subtle, non-local ways, which requires fine-grained checks.

We only consider term-level cyclic value definitions, a simpler problem domain where less static sophistication is demanded. In fact, we do not aim at accepting substantially more recursive definitions than the previous OCaml syntactic check, only to be more trustworthy.

Name access as an effect Dreyer (2004) proposes to track usage of recursively-defined variables as an effect, and designs a type-and-effect system whose effects annotations are sets of abstract names, maintained in one-to-one correspondence with letrec-bound variables. The construction letrec \( X \triangleright x : \tau = e \) introduces the abstract type-level name \( X \) corresponding to the recursive variable \( x \). This recursive variable is made available in the scope of the right-hand-side \( e : \tau \) at the type \( \text{box}(X,\tau) \) instead of \( \tau \) (reminding us of guardedness modalities). Any dereference of \( x \) must explicitly “unbox” it, adding the name \( X \) to the ambient effect.

This system is very powerful, but we view it as a core language rather than a surface language: encoding a specific usage pattern may require changing the types of the components involved, to introduce explicit box modalities:

- When one defines a new function from \( \tau \) to \( \tau' \), one needs to think about whether it may be later used with still-undefined recursive names as argument – assuming it indeed makes delayed uses of its argument. In that case, one should use the usage-polymorphic type function type \( \forall X.\text{box}(X,\tau) \rightarrow \tau' \) instead of the simple function type \( \tau \rightarrow \tau' \). (It is possible to inject \( \tau \) into \( \text{box}(X,\tau) \), so this does not restrict non-recursive callers.)
- One could represent cyclic data such as let rec ones = 1 :: ones in this system, but it would require a non-modular change of the type of the list-cell constructor from \( \forall \alpha.\alpha \rightarrow \text{List}(\alpha) \rightarrow \text{List}(\alpha) \) to the box-expecting type \( \forall \alpha.\alpha \rightarrow \forall X.\text{box}(X,\text{List}(\alpha)) \rightarrow \text{List}(\alpha) \).

In particular, one cannot directly use typability in this system as a static analysis for a source language; this work needs to be complemented by a static analysis such as ours, or the safety-analysis argument has to be performed manually by the user placing box annotations and operations.

Graph typing Hirschowitz also collaborated on static analyses for recursive definitions in Hirschowitz and Lenglet (2005); Bardou (2005). The design goal was to have a simpler system than existing work aiming for expressivity, and in particular make inference as simple as possible.
As a generalization of Boudol’s binary degrees they use compactified numbers \( \mathbb{N} \cup \{-\infty, \infty\} \). The degree of a free variable “counts” the number of subsequent \( \lambda \)-abstractions that have to be traversed before the variable is used; \( x \) has degree 2 in \( \lambda y. \lambda z. x \). A \( -\infty \) is never safe, it corresponds to our \textit{Dereference} mode. 0 conflates our \textit{Guard} and \textit{Return} mode (an ad-hoc syntactic restriction on right-hand-sides is used to prevent under-determined definitions), the \( n+1 \) are fine-grained representations of our \textit{Delay} mode, and finally \( +\infty \) is our \textit{Ignore} mode.

Another salient aspect of their system is the use of “graphs” in the typing judgment: a use of \( y \) within a definition \texttt{let x = e} is represented as an edge from \( y \) to \( x \) (labelled by the usage degree), in a constraint graph accumulated in the typing judgment. The correctness criterion is formulated in terms of the transitive closure of the graph: if \( x \) is later used somewhere, its usage implies that \( y \) also needs to be initialized in this context.

Our work does not need such a transitive-computation device, as our \texttt{let} rule uses a simple form of mode substitution to propagate usage information as expected. One contribution of our work is thus to show that the graph representation can be replaced by a more standard syntactic approach.

Finally, their static analysis mentions the in-memory size of value, which needs to be known statically, in the OCaml compilation scheme, to create uninitialized memory blocks for the recursive names before evaluating the recursive definitions. Our type system does not mention size at all, it is complemented by an independent (and simpler) analysis of static-size deduction, which is outside the scope of the present formalization; we describe it briefly in Section 6.1.

\( \texttt{F\# Syme (2006)} \) proposes a simple translation of mutually-recursive definitions into lazy/force constructions. For example, \texttt{let rec x = t and y = u} is turned into

\[
\begin{align*}
\texttt{let rec} \ x\texttt{thunk} &= \texttt{lazy (t[force x\texttt{thunk}/x, force y\texttt{thunk}/y])} \\
\texttt{and} \ y\texttt{thunk} &= \texttt{lazy (u[force x\texttt{thunk}/x, force y\texttt{thunk}/y])} \\
\texttt{let} \ x &= \texttt{force x\texttt{thunk}} \\
\texttt{let} \ y &= \texttt{force y\texttt{thunk}}
\end{align*}
\]

With this semantics, evaluation happens on-demand, which the recursive definitions evaluated at the time where they are first accessed. This implementation is very simple, but it turns vicious definitions into dynamic failures – handled by the lazy runtime which safely raises an exception. However, this elaboration cannot support cyclic data structures: \texttt{let rec ones = 1 :: ones} gets translated into \texttt{let rec ones\_thunk = lazy (1 :: force ones\_thunk)}, which fails immediately when first forced.

Nowadays, the \texttt{F\#} language reference (\textit{Syme, 2012}), provides an ad-hoc syntactic criterion, the “Recursive Safety Analysis”, roughly similar to the previous OCaml syntactic criterion, for safe cyclic value definitions, which do not use the thunk-introducing translation. (A set of mutually-recursive definitions can contain “safe” and “unsafe” definitions, and only the latter get thunked).

Finally, the implementation also performs a static analysis to detect some definitions that are bound to fail – it over-approximates safety by considering
ignoring occurrences within function abstractions, objects or lazy thunks, even if those delaying terms may themselves be called/accessed/forced at definition time. We believe that we could recover a similar analysis by changing our typing rules for our constructions – but with the OCaml compilation scheme we must absolutely remain sound.

**Operational semantics** Hirschowitz, Leroy, and Wells (2003, 2009) give operational semantics for a source-level language (floating `letrec` bindings) and a small-step semantics for their compilation-target language with mutable stores. Boudol and Zimmer (2002) and Dreyer (2004) use an abstract machine. Syme (2006) translates recursive definitions into lazy constructions, so the usual thunk-store semantics of laziness can be used to interpret recursive definitions. Finally, Nordlander, Carlsson, and Gill (2008) gives the simplest presentation of a source-level semantics we know of; we extend it with algebraic datatypes and pattern-matching, and use it as a reference to prove the soundness of our analysis.

One inessential detail in which the semantics often differ is the evaluation order of mutually-recursive right-hand-sides. Many presentations enforce an arbitrary (e.g. left-to-right) evaluation order. Some systems (Syme, 2006; Nordlander, Carlsson, and Gill, 2008) allow a reduction to block on a variable whose definition is not yet evaluated, and go evaluate it in turn; this provides the “best possible order” for the user. Another interesting variant would be to say that the reduction order is unspecified, and that trying to evaluate an uninitialized is always a fatal error / stuck term; this provides the “worst possible order”, failing as much as possible; as far as we know, the previous work did not propose it, although it is a simple presentation change. Most static analyses are evaluation-order-independent, so they are sound and complete with respect to the “worst order” interpretation.
Bibliography