Sound Invariant Checking Using Type Modifiers and Object Capabilities.

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Abstract

In this paper we use pre existing language support for type modifiers and object capabilities to enable a system for sound runtime verification of invariants. Our system guarantees that class invariants hold for all objects involved in execution. Invariants are specified simply as methods whose execution is statically guaranteed to be deterministic and not access any externally mutable state. We automatically call such invariant methods only when objects are created or the state they refer to may have been mutated. Our design restricts the range of expressible invariants but improves upon the usability and performance of our system compared to prior work. In addition, we soundly support mutation, dynamic dispatch, exceptions, and non determinism, while requiring only a modest amount of annotation.

We present a case study showing that our system requires a lower annotation burden compared to Spec#, and performs orders of magnitude less runtime invariant checks compared to the widely used ‘visible state semantics’ protocols of D, Eiffel. We also formalise our approach and prove that such pre existing type modifier and object capability support is sufficient to ensure its soundness.

2012 ACM Subject Classification  Theory of computation → Invariants, Theory of computation → Program verification, Software and its engineering → Object oriented languages

Keywords and phrases  type modifiers, object capabilities, runtime verification, class invariants

Digital Object Identifier 10.4230/LIPIcs.CVIT.2016.23

1 Introduction

Object oriented programming languages provide great flexibility through subtyping and dynamic dispatch: they allow code to be adapted and specialised to behave differently in different contexts. However this flexibility hampers code reasoning, since object behaviour is usually nearly completely unrestricted. This is further complicated with the support OO languages typically have for exceptions, memory mutation, and I/O.

Class invariants are an important concept when reasoning about software correctness. They can be presented as documentation, checked as part of static verification, or, as we do in this paper, monitored for violations using runtime verification. In our system, a class specifies its invariant by defining a boolean method called invariant. We say that an object’s invariant holds when its invariant method would return true. We do this, like Dafny [44], to minimise special syntactic and type-system treatment of invariants, making
them easier to understand for users. Whereas most other approaches treat invariants as a special annotation with its own syntax.

An invariant protocol [72] specifies when invariants need to be checked, and when they can be assumed; if such checks guarantee said assumptions, the protocol is sound. The two main sound invariant protocols present in literature are visible state semantic [51] and the Boogie/Pack-Unpack methodology [5]. The visible state semantics expect the invariants of receivers to hold before and after every public method call, and after constructors. Invariants are simply checked at all such points, thus this approach is obviously sound; however this can be incredibly inefficient, even in simple cases. In contrast, the pack/unpack methodology marks all objects as either packed or unpacked, where a packed object is one whose invariant is expected to hold. In this approach, an object’s invariant is checked only by the pack operation. In order for this to be sound, some form of aliasing and/or mutation control is necessary. For example, Spec#, which follows the pack/unpack methodology, uses a theorem prover, together with source code annotations. While Spec# can be used for full static verification, it conveniently allows invariant checks to be performed at runtime, whilst statically verifying aliasing, purity and other similar standard properties. This allows us to closely compare our approach with Spec#.

Instead of using automated theorem proving, it is becoming more popular to verify aliasing and immutability using a type system. For example, three languages: L42 [68, 67, 43, 35], Pony [22, 23], and the language of Gordon et. al. [38] use Type Modifiers (TMs) and Object Capabilities (OCs) to ensure safe and deterministic parallelism. While studying those languages, we discovered an elegant way to enforce invariants.

We use the guarantees provided by these systems to ensure that at that all times, if an object is usable in execution, its invariant holds. What this means is that if you can do anything with an object, such as by using it as an argument/receiver of a method call, we know that the invariant of it, and all objects reachable from it, holds. In order to achieve this, we use TMs and OCs to restrict how the result of invariant methods may change, this is done by restricting I/O as well as what state the invariant can refer to and what can alias/mutate such state. We use these restrictions to reason as to when an object’s invariant could have been violated, and when such object can next be used, we then inject a runtime check between these two points. See Section 3 for the exact details of our invariant protocol.

**Example**

Here we show an example illustrating our system in action. Suppose we have a Cage class which contains a Hamster; the Cage will move its Hamster along a path. We would like to ensure that the Hamster does not deviate from the path. We can express this as the invariant of Cage: the position of the Cage’s Hamster must be within the path (stored as a field of Cage).

```java
class Point {
    Double x; Double y;
    Point(Double x, Double y) {...}
    @Override read method Bool equals (read Object that) {
        return that instanceof Point &&
        this.x == ((Point)that).x && this.y == ((Point)that).y;
    }
}

class Hamster {Point pos; //pos is imm by default
    Hamster(Point pos) {...}
```

1 TMs are called reference capabilities in other works. We use the term TM here to not confuse them with object capabilities, another technique we also use in this paper.
class Cage {
    capsule Hamster h;
    List<Point> path; // path is imm by default
    Cage(capsule Hamster h, List<Point> path) {..}
    read method Bool invariant() {
        return this.path.contains(this.h.pos);
    }
    mut method Void move() {
        Int index = 1 + this.path.indexOf(this.h.pos);
        this.moveTo(this.path.get(index % this.path.size()));
    }
    mut method Void moveTo(Point p) { this.h.pos = p; }
}

Many verification approaches take advantage of the separation between primitive/value
types and objects, since the former are immutable and do not support reference equality.
However, our approach works in a pure OO setting without such a distinction. Hence we
write all type names in BoldTitleCase to underline this. Note: to save space, here and in
the rest of the paper we omit the bodies of constructors that simply initialise fields with
the values of constructor parameters, but we show their signature in order to show any
annotations.

We use the read annotation on equals to express that it does not modify either the
receiver or the parameter. In Cage we use the capsule annotation to ensure that the
Hamster’s reachable object graph (ROG) is fully under the control of the containing Cage.
We annotated the move and moveTo methods with mut, since they modify their receivers
ROG. The default annotation is always imm, thus Cage’s path field is a deeply immutable
list of Points. Our system performs runtime checks for the invariant at the end of Cage’s
constructor, moveTo method, and after any update to one of its fields. The moveTo method
is the only one that may (directly) break the Cage’s invariant. However, there is only a
single occurrence of this and it is used to read the h field. We use the guarantees of TMs
to ensure that no alias to this could be reachable from either h or the immutable Point
parameter. Thus, the potentially broken this object is not visible while the Hamster’s
position is updated. The invariant is checked at the end of the moveTo method, just before
this would become visible again. This technique loosely corresponds to an implicit pack
and unpack: we use this only to read the field value, then we work on its value while the
invariant of this is not known to hold, finally we check the invariant before allowing the
object to be used again.

Note: since only Cage has an invariant, only Cage has special restrictions, allowing the
code for Point and Hamster to be unremarkable. This is not the case in Spec#: all code
involved in verification needs to be designed with verification in mind \[7\].

**Spec# Example**
Here we show the previous example in Spec#, the system most similar to ours (see appendix
\[B\] for a more detailed discussion about this solution):

```
// Note: assume everything is ‘public’
class Point { double x; double y; Point(double x, double y) {..}
    [Pure] bool Equal(double x, double y) {
        return x == this.x && y == this.y; }
}
class Hamster{[Peer]Point pos;
```
Hamster([Captured] Point pos){..}
}
class Cage {
    [Rep] Hamster h; [Rep, ElementsRep] List<Point> path;

    Cage([Captured] Hamster h, [Captured] List<Point> path)
        requires Owner.Same(Owner.ElementProxy(path), path); {
            this.h = h; this.path = path; base(); }

    invariant exists {int i in (0 : this.path.Count);
        this.path[i].Equal(this.h.pos.x, this.h.pos.y) ;}
    void Move() {
        int i = 0;
        while(i<path.Count && !path[i].Equal(h.pos.x,h.pos.y)){i++;}
        expose(this) {this.h.pos = this.path[i%this.path.Count];}}
}

In both versions, we designed Point and Hamster in a general way, and not solely to be
used by classes with an invariant, in particular Point is not an immutable class. However,
doing this in Spec# proved difficult, in particular we were unable to override Object.Equals,
or even define a usable equals method that takes a Point, as such we could not call either
List<Point>.Contains or List<Point>.IndexOf.

Even with all of the above annotations, we still needed special care in creating Cages:

List<Point> pl = new List<Point>{new Point(0,0),new Point(0,1)};
Owner.AssignSame(pl, Owner.ElementProxy(pl));
Cage c = new Cage(new Hamster(new Point(0,0)), pl);

Whereas with our system we can simply write:

List<Point> pl = List.of(new Point(0,0), new Point(0,1));
Cage c = new Cage(new Hamster(new Point(0,0)), pl);

In Spec# we had to add 10 different annotations, of 8 different kinds; some of which
were quite involved. In comparison, our approach requires only 7 simple keywords, of 3
different kinds; however we needed to write a separate moveTo method, since we do not want
to burden our language with extra constructs such as Spec#’s expose.

Summary

We have fully implemented our protocol in L42 and we used this implementation to implement
and test an interactive GUI involving a class with an invariant. On a test case with 5 objects
with an invariant, our protocol performed only 77 invariant checks, whereas the visible state
semantic invariant protocols of D and Eiffel perform 53 and 14 million checks (respectively).
See Section 7 for an explanation of these result. We also compared with Spec#, whose
invariant protocol performs the same number of checks as ours, however the annotation
burden was almost 4 times higher than ours.

In this paper we argue that our protocol is not only more succinct than the pack/unpack
approach, but is also easier and safer to use. Moreover, our approach deals with more
scenarios than most prior work: we allow sound catching of invariant failures and also

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2 A suitably anonymised, experimental version of L42, supporting the protocol described in this paper,
   together with the full code of our case studies, is available at [http://l42.is/EcoopArtifact.zip](http://l42.is/EcoopArtifact.zip).
3 We also believe it would be easy to implement our protocol in Pony and Gordon et al.’s language.
carefully handle non deterministic operations like I/O. Section 2 explains the type modifier and object capability support we use for this work. Section 3 explains the details of our invariant protocol, and section 4 formalises a language enforcing this protocol. Sections 5 and 6, respectively, explain and motivate how our protocol can handle invariants over immutable and encapsulated data. Section 7 presents our GUI case study and compares it against visible state semantics and Spec#. Sections 8 and 9 provide related work and conclusions.

Appendix A provides a proof that our invariant protocol is sound. Appendices B and C provide further case studies and comparisons against Spec#, D and Eiffel.

2 Type Modifiers and Object Capabilities

Reasoning about imperative object oriented (OO) programs is a non trivial task, made particularly difficult by mutation, aliasing, dynamic dispatch, I/O, and exceptions. There are many ways to perform such reasoning, here we use the type system to restrict, but not prevent such behaviour in order to be able to soundly enforce invariants with runtime verification (RV).

Type Modifiers (TMs)

TMs, as used in this paper, are a type system feature that allows reasoning about aliasing and mutation. Recently a new design for them has emerged that radically improves their usability; three different research languages are being independently developed relying on this new design: the language of Gordon et. al. [35], Pony [22, 23], and L42 [68, 67, 43, 35]. These projects are quite large: several million lines of code are written in Gordon et. al.’s language and are used by a large private Microsoft project; Pony and L42 have large libraries and are active open source projects. In particular the TMs of these languages are used to provide automatic and correct parallelism [35, 22, 23, 67].

While we focus on the specific TMs provided by L42, Pony, and Gordon et. al., type modifiers are a well known language mechanism [75, 11, 61, 22, 35, 38] that allow statically reasoning about mutability and aliasing properties of objects. With slightly different names and semantics, the four most common modifiers for references to objects are:

- Mutable (mut): the referenced object can be mutated, as in most imperative languages without modifiers. If all types are mut, there is no restriction on aliasing/mutation.
- Readonly (read): the referenced object cannot be mutated by such references, but there may be mutable aliases to such object, thus mutation can still be observed.
- Immutable (imm): the referenced object can never mutate. Like read references, one cannot mutate through an imm reference, however imm references also guarantee that the referenced object will not mutate through any other alias.
- Encapsulated (capsule): everything in the reachable object graph (ROG) of a capsule reference (including itself) is mutable only through that reference; however immutable references can be freely shared across capsule boundaries.

TMs are different to field or variable modifiers like Java’s final: TMs apply to references, whereas final applies to fields themselves. Unlike a variable/field of a read type, a final variable/field cannot be reassigned, it always refers to the same object, however the variable/field can still be used to mutate the referenced object. On the other hand, an object cannot be mutated through a read reference, however a read variable can still be reassigned.

\[^4\] In C, this is similar to the difference between A* const (like final) and const A* (like read), where const A* const is like final read.
Consider the following example usage of \texttt{mut}, \texttt{imm}, and \texttt{read}, where we can observe a change in \texttt{rp} caused by a mutation inside \texttt{mp}.

\begin{verbatim}
mut Point mp = new Point(1, 2);
mp.x = 3; // ok
imm Point ip = new Point(1, 2);
// ip.x = 3; // type error
read Point rp = mp; // ok, read is a common supertype of imm/mut
// rp.x = 3; // type error
mp.x = 5; // ok, now we can observe rp.x == 5
ip = new Point(3, 5); // ok, ip is not final
\end{verbatim}

There are several possible interpretations of the semantics of type modifiers. Here we assume the full/deep meaning [79, 66]:
- the objects in the ROG of an immutable object are immutable,
- a mutable field accessed from a \texttt{read} reference produces a \texttt{read} reference,
- no casting/promotion from \texttt{read} to \texttt{mut} is allowed.

There are many different existing techniques and type systems that handle the modifiers above [79, 21, 39, 38, 68]. The main progress in the last few years is with the flexibility of such type systems: where the programmer should use \texttt{imm} when representing immutable data and \texttt{mut} nearly everywhere else. The system will be able to transparently promote/recover [38, 22, 68] the type modifiers, adapting them to their use context. To see a glimpse of this flexibility, consider the following example:

\begin{verbatim}
mut Circle mc = new Circle(new Point(0, 0), 7);
capsule Circle cc = new Circle(new Point(0, 0), 7);
imm Circle ic = new Circle(new Point(0, 0), 7);
\end{verbatim}

Here \texttt{mc}, \texttt{cc}, and \texttt{ic} are syntactically initialised with the same expression: \texttt{new Circle(..)}. The \texttt{new} expression returns a \texttt{mut}, so \texttt{mc} is obviously ok. Moreover, the expression does not use any \texttt{mut} local variables, thus the flexible TM system allows the \texttt{mut} result to be promoted to \texttt{capsule}, thus \texttt{cc} is ok. Additionally, a \texttt{capsule} can be implicitly converted to \texttt{imm}, thus \texttt{ic} is also ok. We want to emphasise that this is not a special feature of \texttt{new} expressions: any expression of a \texttt{mut} type that uses no free \texttt{mut} variables declared outside can be implicitly promoted to \texttt{capsule/imm} \footnote{This requires some restrictions on \texttt{read} fields not discussed in detail for lack of space.}. This is the main improvement on the flexibility of TMs in recent literature [67, 68, 38, 22, 23]. Former work [16, 14, 40, 69, 2], which eventually enabled the work of Gordon et. al., does not consider promotion and infers uniqueness/isolation/immutability only when starting from references that have been tracked with restrictive annotations along their whole lifetime. From a usability perspective, this improvement means that these TMs are opt-in: a programmer can write large sections of code mindlessly using \texttt{mut} types and be free to have rampant aliasing. Then, at a later stage, another programmer may still be able to encapsulate those data structures into an \texttt{imm} or \texttt{capsule} reference.

The \texttt{capsule} modifier (sometimes called isolated/\texttt{iso}) is possibly the one whose details differ the most in literature. Here we refer to the interpretation of [38], that introduced the concept of recovery/promotion. This concept is the basis for L42, Pony, and Gordon et. al.’s type systems [38, 67, 68, 67, 22, 23].
The capsule/isolated fields of Gordon et. al. and Pony rely on destructive reads. In order to read them, a new value (such as null) will be assigned to them. In contrast, L42 does not require such destructive reads, thus capsule fields can be accessed many times, and their content can be seen from outside; but only in controlled ways. Both Gordon et. al. and Pony restrict how capsule local variables can be used by changing the type they are seen as, however both allow the local variable to be ‘consumed’, allowing them to be used as normal capsule/isolated expressions, at the cost of being unable to use the variable again. L42 however uses a simpler approach where all accesses to capsule local variables consume them: they are expressed using linear/affine types, thus they can only be used once.

**Exceptions**

In most languages exceptions may be thrown at any point; combined with mutation this complicates reasoning about the state of programs after exceptions are caught: if an exception was thrown whilst mutating an object, what state is that object in? Does its invariant hold? The concept of strong exception safety (SES) simplifies reasoning: if a try-catch block caught an exception, the state visible before execution of the try block is unchanged, and the exception object does not expose any object that was being mutated. L42 already enforces SES for unchecked exceptions. L42 enforces SES using TMs in the following way:

- Code inside a try block capturing unchecked exceptions is typed as if all mut variables declared outside of the block are read.
- Only imm objects may be thrown as unchecked exceptions.

This strategy does not restrict throwing exceptions, but only catching unchecked ones. SES allows us to soundly capture invariant failures as unchecked exceptions: the broken object is guaranteed to be garbage collectable when the exception is captured. For the purposes of soundly catching invariant failures, it would be sufficient to enforce SES only when capturing exceptions caused by such failures.

**Object Capabilities (OCs)**

OCs, which L42, Pony, and Gordon et. al.’s work have, are a widely used programming style that allows associating resources with objects. When this style is respected, code that does not possess an alias to such an object cannot use its associated resource. Here, as in Gordon et. al.’s work, we use OCs to reason about determinism and I/O. To properly enforce this, the OC style needs to be respected while implementing the primitives of the standard library and when performing foreign function calls that could be non deterministic, such as operations that read from files or generate random numbers. Such operations would not be provided by static methods, but instead instance methods of classes whose instantiation is kept under control.

For example, in Java, System.in is a capability object that provides access to the standard input resource, however, as it is globally accessible it completely prevents reasoning about determinism.

In contrast, if Java were to respect the object capability style, the main method could take a System parameter, as in main(mut System s) {.. s.in.read() ..}. Calling methods

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6 This is needed to support safe parallelism. Pony takes a more drastic approach and does not support exceptions in the first place. We are not aware of how Gordon et. al. handles exceptions, however in order for it to have sound unobservable parallelism it must have some restrictions.

7 Transactions are another way of enforcing strong exception safety, but they require specialized and costly run time support.

8 A formal proof of why these restriction are sufficient is presented in the work of Lagorio.
on that System instance would be the only way to perform I/O; moreover, the only System instance would be the one created by the runtime system before calling main. This design has been explored by Joe-E [32]. OCs are typically not part of the type system nor do they require runtime checks or special support beyond that provided by a memory safe language. However, since L42 allows user code to perform foreign calls without going through a predefined standard library, its type system enforces the OC pattern over such calls:

- Foreign methods (which have not been whitelisted as deterministic) and methods whose names start with #$ are capability methods.
- Constructors of classes declared as capability classes are also capability methods.
- Capability methods can only be called by other capability-methods or mut/capsule methods of capability classes.
- In L42 there is no main method, rather it has several main expressions; such expressions can also call capability methods, thus they can instantiate capability objects and pass them around to the rest of the program.

L42 expects capability methods to be used mostly internally by capability classes, whereas user code would call normal methods on already existing capability objects.

For the purposes of invariant checking, we only care about the effects that methods could have on the running program and heap. As such, output methods (such as a print method) can be whitelisted as ‘deterministic’, provided they do not affect program execution, such as by non deterministically throwing I/O errors.

**Purity**

TMs and OCs together statically guarantee that any method with only read or imm parameters (including the receiver) is pure; we define pure as being deterministic and not mutating existing memory. Such methods are pure because:

- the ROG of the parameters (including this) is only accessible as read (or imm), thus it cannot be mutated.
- if a capability object is in the ROG of any of the arguments (including the receiver), then it can only be accessed as read, preventing calling any non deterministic (capability) methods,
- no other preexisting objects are accessible (as L42 does not have global variables).  

3 **Our Invariant Protocol**

Our invariant protocol guarantees that the whole ROG of any object involved in execution (formally, in a redex) is valid: if you can call methods on an object, calling invariant on it is guaranteed to return true in a finite number of steps. However, calls to invariant that are generated by our runtime monitoring (see below) can access the fields of a potentially invalid this. This is necessary to allow for the invariant method to do its job: namely distinguish between valid and invalid objects. However, as for any other method, calls to invariant written explicitly by users are guaranteed to have a valid receiver.

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9 This is even true in the concurrent environments of Pony and Gordon, since they ensure that no other thread/actor has access to a mut/capsule alias of this. Thus, since such methods do not write to memory accessible by another thread, nor read memory that could be mutated by another thread, they are atomic.

10 If L42 did have static variables, getters and setters for them would be capability methods. Even allowing unrestricted access to imm static variables would prevent reasoning over determinism, due to the possibility of global variable updates; however constant/final globals of an imm type would not cause such problems.
For simplicity, in the following explanation and in our formalism we require receivers to always be specified explicitly, and require that the receivers of field accesses and updates are always \texttt{this}; that is, all fields are instance private. We also do not allow explicit constructor definitions, instead we assume constructors are of the standard form
\begin{verbatim}
C(T_1 x_1,..,T_n x_n) {this.f_1=x_1;...;this.f_n=x_n;},
\end{verbatim}
where the fields of \( C \) are \( T_1 f_1;..;T_n f_n; \). This ensures that partially uninitialised (and likely invalid) objects are not passed around or used. These restrictions only apply to our formalism; our code examples and the L42 implementation soundly relax these, see below for a discussion.

\textbf{Invariants}

We require that all classes contain a \texttt{read} method \texttt{Bool invariant() {..}}, if no \texttt{invariant} method is present, a trivial one returning \texttt{true} will be assumed. As this method only takes a \texttt{read} parameter (the receiver), we can be sure that it is pure \cite{section2} as discussed in Section \ref{section2}. The bodies of \texttt{invariant} methods are limited in their usage of \texttt{this}: \texttt{this} can only be used to access \texttt{imm} and \texttt{capsule} fields. This restriction ensures that an invalid \texttt{this} cannot be passed around. We prevent accessing \texttt{mut} fields since their ROG could be changed by unrelated code (see Section \ref{section5}). Note that we do not require such fields to be \texttt{final}: when a field is updated, we simply check the invariant of the receiver of the update.

\textbf{Capsule mutators}

In order to allow complex mutations of objects with invariants we introduce the notion of \texttt{capsule mutator}. A \texttt{capsule mutator} can perform an arbitrarily complex mutation of the ROG of a capsule field. We use TMs to ensure that the object containing the capsule field is not usable whilst the fields ROG is mutated, and it’s invariant is checked immediately afterwards.

Formally, \texttt{capsule mutators} are \texttt{mut} methods whose body accesses a \texttt{capsule} field mentioned in the invariant of the class containing the field. Capsule mutators must use \texttt{this} exactly once in their body, since fields are instance private, this will be to access the \texttt{capsule} field. Excluding the \texttt{mut} receiver, such methods cannot have any \texttt{mut} or \texttt{read} parameters, their return type must not be \texttt{mut}, and their \texttt{throws} clause must be empty \cite{section3}.

As capsule mutators use \texttt{this} only once, and have no \texttt{read} or \texttt{mut} parameters, \texttt{this} will not be accessible during execution. This is important, as it allows the invariant to be violated part way through the capsule mutator, but re established by the end. Preventing \texttt{mut} return types ensures that such methods cannot leak out a mutable alias to the \texttt{capsule} field, which could then be used to break the invariant. Note that these restrictions do not apply when the receiver of the field access is \texttt{capsule}, since we guarantee that the receiver is not in the ROG of any of its \texttt{capsule} fields, and hence it can never be seen afterwards.

\textbf{Monitoring}

The language runtime will insert automatic calls to \texttt{invariant}, if such a call returns \texttt{false}, an unchecked exception will be thrown. Such calls are inserted in the following points:

\begin{itemize}
  \item After a constructor call, on the newly created object.
  \item After a field update, on the receiver.
  \item After a capsule mutator method returns, on the receiver of the method\cite{section4}
\end{itemize}

\footnote{If the invariant were not pure, it would be nearly impossible to ensure that it would return \texttt{true} at any point.}

\footnote{To allow capsule mutators to leak checked exceptions, we would need check the invariant when such exceptions are leaked. However, this would make the runtime semantics of checked exceptions inconsistent with unchecked ones.}

\footnote{The invariant is not checked if the call was terminated via an an unchecked exception, since strong
In Appendix A we show that these checks, together with our aforementioned restrictions, are sufficient to ensure our guarantee that all objects involved in execution (except as part of an invariant check) are valid.

Relaxations
The above restrictions can be partially relaxed without breaking soundness, however this would not make the proof more interesting. In particular:

- invariant methods can be allowed to call instance methods that in turn only use this to read imm or capsule, or call other such instance methods. With this relaxation, the semantics of invariant needs to be understood with the body of those methods inlined; thus the semantics of the inlined code needs to be logically reinterpreted in the context of invariant, where this may be invalid. In some sense, those inlined methods and field accesses can be thought of as macro expanded, and hence are not dynamically dispatched. Such inlining has been implemented in L42.

- We could allow all fields to be public, however capsule fields, mentioned in the invariant of their containing class, should not be accessible over a mut receiver other than this. Even without this relaxation, however, getters and setters could be used to simulate public fields.

- Unrestricted readonly access to capsule fields can be allowed by automatically generated getters of the form read method read C f() { return this.f; }. Such getters are already a fundamental part of the L42 language.

- Java style constructors could be allowed, provided that this is only used as the receiver of field initialisations. L42 does not provide such constructors, but one can always write a static factory method that behaves equivalently.

Both L42, and our formal language (see Section 4) do not have traditional subclassing, rather all ‘classes’ are either interfaces (which only have abstract methods), or are final (which cannot be subtyped). In a language with traditional subclassing, invariant methods would implicitly start with a check that super.invariant() returns true. Note that invariant checks would not be performed at the end of super(..) constructor calls, but only at the end of new expressions, as happens in §30.

4 Formal Language Model

In order to model our system, we need to formalise an imperative object oriented language with exceptions, object capabilities, and rich type system support for TMs and strong exception safety. Formal models of the runtime semantics of such languages are simple, but defining and proving, such a type system would require a paper of its own, and indeed many such papers exist in literature [67, 65, 58, 22, 43]. Thus we are going to assume that we already have an expressive and sound type system enforcing the properties we need, and instead focus on invariant checking. We clearly list in Appendix A the assumptions we make on such a type system, so that any language satisfying them, such as L42, can soundly support our invariant protocol.

To keep our small step semantics as conventional as possible, we follow Pierce [64] and Featherweight Java [11], and assume:

- An implicit program/class table.
class implements = expressions are of the form

To encode object capabilities and I/O, we assume a special location

of memory. This is used to model the guarantee of strong exception safety, that is, the
unchecked exception.

Here, \( l \) is the object being monitored, \( e_1 \) is the expression which is being monitored, and \( e_2 \) denotes the evaluation of \( l . i n v a r i a n t () \). If, at any point in execution, \( e_2 \) is \( \text{false} \), then \( l \)'s invariant failed to hold; such a monitor expression corresponds to the throwing of an unchecked exception.

In addition, our reduction rules will annotate \( \text{try} \) expressions with the original state of memory. This is used to model the guarantee of strong exception safety, that is, the annotated memory will not be mutated by executing the body of the \( \text{try} \).
Well Formedness Criteria
We additionally restrict the grammar with the following well formedness criteria:
- **invariant** methods and capsule mutators satisfy the restrictions in Section 3.
- Field accesses and updates in methods are of the form `this.f` or `this.f = e`, respectively.
- Field accesses and updates in the main expression are of the form `l.f` or `l.f = e`, respectively.
- Locations that are preserved by `try` blocks are never monitored, that is, for `try` expressions, and calls to capsule mutators. Monitor expressions are only a proof device, they need not be implemented directly as presented. For example, in L42 we implement them by statically injecting calls to `invariant` at the end of setters, factory methods and capsule mutators; this works as L42 does not have primitive expression forms for field updates and constructors, rather they are uniformly represented as method calls.

Our **CTXV** rule evaluates monitor expressions, `[](l;_;_`), by first evaluating `e1` and then `e2`. If `e2` evaluates to `true`, then the monitor succeeded, and will yield the result of `e1`. If however `e2` evaluated to `false`, then the monitor failure will be caught by our **TRY ERROR** rule, as will any other uncaught monitor failure in `e1` or `e2`.

**Statement of Soundness**
We define a deterministic reduction to mean that exactly one reduction is possible:

\[ σ₀|e₀ → σ|e₁ \] if \( \{σ₁|e₁\} = \{σ|e\} \) where \( σ₀|e₀ → σ|e \)

An object is **valid** if calling its `invariant` method would deterministically produce `true` in a finite number of steps, i.e. it does not evaluate to `false`, fail to terminate, or produce an
error. We also require evaluating invariant to preserve existing memory ($\sigma$), however new objects ($\sigma'$) can be created and freely mutated.

$\text{valid}(\sigma, l) \iff \sigma|l$.\text{invariant}() $\rightarrow^+ \sigma, \sigma'|\text{true}$.

To allow the invariant method to be called on an invalid object, and access fields on such object, we define the set of trusted execution steps as the the call to invariant itself, and any field accesses inside its evaluation. Note that this only applies to single small step reductions, and not the entire evaluation of invariant.

$\text{trusted}(E_v, r_l) \iff$

- either $r_l = l$.\text{invariant}() and $E_v = E'_v[l(l;v;\Box)]$, or
- $r_l = l.f$ and $E_v = E'_v[l(l;v;E''_v)]$.

Finally, we define what it means to soundly enforce our invariant protocol: every object referenced by any untrusted redex is valid.

▶ Theorem 1 (Soundness). if $c: \text{Cap}\emptyset \vdash e: T$ and $c \mapsto \text{Cap}[_{-}]|e \rightarrow^+ \sigma|\text{true}[E_v][r_l]$, then either $\text{valid}(\sigma, l)$ or $\text{trusted}(E_v, r_l)$.

5 Invariants Over Immutable State

In this section we consider validation over fields of imm types. In the next section we detail our technique for capsule fields.

In the following code Person has a single immutable (non final) field name:

```java
class Person {
    read method Bool invariant() { return !name.isEmpty(); } 
    private String name; // the default modifier imm is applied here
    read method String name() { return this.name; } 
    mut method String name(String name) { this.name = name; } 
    Person(String name) { this.name = name; }
}
```

Person only has immutable fields and its constructor only uses this to initialise them. Note that the name field is not final, thus Person objects can change state during their lifetime. This means that the ROGs of all Person fields are immutable, but Persons themselves may be mutable. We can easily enforce Person’s invariant by generating checks on the result of this.invariant(): immediately after each field update, and at the end of the constructor.

```java
class Person {
    .. // Same as before
    mut method String name(String name) {
        this.name = name; // check after field update
        if (!this.invariant()) { throw new Error(...); }
    }
    Person(String name) {
        this.name = name; // check at end of constructor
        if (!this.invariant()) { throw new Error(...); }
    }
}
```

Such checks will be generated/injected, and not directly written by the programmer. If we were to relax (as in Rust), or even eliminate (as in Java), the support for TMs or OCs, the enforcement of our invariant protocol for the Person class would become harder, or even impossible.

Unrestricted use of non determinism

Allowing the invariant method to (indirectly) perform a non deterministic operation, such
as by creating new capability objects, could break our guarantee that (manually) calling it always returns `true`. For example consider this simple and contrived (mis)use of person:

```java
class EvilString extends String {
    @Override
    read method Bool isEmpty() {
        // Create a new capability object out of thin air
        return new Random().bool();
    }
}
method mut Person createPersons(String name) {
    // we can not be sure that name is not an EvilString
    mut Person schrodinger = new Person(name); // exception here?
    assert schrodinger.invariant(); // will this fail?
}
```

Despite the code for `Person.invariant` intuitively looking correct and deterministic, the above call to it is not. Obviously this breaks any reasoning and would make our protocol unsound. In particular, note how in the presence of dynamic class loading, we have no way of knowing what the type of `name` could be. Since our system allows non determinism only through capability objects, and restricts their creation, the above example would be prevented.

### Allowing Internal Mutation Through Back Doors

Suppose we relax our rules by allowing interior mutability as in Rust and Java, where sneaky mutation of the ROG of an ‘immutable’ object is allowed. Those back doors are usually motivated by performance reasons, however in [38] they briefly discuss how a few trusted language primitives can be used to perform caching and other needed optimisations, without the need for back doors.

Our example shows that such back doors can be used to break determinism of `invariant` methods, by allowing the invariant to store and read information about previous calls. In the following example we use `MagicCounter` as a back door to remotely break the invariant of `person` without any interaction with the `person` object itself:

```java
class MagicCounter {
    method Int increment() {
        // Magic mutation through an imm receiver, equivalent to i++
    }
}
class NastyS extends String {
    MagicCounter evil = new MagicCounter(0);
    @Override
    read method Bool isEmpty() {
        return this.evil.increment() != 2;
    }
}
NastyS name = new NastyS("bob"); // TMs believe name’s ROG is imm
Person person = new Person(name); // person is valid, counter=1
name.increment(); // counter == 2, person is now broken
person.invariant(); // returns false!, counter == 3
person.invariant(); // returns true, counter == 4
```

### Strong Exception Safety

The ability to catch and recover from invariant failures is extremely useful as it allows programs to take corrective action. Since we represent invariant failures by throwing unchecked exceptions, programs can recover from them with a conventional `try-catch`. Due
to the guarantees of strong exception safety, any object that has been mutated during a `try` block is now unreachable (as happens in alias burying [13]). In addition, since unchecked exceptions are immutable, they cannot contain a `read` reference to any object (such as the `this` reference seen by `invariant` methods). These two properties ensure that an object whose invariant fails will be unreachable after the invariant failure has been captured. If instead we were to not enforce strong exception safety, an invalid object could be made reachable:

```java
mut Person bob = new Person("bob");
// Catch and ignore invariant failure:
try { bob.name(""}; } catch (Error t) { } // ill typed in L42
assert bob.invariant(); // bob is invalid!
```

As you can see, recovering from an invariant failure in this way is unsound and would break our protocol.

## 6 Invariants over encapsulated state

Consider managing the shipment of items, where there is a maximum combined weight:

```java
class ShippingList {
    capsule Items items;
    read method Bool invariant() {
        return this.items.weight() <= 300;
    }
    ShippingList(capsule Items items) {
        this.items = items;
        if (!this.invariant()) {throw Error(...};} // injected check
    }
    mut method Void addItem(Item item) {
        this.items.add(item);
        if (!this.invariant()) {throw Error(...};} // injected check
    }
}
```

To handle this class we just inject calls to `invariant` at the end of the constructor and the `addItem` method. This is safe since the `items` field is declared `capsule`. Relaxing our system to allow a `mut` modifier for the `items` field and the corresponding constructor parameter breaks the code: the cargo we received in the constructor may already be compromised:

```java
mut Items items = ...;
mut ShippingList l = new ShippingList(items); // l is valid
items.addItem(new HeavyItem()); // l is now invalid!
```

As you can see it would be possible for external code with no knowledge of the `ShippingList` to mutate its items[14]

Our restrictions on capsule mutators ensure that capsule fields are essentially an exclusive mutable reference. Removing these restrictions would break our invariant protocol. If we were to allow `x.items` to be seen as `mut`, where `x` is not `this`, then even if the `ShippingList`

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[14] Conventional ownership solves these problems by requiring a deep clone of all the data the constructor takes as input, as well as all exposed data (possibly through getters). In order to write correct library code in mainstream languages like Java and C++, defensive cloning [12] is needed. For performance reasons, this is hardly done in practice and is a continuous source of bugs and unexpected behaviour [12].
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has full control at initialisation time, such control may be lost later, and code unaware of the ShippingList could break it:

```java
mut ShippingList l = new ShippingList(new Items()); // l is ok
mut Items evilAlias = l.items // here l loses control
evilAlias.addItem(new HeavyItem()); // now l is invalid!
```

If we allowed a mut return type the following would be accepted:

```java
mut method mut Items expose(C c) { return c.foo(this.items);}
```

Depending on dynamic dispatch, c.foo() may just be the identity function, thus we would get in the same situation as the former example.

Allowing this to be used more than once can also cause problems:

```java
mut method imm Void multiThis(C c) {
    read Foo f = c.foo(this);
    this.items.add(new HeavyItem());
    f.hi(); } // Can ‘this’ be observed here?
```

If the former code were accepted, this may be reachable from f, thus f.hi() may observe an invalid object.

In order to ensure that a second reference to this is not reachable through the parameters, we only accept imm and capsule parameters. If we were however to accept a read parameter, as in the example below, we would be in the same situation as before, where f may contain a reference to this:

```java
mut method imm Void addHeavy(read Foo f) {
    this.items.add(new HeavyItem());
    f.hi(); } // Can ‘this’ be observed here?
...
mut ShippingList l = new ShippingList();
read Foo f = new Foo(l);
l.addHeavy(f); // We pass another reference to ‘l’ through f
```

7 GUI Case study

Here we show that we are able to verify classes with circular mutable object graphs, that interact with the real world using I/O. Our case study involves a GUI with containers (SafeMovable) and Buttons; the SafeMovable class has an invariant to ensure that its children are completely contained within it and do not overlap. The Buttons move their SafeMovable when pressed. We have a Widget interface which provides methods to get Widgets’ size and position as well as children (a list of Widgets). Both SafeMatables and Buttons implement Widget. Crucially, since the children of SafeMovable is a list of Widgets it can contain other SafeMatables, and all queries to their size and position are dynamically dispatched, such queries are also used in SafeMovable’s invariant. Here we show a simplified version\(^{15}\) where SafeMovable has just one Button, and certain sizes and positions are fixed. Note that Widgets is a class representing a mutable list of mut Widgets.

\(^{15}\)The full version, written in L42, which uses a different syntax, is available in our artifact at http://l42.is/EcoopArtifact.zip
class SafeMovable implements Widget {
capsule Box box;

@Override read method Int left() { return this.box.l; }
@Override read method Int top() { return this.box.t; }
@Override read method Int width() { return 300; }
@Override read method Int height() { return 300; }
@Override read method read Widgets children() {
    return this.box.c;
}
@Override mut method Void dispatch(Event e) {
    for (Widget w: this.box.c) { w.dispatch(e); }
}
read method Bool invariant() {}

SafeMovable(capsule Widgets cs) { this.box = makeBox(c); }
static method capsule Box makeBox(capsule Widgets c) {
    mut Box b = new Box(5, 5, cs);
    b.c.add(new Button(0, 0, 10, 10, new MoveAction(b));
    return b; } // mut b is soundly promoted to capsule
}
class Box { Int l; Int t; mut Widgets c; }
Box(Int l, Int t, mut Widgets c) {...}
class MoveAction implements Action {
    mut Box outer;
    MoveAction(mut Box outer) { this.outer = outer; }
    mut method Void process(Event event) { this.outer.l += 1; }
}

// main expression; #$ is a capability method making a Gui object
Gui.#$().display(new SafeMovable(..));

As you can see, Boxes encapsulate the state of the SafeMovable that can change over time: left, top, and children. Also note how the ROG of Box is circular: since the MoveActions inside Buttons need a reference to the containing Box in order to move it. Even though the children of SafeMovable are fully encapsulated, we can still easily dispatch events to them using dispatch. Once a Button receives an Event with a matching ID, it will call its Action’s process method.

Our example shows that the restrictions of TMs and OCs are flexible enough to encode interactive GUI programs, where widgets may circularly reference other widgets. In order to perform this case study we had to first implement a simple GUI Library in L42. This library uses object capabilities to draw the widgets on screen, as well as fetch and dispatch the events. Importantly, neither our application, nor the underlying GUI library require back doors into either our type modifier or capability system to function, demonstrating the practical usability of our restrictions.

The Invariant

SafeMovable is the only class in our GUI that has an invariant, our system automatically checks it in two places: the end of its constructor and the end of its dispatch method (is a capsule mutator). There are no other checks inserted since we never do a field update on a SafeMovable. The code for the invariant is just a couple of simple nested loops:

read method Bool invariant() {
    for(Widget w1 : this.box.c) {
        ...


```java
if(!this.inside(w1)) { return false; }
for(Widget w2 : this.box.c) {
    if(w1!=w2 && SafeMovable.overlap(w1, w2)){return false;}}
return true;
```

Here `SafeMovable.overlap` is a static method that simply checks that the bounds of the widgets don’t overlap. The call to `this.inside(w1)` similarly checks that the widget is not outside the bounds of `this`; this instance method call is allowed as `inside` only uses `this` to access its fields.

Our Experiment
As shown in the figure to the left, counting both `SafeMovable` and `Button`s, our main method creates 21 widgets: a top level (green) `SafeMovable` without buttons, containing 4 (red, blue, and black) `SafeMovable`s with 4 (gray) buttons each. When a button is pressed it moves the containing `SafeMovable` a small amount in the corresponding direction. This setup is not overly complicated, the maximum nesting level of `Widget` is 5. Our main method automatically presses each of the 16 buttons once. In L42, using the approach of this paper, this resulted in 77 calls to `SafeMovable`’s invariant.

Comparison With Visible State Semantics
As an experiment, we set our implementation to generate invariant checks following the visible state semantics approaches of D and Eiffel [3, 24], where the invariant of the receiver is instead checked at the start and end of every public (in D) and qualified (in Eiffel) method calls. In our `SafeMovable` class, all methods are public, and all calls are qualified, thus this difference is irrelevant. Neither protocol performs invariant checks on field accesses or updates, however due to the ‘uniform access principle’, Eiffel allows fields to directly implement methods, allowing the `width` and `height` fields to directly implement `Widget`’s `width` and `height` methods. On the other hand in D, one would have to write getter methods, which would invoke invariant checks. When we ran our test case following the D approach, the `invariant` method was called 52,734,053 times, whereas the Eiffel approach ‘only’ called it 14,816,207 times; in comparison our invariant protocol only performed 77 calls. The number of checks is exponential in the depth of the GUI: the invariant of a `SafeMovable` will call the `width`, `height`, `left`, and `top` methods of its children, which may themselves be `SafeMovable`s, and hence such calls may invoke further invariant checks. Note that `width` and `height` are simply getters for fields, whereas the other two are non trivial methods.

Spec# Comparison
We also encoded our example in Spec#[17], which like L42, statically verifies aliasing/ownership properties, as well as the admissibility of invariants. The backend of the L42 GUI library is written in Java, we did not port it to Spec#, rather we just simulate the backend, and don’t actually display a GUI in Spec#.

16 That is, the receiver is not `this`.
17 We compiled Spec# using the latest available source (from 19/9/2014). The verifier available online at `rise4fun.com/SpecSharp` behaves differently.
We ran our code through the Spec\# verifier (powered by Boogie \[4\]), which only gave us 2 warnings\[18\] that the invariant of SafeMovable was not known to hold at the end of its constructor and dispatch method. Like our system however, Spec\# checks the invariant at those two points at runtime. Thus the code is equivalently verified in both Spec\# and L42; in particular it performed exactly the same number (77) of runtime invariant checks\[19\].

We found it quite difficult to encode the GUI in Spec\#, due to its unintuitive and rigid ownership discipline. In particular we needed to use many more annotations, which were both larger and had greater variety. In the following table we summarise the annotation burden, for the program that defines and displays the SafeMovable\[20\] and our GUI; as well as the library which defines Buttons, Widget, and event handling.

|                          | Spec\# program | Spec\# library | L42 program | L42 library |
|--------------------------|----------------|----------------|-------------|-------------|
| Total number of annotations | 40             | 19             | 19          | 18          |
| Tokens (except \.;\[\]) and whitespace) | 106            | 34             | 18          | 18          |
| Characters (with minimal whitespace) | 619            | 207            | 74          | 60          |

To encode the GUI example in L42, the only annotations we needed were the 3 type modifiers: mut, read, and capsule. Our Spec\# code requires things such as, purity, immutability, ownership, method pre/post conditions and method modification annotations. In addition, it requires the use of 4 different ownership functions including explicit ownership assignments. In total we used 18 different kinds of annotations in Spec\#. Together these annotations can get quite long, such as the following precondition on SafeMovable’s constructor:

```
requires Owner.Same(Owner.ElementProxy(children), children);
```

The Spec\# code also required us to deviate from the style of code we showed in our simplified version: we could not write a usable children method in Widget that returns a list of children, instead we had to write children_count() and children(int i) methods; we also needed to create a trivial class with a [Pure] constructor (since Object’s one is not marked as such). In contrast, the only strange thing we had to in L42 was creating Boxes by using an additional variable in a nested scope. This is needed to delineate scopes for promotions. Based on these results, we believe our system is significantly simpler and easier to use.

**The Box Pattern**

Our design, using an inner Box object, is a common pattern in static verification: where one encapsulates all relevant mutating state into an encapsulated sub object which is not exposed to users.

Both our L42 and Spec\# code required us to use the box pattern for our SafeMovable, due to the circular object graph caused by the Actions of Buttons needing to change their enclosing SafeMovable’s position.

**The Transform Pattern**

Suppose we want to scale a Widget, we could add mut setters for width, height, left, and top in the Widget interface. However, if we also wish to scale its children we have a problem,

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\[18\] We used assume statements, equivalent to Java’s assert, to dynamically check array bounds. This aligns the code with L42, which also performs such checks at runtime.

\[19\] We also encoded our GUI in Microsoft Code Contracts \[27\], whose unsound heuristic also calls the invariant 77 times; however Code Contract does not enforce the encapsulation of children, thus their approach would not be sound in our context.

\[20\] We only count constructs Spec\# adds over C# as annotations, we also do not count annotations related to array bounds or null checks.
since `Widget.children` returns a `read Widgets`, which does not allow mutation. We could of course add a `mut` method `zoom` to the `Widget` interface, however this does not scale if more operations are desired. If instead `Widget.children` returned a `mut Widgets`, it would be difficult for `Widget` implementations, such as `SafeMovable`, to keep control of their ROGs.

A simple and practical solution would be to define a `transform` method in `Widget`, and a `Transformer` interface like so:

```java
interface Transformer<T> { method Void apply(mut T elem); }
interface Widget {
    mut method Void top(Int that); // setters for immutable data
    mut method read Void transform(Transformer<Widgets> t);
} // transformers for possibly encapsulated data
class SafeMovable {
    mut method Void transform(Transformer<Widgets> t) {
        return t.apply(this.box.c); }
} // Well typed capsule mutator
```

The `transform` method offers an expressive power similar to `mut` getters, but prevents `Widgets` from leaking out. With a `Transformer`, a `zoom` function could be simply:

```java
static method Void zoom(mut Widget w) {
    w.transform(ws -> {
        for (wi : ws) { zoom(wi, scale); }
    });
    w.width(w.width() / 2); ..; w.top(w.top() / 2); }
```

## 8 Related work

### Type Modifiers

We rely on a combination of modifiers that are supported by at least 3 languages/lines of research: L42 [68, 67, 43, 35], Pony [22, 23], and Gordon et. al. [38]; each of these works is accompanied by proofs about the properties of those modifiers. Since such proofs have already been done, in this work we just assume the required properties. Those approaches all support deep/strong interpretation, without back doors.

TM approaches like Javari [75, 15] and Rust [49] are unsuitable since they introduce back doors which are not easily verifiable as being used properly. Many approaches just try to preserve purity (as for example [63]), but here we also need aliasing control. Ownership [20, 79, 25] is another popular form of aliasing control that can be used as a building block for static verification [57, 7]. Capsule/isolated local variables are affine in that they can be used only once, however this linearity is a property of variables, not expressions or fields. Linear/affine types extend this idea further, however they usually do not consider the ROGs of such types, or work in an OO setting [56, 29].

### Object Capabilities

In literature, OCs are used to provide a wide range of guarantees, and many variations are present. Object capabilities [55], in conjunction with type modifiers, are able to enforce purity of code in a modular way, without requiring the use of monads. L42 and Gordon use OCs simply to reason about I/O and non determinism. This approach is best exemplified by Joe-E [32], which is a self contained and minimalistic language using OCs over a subset of Java in order to reason about determinism. However, in order for Joe-E to be a subset of

---

21 A more general transformer could return a generic `read R`. 

Java, they leverage on a simplified model of immutability: immutable classes must be final with only final fields that refer to immutable classes. In Joe-E, every method that only takes instances of immutable classes is pure. Thus their model would not allow the verification of purity for invariant methods of mutable objects. In contrast our model has a more fine grained representation of mutability: it is reference based instead of class based. In our work, every method taking only read or imm references is pure, regardless of their class type.

**Class invariant protocols**

Class invariants are a fundamental part of the design by contract methodology. Invariant protocols differ wildly and can be unsound or complicated, particular due to re entrancy and aliasing [45, 26, 53].

While invariant protocols all seem to check and assume the invariant of an object after its construction, they handle invariants differently across object lifetimes; popular sound approaches include:

- The invariants of objects in a steady state are known to hold: that is when execution is not inside any of the objects public methods [36]. Invariants need to be constantly maintained between calls to public methods [77].
- The invariant of the receiver before a public method call and at the end of every public method body needs to be ensured. The invariant of the receiver at the beginning of a public method body and after a public method call can be assumed [17, 26]. Some approaches ensure the invariant of the receiver of the calling method, rather than the called method [58]. JML [34] relaxes these requirements for helper methods, whose semantic is the same as if they were inlined.
- The same as above, but only for the bodies of ‘selectively exported’ (i.e. non instance private) methods, and only for ‘qualified’ (i.e. not this) calls [53].
- The invariant of an object is assumed only when a contract requires the object be ‘packed’. It is checked after an explicit ‘pack’ operation, and objects can later be ‘unpacked’ [5].
- Or, as in this work, the invariant of any object which could be involved in execution is assumed to hold. It is checked after every modification of the object or its encapsulated ROG.

These different protocols can be deceivingly similar, and some approaches like JML suggest verifying a simpler approach (that method calls preserve the invariant of the receiver) but assume a stronger one (the invariant of every object, except this, holds).

**Runtime Verification Tools**

Many languages and tools support some form of runtime invariant checking (e.g. Eiffel [52], D [3], and JML [17]). By looking to a survey by Voigt et al. [76] and the extensive MOP project [50], it seems that most runtime verification tools (RV) empower users to implement the kind of monitoring they see fit for their specific problem at hand. This means that users are responsible for deciding, designing, and encoding both the logical properties and the instrumentation criteria [50]. In the context of class invariants, this means the user defines the invariant protocol and the soundness of such protocol is not checked by the tool.

In practice, this means that the logic, instrumentation, and implementation end up connected: a specific instrumentation strategy is only good to test certain logic properties in certain applications. No guarantee is given that the implemented instrumentation strategy is able to support the required logic in the monitored application. Some of these tools are designed to support class invariants: for example InvTS [37] lets you write Python conditions that are verified on a set of Python objects, but the programmer needs to be able to predict which objects are in need of being checked and to use a simpler domain specific language to target them. Hence if a programmer makes a mistake while using this domain specific
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language, invariant checking will not be triggered. Some tools are intentionally unsound and just perform invariant checking following some heuristic that is expected to catch most failures: such as jmlrac [17] and Microsoft Code Contracts [28].

Many works attempt to move out of the ‘RV tool’ philosophy to ensure RV monitors work as expected, as for example the study of contracts as refinements of types [31]. However, such work is only interested in pre and post conditions, not class invariants.

Our invariant protocol is much stronger then visible state semantics, and keeps the invariant under tight control. Gopinathan et. al.’s. approach keeps a similar level of control: relying on powerful aspect oriented support, they detect any field update in the whole ROG of any object, and check all the invariants that such update may have violated.

We agree with their criticism of visible state semantics, where methods still have to assume that any object may be broken; in such case calling any public method would trigger an error, but while the object is just passed around (and for example stored in collections), the broken state will not be detected; Gopinathan et. al. says “there are many instances where o’s invariant is violated by the programmer inadvertently changing the state of p when o is in a steady state. Typically, o and p are objects exposed by the API, and the programmer (who is the user of the API), unaware of the dependency between o and p, calls a method of p in such a way that o’s invariant is violated. The fact that the violation occurred is detected much later, when a method of o is called again, and it is difficult to determine exactly where such violations occur.”

However, their approach addresses neither exceptions nor non determinism caused by I/O, so their work is unsound if those aspects are taken into consideration.

Their approach is very computationally intensive, but we think it is powerful enough that it could even be used to roll back the very field update that caused the invariant to fail, making the object valid again. We considered a roll back approach for our work, however rolling back a single field update is likely to be completely unexpected, rather we should roll back more meaningful operations, similarly to what happens with transactional memory, and so is likely to be very hard to support efficiently. Using TMs to enforce strong exception safety is a much simpler alternative, providing the same level of safety, albeit being more restrictive (namely that if the operation did succeed it is still effectively rolled back).

Chaperones and impersonators [71] lifts the techniques of gradual typing [73, 74, 78] to work on general purpose predicates, where values can be wrapped to ensure an invariant holds. This technique is very powerful and can be used to enforce pre and post conditions by wrapping function arguments and return values. This technique however does not monitor the effects of aliasing, as such they may notice if a contract has broken, but not when or why. In addition, due to the difficulty of performing static analysis in weakly typed languages, they need to inject runtime checking code around every user facing operation. Aspect oriented systems like Jose [30], similarly wrap invariant checks around method bodies.

Security and Scalability

Our approach allows verifying an object’s invariant independently of the actual invariants of other objects. This is in contrast with the main strategy of static verification: to verify a method, the system assumes the contracts of other methods, and the content of those contracts is the starting point for their proof. Thus, static verification proceeds like a mathematical proof: a program is valid if it is all correct, but a single error invalidates all claims. This makes it hard to perform verification on large programs, or when independently maintained third party libraries are involved. This is less problematic with a type system, since its properties are more coarse grained, simpler and easier to check. Static verification has more flexible and fine grained annotations and often relies on a fragile theorem prover as
To soundly verify code embedded in an untrusted environment, as in gradual typing \[74,78\], it is possible to consider a verified core and a runtime verified boundary. You can see our approach as an extremely modularized version of such system: every class is its own verified core, and the rest of the code could have Byzantine behaviour. Our formal proofs show that every class that compiles/type checks is soundly handled by our protocol, independently of the code that uses such class or any other surrounding code.

Our approach works both in a library setting and with the open world assumption. Consider for example the work of Parkinson \[62\]: in his short paper he verified a property of the Subject/Observer pattern. However, the proof relies on (any override of) the Subject.register(Observer) method respecting its contract. Such assumption is unrealistic in a real world system with dynamic class loading, and could trivially be broken by a user defined EvilSubject.

Static Verification
Spec# \[8\] is a language built on top of C#, it adds various annotations such as method contracts and class invariants. It primarily follows the Boogie methodology \[59\] where (implicit) annotations are used to specify and modify the owner of objects and whether their invariants are required to hold. Invariants can be ownership based \[5\], where an invariant only depends on objects it owns; or visibility based \[6, 46\], where an invariant may depend on objects it doesn’t own, provided that the class of such objects know about this dependence. Unlike our approach, Spec# does not restrict the aliases that may exist for an object, rather it restricts object mutation: an object cannot be modified if the invariant of its owner is required to hold. This is more flexible than our approach as it also allows only part of an object’s ROG to be owned/encapsulated. However as we showed in Section \[7\] it can become much more difficult to work with and requires significant annotation since merely having an alias to an object is insufficient to modify it or call methods on it. Spec# also works with existing .NET libraries by annotating them with contracts, however such annotations are not verified. Spec#, like us, does perform runtime checks for invariants and throws unchecked exceptions on failure. However Spec# does not allow soundly recovering from an invariant failure, since catching unchecked exceptions in Spec# is intentionally unsound. \[48\]

Another system is AutoProof \[65\], a static verifier for Eiffel that also follows the Boogie methodology, but extends it with semantic collaboration where objects keep track of their invariants’ dependencies using ghost state. Dafny \[44\] is a new language where all code is statically verified, it supports invariants by injecting pre and post conditions following visible state semantics; however it requires objects to be newly allocated (or cloned) before another object’s invariant may depend on it. Dafny is also generally highly restrictive with its rules for mutation, and object construction, it also does not provide any means of performing non deterministic I/O.

Specification languages
Using a specification language based on the mathematical metalanguage and different from the program language’s semantics may seem attractive, since it can express uncomputable concepts, has no mutation or non determinism, and is often easier to formally reason about.

However, a study \[18\] discovered that developers expect specification languages to follow the semantics of the underlying language, including short circuit semantics and arithmetic exceptions; thus for example \(1/0 \lor 2 \lor 1\) should not hold, while \(2 \lor 1 \lor 1/0\) should, thanks to short circuiting. This study was influential enough to convince JML to change its interpretation of logical expressions accordingly \[19\]. Dafny \[44\] uses a hybrid approach: it has mostly the same language for both specification and execution. Specification (‘ghost’)
contexts can use uncomputable constructs such as universal quantification over infinite sets. Whereas runtime contexts allow mutation, object allocation and print statements. The semantics of shared constructs (such as short circuiting logic operators) is the same in both contexts.

Most runtime verification systems, such as ours, use a metacircular approach: specifications are simply code in the underlying language. Since specifications are checked at runtime, they are unable to verify uncomputable contracts. Ensuring determinism in a non functional language is challenging. Spec# recognizes the need for purity/determinism when method calls are allowed in contracts. There are three main current approaches: a) forbid the use of functions in specifications, b) allow only provably pure functions, or c) allow programmers free use of functions. The first approach is not scalable, the second overly restrictive and the third unsound.

They recognize that many tools unsoundly use option (c), such as AsmL. Spec# aims to follow (b) but only considers non determinism caused by memory mutation, and allows other non deterministic operations, such as I/O and random number generation. For example, the following method verifies:

\[
\text{[Pure]} \quad \text{bool uncertain()} \{ \text{return new Random().Next()} \% 2 == 0; \}
\]

And so assert uncertain() == uncertain(); also verifies, but randomly fails with an exception at runtime. As you can see failing to handle non determinism jeopardises reasoning.

A simpler and more restrictive solution to these problems is to prevent ‘pure’ functions from reading or writing to any non final fields, or calling any impure functions. This is the approach used by, one advantage of their approach is that invariants (which must be ‘pure’) can read from a chain of final fields, even when they are contained in otherwise mutable objects. However their approach completely prevents invariants from mutating newly allocated objects, thus greatly restricting how computations can be performed.

9 Conclusions and Future Work

Our approach follows the principles of offensive programming, where no attempt to fix or recover an invalid object is performed and failures (unchecked exceptions) are raised close to their cause: at the end of constructors creating invalid objects and immediately after field updates and instance methods that invalidate their receivers.

Our work builds on a specific form of TMs and OCs, whose popularity is growing, and we expect future languages to support some variation of these. Crucially, any language already designed with such TMs and OCs can also support our invariant protocol with minimal added complexity.

We demonstrated the applicability and simplicity of our approach with a GUI example. Our invariant protocol performs several orders of magnitude less checks than visible state semantics, and requires much less annotation than Spec#, (the system with the most comparable goals). In Section 4 we formalised our invariant protocol and in Appendix A we prove it sound. To stay parametric over the various existing type systems which provably enforce the properties we require for our proof (and much more), we do not formalise any specific type system.

One interesting avenue for future work would be to use invariants to encode pre and post conditions, as done by, where pre and post conditions are encoded as the invariants of the parameter and return types (respectively). Without good syntax sugar, such an approach could be quite verbose, however it would ensure that a methods precondition holds during
the entire execution of a method, and not just the beginning. In addition this could be more efficient than traditional runtime checking when the same argument is used in the invocations of methods with the same pre condition, as happens often in practice for recursive methods: where many parameters are simply parsed unmodified in recursive calls.

The language we presented here restricts the forms of invariant and capsule mutator methods; such strong restrictions allow for sound and efficient injection of invariant checks. These restrictions do not get in the way of writing invariants over immutable data, but the box pattern is required for verifying complex mutable data structures. We believe this pattern, although verbose, is simple and understandable. While it may be possible for a more complex and fragile type system to reduce the need for the pattern whilst still ensuring our desired semantics, we prioritize simplicity and generality.

In order to obtain safety, simplicity, and efficiency we traded some expressive power: the invariant method can only refer to immutable and encapsulated state. This means that while we can easily verify that a doubly linked list of immutable elements is correctly linked up, we cannot do the same for a doubly linked lists of mutable elements. Our approach does not prevent correctly implementing such data structures, but the invariant method would be unable to access the list’s nodes, since they would contain mut references to shared objects. In order to verify such data structures we could add a special kind of field which cannot be (transitively) accessed by invariants; such fields could freely refer to any object. We are however unsure if such complexity would be justified.

For an implementation of our work to be sound, catching exceptions like stack overflows or out of memory cannot be allowed in invariant methods, since they are not deterministically thrown. Currently L42 never allows catching them, however we could also write a (native) capability method (which can’t be used inside an invariant) that enables catching them. Another option worth exploring would be to make such exceptions deterministic, perhaps by giving invariants fixed stack and heap sizes.

Other directions that could be investigated to improve our work include the addition of syntax sugar to ease the burden of the box and the transform patterns; type modifier inference, and support for flexible ownership types.

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A Proof and Axioms

Axiomatic Type Properties

As previously discussed, instead of providing a concrete set of type rules, we provide a set of properties that the type system needs to respect. To express these properties, we first need some auxiliary definitions.

The encapsulated ROG of $l_0$ is composed of all the objects in the ROG of its immutable and capsule fields:

$$ l \in \text{erog}(\sigma, l_0) \text{ iff } \exists f, \Sigma caret(l_0). f \in \text{\{imm, capsule\}} \text{ and } l \in \text{rog}(\sigma, \sigma(l_0). f) $$

An object is \textit{mutable} in a $\sigma$ and $e$ if there is an occurrence of $l$ in $e$, that when seen as \textit{imm} makes the expression ill typed:

$$ \textit{mutable}(l, \sigma, e) \text{ iff for some } T = \text{imm} \Sigma caret(l) \text{ and } \Sigma caret[l] = e, $$

$$ \Sigma caret; x : T \vdash E[\Sigma caret[x]] : T' \text{ does not hold for any } T'.$$

Here we assume the usual Progress and Subject Reduction Base. Note that Subject Reduction Base only ensures properties about type checking, not invariant checking.

\textbf{Assumption 1 (Progress).} if $\Sigma caret; \emptyset \vdash e_0 : T_0$, and $e_0$ is not form $l$ or $error$, then $\sigma_0|e_0 \rightarrow \sigma_1|e_1$.

\textbf{Assumption 2 (Subject Reduction Base).} if $\Sigma caret; \emptyset \vdash e_0 : T_0$, $\sigma_0|e_0 \rightarrow \sigma_1|e_1$, then $\Sigma caret_1; \emptyset \vdash e_1 : T_1$.

If the result of a field access is \textit{mut}, the receiver is also \textit{mut}; field updates are only allowed on \textit{mut} receivers.

\textbf{Assumption 3 (Mut Field).}

(1) if $\Sigma; \Gamma \vdash e.f : \text{mut } _{}$ then $\Sigma; \Gamma \vdash e : \text{mut } _{}$ and

(2) if $\Sigma; \Gamma \vdash e_0.f = e_1 : T$ then $\Sigma; \Gamma \vdash e_0 : \text{mut } _{}$.

An object is not part of the ROG of its immutable or capsule field.

\textbf{Assumption 4 (Head Not Circular).} if $\Sigma caret; \Gamma \vdash l : T$, then $l \notin \text{erog}(\sigma, l)$.

In a well typed $\sigma$ and $e$, if mutableable $l_2$ is reachable through the $\text{erog}$ of $l_1$, and $l_1$ is reachable through the $\text{erog}$ of $l_0$, then all the paths connecting $l_0$ and $l_2$ pass through $l_1$; thus if we were to remove $l_1$ from the object graph, $l_0$ would no longer reach $l_2$.

\textbf{Assumption 5 (Capsule Tree).} If $\Sigma caret; \Gamma \vdash e : T$, $l_2 \in \text{erog}(\sigma, l_1)$, $l_1 \in \text{erog}(\sigma, l_0)$, and $\text{mutable}(l_2, \sigma, e)$ then $l_2 \notin \text{erog}(\sigma \setminus l_1, l_0)$.

Capsule Tree and Head Not Circular together imply that capsule fields section the object graph into a tree of nested ‘balloons’, where nodes are mutable encapsulated objects and edges are given by reachability between those objects in the original memory: if $l_2$ is in the encapsulated ROG of $l_1$, and $l_2$ is mutableable and reachable through $l_1$, then $l_2$ must be reachable by a capsule field. Thanks to Head Not Circular and $l_1 \in \text{erog}(\sigma, l_0)$ we can derive that $l_2 \notin \text{erog}(\sigma, l_1)$.

The execution of an expression with no \textit{mut} free variables is deterministic and does not mutate pre existing memory (and thus does not not perform I/O by mutating the pre existing $e$):

\textbf{Assumption 6 (Determinism).} if $\emptyset ; \Gamma \vdash e : T$, $\forall x (\Gamma(x) \neq \text{mut } _{})$, and $\sigma|e' \rightarrow^+ \sigma'|e''$ then $\sigma|e' \rightarrow^+ \sigma, _{}|e''$, where $e' = e[x_1 = l_1, \ldots, x_n = l_n]$ and $\Sigma caret; \emptyset \vdash e' : T$.

\footnote{This is not strictly true in L42, as 42 allows circular objects with \textit{fwd imm} fields, however such objects cannot have an invariant.}
For each try-catch, execution preserves the memory needed to continue the execution in case of an error (the memory visible outside of the try).

**Assumption 7 (Strong Exception Safety)**, if \( \Sigma, \sigma, \sigma' ; \emptyset \vdash \mathcal{E}[\text{try}^\sigma \{ e_0 \} \text{ catch } \{ e_1 \}] : T \) and \( \sigma, \sigma' \vdash \mathcal{E}[\text{try}^\sigma \{ e_0 \} \text{ catch } \{ e_1 \}] \rightarrow \sigma'' \vdash \mathcal{E}[\text{try}^\sigma \{ e_0 \} \text{ catch } \{ e_1 \}] \) then \( \sigma'' = \sigma, \_ \) and \( \Sigma, \sigma ; \emptyset \vdash e_1 : T \)

Note that our last well formedness rule requires update and mcall to introduce monitor expressions only over locations that are not preserved by try blocks. This can be achieved, since monitors are introduced around mutation operations (and new expression), and Strong Exception Safety ensures no mutation happens on preserved memory.

**Proof of Soundness**

It is hard to prove Soundness directly, so we first define a stronger property, called Stronger Soundness, and show that it is preserved during reductions by means of conventional Progress and Subject Reduction (Progress is one of our assumptions, while Subject Reduction relies heavily upon Subject Reduction Base). That is:

- Progress \& Subject Reduction \( \Rightarrow \) Stronger Soundness, and
- Stronger Soundness \( \Rightarrow \) Soundness.

**Stronger Soundness \( \Rightarrow \) Soundness**

Stronger Soundness depends on wellEncapsulated, monitored and OK:

\[
\text{wellEncapsulated}(\sigma, e, l_0) \iff \forall l \in \text{erog}(\sigma, l_0), \text{not mutable}(l, \sigma, e).
\]

The main idea is that an object is well encapsulated if its encapsulated state cannot be modified by \( e \).

An object is monitored if execution is currently inside of a monitor for that object, and the monitored expression \( e_1 \) does not contain \( l \) as a proper subexpression:

\[
\text{monitored}(e, l) \iff e = \mathcal{E}_v[l ; e_1] \quad \text{and either} \quad e_1 = l, \quad \text{or} \quad l \text{ is not inside } e_1.
\]

A monitored object is associated with an expression that can not observe it, but may reference its internal representation directly. In this way, we can safely modify its representation before checking its invariant.

The idea is that at the start the object will be valid and \( e_1 \) will reference \( l \); but during reduction, \( l \) will be used to modify the object; only after that moment, the object may become invalid.

**Define** \( \text{OK}(\sigma, e) \):

\[
\forall l \in \text{dom}(\sigma) \quad \text{either}
\]

1. garbage\((l, \sigma, e)\),
2. valid\((\sigma, l)\) \& wellEncapsulated\((\sigma, e, l)\), or
3. monitored\((e, l)\).

Finally, the system is in an OK state if all objects in memory, are either not (transitively) reachable from the expression (thus can be garbage collected), valid and encapsulated, or currently monitored.

**Theorem 2 (Stronger Soundness)**, if \( c : \text{Cap} ; \emptyset \vdash e_0 : T_0 \) and \( c \mapsto \text{Cap}[\_] | e_0 \rightarrow^* \sigma | e \), then \( \text{OK}(\sigma, e) \).

Starting from only the capability object, any well typed expression \( e_0 \) can be reduced in an arbitrary number of steps, and \( \text{OK} \) will always hold.

**Theorem 3.** Stronger Soundness \( \Rightarrow \) Soundness

**Proof.** By Stronger Soundness, each \( l \) in the current redex must be OK:

1. If \( l \) is garbage, it cannot be in the current redex, a contradiction.
2. If \( \text{valid}(\sigma, l) \), then \( l \) is valid, so thanks to Determinism no invalid object could be observed.

3. Otherwise, if \( \text{monitored}(e, l) \) then either:
   1. we are executing inside of \( e_1 \), thus the current redex is inside of a sub expression of the monitor that does not contain \( l \), a contradiction.
   2. or we are executing inside \( e_2 \): by our reduction rules, all monitor expressions start with \( e_2 = l.\text{invariant()} \), thus the first execution step of \( e_2 \) is trusted. Further execution steps are also trusted, since by well formedness the body of invariant methods only use \textit{this} (now replaced with \( l \)) to read fields.

In any of the possible cases above, \textit{Soundness} holds for \( l \), and so it holds for all redexes.

\textbf{Subject Reduction}

\textbf{Define} fieldGuarded(\( \sigma, e \)):
\[
\forall \mathcal{E} \text{ such that } e = \mathcal{E}[l.f] \text{ and } \Sigma^\sigma(l).f = \text{capsule } \_ , \text{ and } f \text{ inside } \Sigma^\sigma(l).\text{invariant}()
\]
   either \( \forall T, \forall C, \Sigma^\sigma; x : \text{mt } C \nvdash \mathcal{E}[x] : T \), or
   \( \mathcal{E} = \mathcal{E}''[c_1; \mathcal{E}'''; c_2] \) and \( l \) is contained exactly once in \( \mathcal{E}'' \).

That is, all \textit{mut} capsule field accesses are individually guarded by monitors. Note how we use \( C \) in \( x : \text{mt } C \) to guess the type of the accessed field, and that we use the full context \( \mathcal{E} \), instead of the evaluation context \( \mathcal{E}_v \), to refer to field accesses everywhere in the expression \( e \).

\textbf{Theorem 4 (Subject Reduction).} If \( \Sigma^\sigma; \emptyset \vdash e_0 : T_0, \sigma_0|e_0 \rightarrow \sigma_1|e_1, \text{OK}(\sigma_0, e_0) \) and \( \text{fieldGuarded}(\sigma_0, e_0) \) then \( \Sigma^\sigma; \emptyset \vdash e_1 : T_1, \text{OK}(\sigma_1, e_1) \) and \( \text{fieldGuarded}(\sigma_1, e_1) \)

\textbf{Theorem 5.} Progress + Subject Reduction \( \Rightarrow \) Stronger Soundness

\textit{Proof.} This proof proceeds by induction in the usual manner.

\textit{Base case:} At the start of execution, memory only contains \( c \) since \( c \) is defined to always be \textit{valid}, and has only \textit{mut} fields, it is trivially \textit{wellEncapsulated}, thus \textit{OK}(\( c \rightarrow \text{Cap}, e \)).

\textit{Induction:} By Progress, we always have another evaluation step to take, by Subject Reduction such a step will preserve \textit{OK}, and so by induction, \textit{OK} holds after any number of steps.

Note how for the proof garbage collectability is important: when the \textit{invariant}() method evaluates to \textit{false}, execution can continue only if the offending object is classified as \textit{garbage}.

\textbf{Exposer Instrumentation}

We first introduce a lemma derived from our well formedness criteria and the type system:

\textbf{Lemma 1 (Exposer Instrumentation).} If \( \sigma_0|e_0 \rightarrow \sigma_1|e_1 \) and \( \text{fieldGuarded}(\sigma_0, e_0) \) then \( \text{fieldGuarded}(\sigma_1, e_1) \).

\textit{Proof.} The only rule that can introduce a new field access is \textit{mcall}. In that case, \textbf{Exposer Instrumentation} holds by well formedness (all field accesses in methods are of the form \textit{this}.\( f \)), since \textit{mcall} inserts a monitor while invoking capsule mutator methods, and not field accesses themselves. If however the method is not a \textit{mut} method but still accesses a capsule field, by \textit{Mut Field} such a field access expression cannot be typed as \textit{mut} and so no monitor is needed.

Note that \textit{MONITOR EXIT} is fine because monitors are removed only when \( e_1 \) is a value.

\textbf{Proof of Subject Reduction}

Any reduction step can be obtained by exactly one application of the \textit{ctxv} rule and one other rule. Thus the proof can simply proceed by cases on the other applied rule.

By Subject Reduction Base and Exposer Instrumentation, \( \Sigma^\sigma; \emptyset \vdash e_1 : T_1 \) and \( \text{fieldGuarded}(\sigma_1, e_1) \). So we just need to proceed by cases on the reduction rule applied to verify that \textit{OK}(\( \sigma_1, e_1 \)) holds:

1. \textbf{UPDATE} \( \sigma|l.f = v \rightarrow \sigma'|e' \):
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- By \textbf{update} \( e' = \text{mut}(l; l).\text{invariant}() \), thus \textit{monitored}(e, l).
- Every \( l_i \) such that \( l \in \text{rog}(\sigma, l_i) \) will verify the same case as the former step:
  - If it was \textit{garbage}, clearly it still is.
  - If it was \textit{monitored}, it still is.
  - Otherwise it was \textit{valid} and \textit{wellEncapsulated}:
    * If \( l \in \text{rog}(\sigma, l_i) \) we have a contradiction since \textit{mutable}(l, \sigma, e), (by Mut Field)
    * Otherwise, by our well formedless criteria that \textit{invariant} only accesses \textit{imm} and \textit{capsule} fields, and by Determinism, it is clearly the case that \textit{valid} still holds;
      By Head Not Circular it cannot be the case that \( l \in \text{rog}(\sigma', l_i) \), and so \( l_i \) is still \textit{wellEncapsulated}.
- Every other \( l_0 \) is not in the reachable object graph of \( l \), thus it being \textit{OK} could not have been affected by this reduction step.

2. \textbf{(access)} \( \sigma[l.f \rightarrow \sigma[v]] \):
- If \( l \) was \textit{valid} and \textit{wellEncapsulated}:
  - If we have now broken \textit{wellEncapsulated}, we must have made something in its \textit{rog} \textit{mutable}. As we can only type \textit{capsule} fields as \textit{mut} and not \textit{imm} fields, by Mut Field we must have that \( f \) is \textit{capsule} and \( l.f \) is being typed as \textit{mut}.
    By \textit{fieldGuarded}(\sigma_0, e_0), the former step must have been inside a monitor \textit{mu}(l; \mathcal{E}_v[l.f]; e) and the \textit{l} under reduction was the only occurrence of \( l \). Since \( f \) is a capsule, we know that \( l \notin \text{rog}(\sigma, l) \) by Head Not Circular. Thus in our new step \( l \) is not \textit{inside} \( \mathcal{E}_v[v] \). Thus \( l \) must be \textit{monitored} and hence it is \textit{OK}.
  - Otherwise, \( l \) is still \textit{OK}
- Suppose some other \( l_0 \) was \textit{wellEncapsulated} and \textit{valid}:
  - If \( l \) was in the \textit{rog} of \( l_0 \), by \textit{Capsule Tree}, if \( l \) was in the \textit{rog} of \( l \), then \( v \) can only be reached from \( l_0 \) by passing through \( l \), and so we could not have made \( l_0 \) non \textit{wellEncapsulated}. In addition, since only things in the \textit{rog} can be referenced by \textit{invariant}, validity can not depend on \( l \), and by Determinism it is still the case that \( l_0 \) is \textit{valid}. And so we can’t have affected \( l_0 \) being \textit{OK}.
  - Otherwise, this reduction step could not have affected \( l_0 \), so \( l_0 \) is still \textit{OK}.
- Nothing that was \textit{garbage} could have been made reachable by this expression, since the only value we produced was \( v \) and it was reachable through \( l \) (and so could not have been \textit{garbage}), thus \( l \) is still \textit{OK}.
- As we don’t change any monitors here, nothing that was \textit{monitored} could have been made \textit{un-monitored}, and so it is still \textit{OK}.

3. \textbf{(mcall, try enter and try ok)}:
These reduction steps do not modify memory, the memory locations reachable inside of main expression, or any monitor expressions. Therefore it cannot have any effect on the \textit{garbage}, \textit{wellEncapsulated}, \textit{valid} (due to Determinism), or \textit{monitored} properties of any memory locations, thus \textit{OK} still holds.

4. \textbf{(new)} \( \sigma[\text{new} \ C(\tau) \rightarrow \sigma, l \rightarrow C(\tau)[\#(l; l).\text{invariant}()] \):
Clearly the newly created object, \( l \), is \textit{monitored}. As for \textit{mcall}, other objects and properties are not disturbed, and so \textit{OK} still holds.

5. \textbf{(monitor exit)} \( \sigma[\#(l; v; \text{true})] \rightarrow \sigma[v] \):
- As monitor expressions are not present in the original source code, it must have been introduced by \textit{update}, \textit{mcall}, or \textit{new}. In each case the 3\textsuperscript{rd} expression started of as \( l.\text{invariant}() \), and it has now (eventually) been reduced to \textit{true}, thus by Determinism \( l \) is \textit{valid}.
If the monitor was introduced by update, then \( v = l \). We must have had that \( l \) was well encapsulated before update was executed (since it can’t have been garbage and monitored, as update itself preserves this property and we haven’t modified memory in anyway, we must still have that \( l \) is wellEncapsulated. As \( l \) is valid and wellEncapsulated, it is OK.

If the monitor was introduced by mcall, then it was due to calling a capsule mutator method that mutated a field \( f \).

- A location that was garbage obviously still is, and so is also OK.
- No location that was valid could have been made invalid since this reduction rule performs no mutation of memory. If a location was wellEncapsulated before, the only way it could be non wellEncapsulated is if we somehow leaked a mut reference to something, but by our well-formedness rules, \( v \) cannot be typed as mut and so we can’t have affected wellEncapsulated, hence such thing is still OK.
- The only location that could have been made un monitored is \( l \) itself. By our well formedness criteria, \( l \) was only used to modify \( f \), and we have no parameters by which we could have made \( f \) non wellEncapsulated, since that would violate Capsule Tree. As nothing else in \( l \) was modified, and it must have been wellEncapsulated before the mcall, and so it still is. In addition since \( l \) is valid, it is OK.

- Otherwise the monitor was introduced by new. Since we require that capsule fields and imm fields are only initialised to capsule and imm expressions, by Capsule Tree, the resulting value, \( l \), must be wellEncapsulated, since \( l \) is also valid we have that \( l \) is OK.

6. (try error) \( \sigma, \sigma_0 | \text{try } \{ \text{error} \} \text{ catch } \{ e \} \rightarrow \sigma | e \):

By Strong Exception Safety, we know that \( \sigma_0 \) is garbage with respect to \( \mathcal{E}_v[e] \). By our well formedness criteria, no location inside \( \sigma \) could have been monitored. Since we don’t modify memory, everything in \( \sigma_0 \) is garbage and nothing inside \( \sigma \) was previously monitored, it is still clearly the case that everything in \( \sigma \) is OK.

## B The Hamster Example in Spec#

In this section we describe exactly why we chose to annotate the example from Section 1 in the way we did. For brevity, we will assume the default accessibility is public, whilst in both Spec# and C#, it is actually private.

### The Point Class

The typical way of writing a Point class in C# is as follows:

```csharp
class Point {
    double x, y;
    Point(double x, double y) { this.x = x; this.y = y; }
}
```

This works exactly as is in Spec#, however we have difficulty if we want to define equality of Points (see below).

### The Hamster Class

The Hamster class in C# would simply be:

```csharp
class Hamster {
    Point pos;
    Hamster(Point pos) { this.pos = pos; }
}
```
Though this is legal in Spec#, it is practically useless. Spec# has no way of knowing whether \texttt{pos} is valid or consistent. If \texttt{pos} is not known to be valid, one will be unable to pass it to almost any method, since by default methods implicitly require their receivers and arguments to be valid (compare this with our invariant protocol, which guarantees that any reachable object is valid). If \texttt{pos} is not known to be consistent, one will be unable to mutate it, by updating one of its fields or by passing it as an argument (or receiver) to a non \texttt{Pure} method. Though we don’t want \texttt{pos} to ever mutate, Spec# currently has no way of enforcing that an \texttt{instance} of a non immutable class is itself immutable \textsuperscript{23}, as such we will simply refrain from mutating it.

To enable Spec# to reason about \texttt{pos}’s validity, we will require that it be a \texttt{peer} of the enclosing \texttt{Hamster}; we can do this by annotating \texttt{pos} with \texttt{[Peer]}. Peers are objects that have the same owner, implying that whenever one is valid and/or consistent, the other one also is. This means that if we have a \texttt{Hamster}, we can use its \texttt{pos}, in the same ways as we could use the \texttt{Hamster}.

To simplify instantiation of \texttt{Hamsters}, their constructors will take unowned \texttt{Point}s. Spec# will then automatically make such \texttt{Point} a peer. This is achieved by taking a \texttt{[Captured]} \texttt{Point} in the constructor (note how similar this is to taking a \texttt{capsule Point}). Note that unlike our system, this prevents multiple \texttt{Hamsters} from sharing the same \texttt{Point}, unless both \texttt{Hamsters} have the same owner, if \texttt{Point} were immutable, there would be no such restriction.

With the aforementioned modifications, the \texttt{Hamster} becomes:

```csharp
class Hamster {
    [Peer] Point pos;
    Hamster([Captured] Point pos) { this.pos = pos; }
}
```

We don’t want \texttt{Point} to be an immutable/value type, however if it were, the original unannotated version would not have any problems.

### The Cage Class

The natural way to write this class in C#, if it had native support for class invariants like Spec#, would be:

```csharp
class Cage {
    Hamster h;
    List<Point> path;
    Cage(Hamster h, List<Point> path) { this.h = h; this.path = path; }
    invariant this.path.Contains(this.h.pos);
    void Move() {
        int index = this.path.IndexOf(this.h.pos);
        this.h.pos = this.path[index % this.path.Count];
    }
}
```

However for the above \texttt{invariant} to be admissible in Spec#, \texttt{this.path} and \texttt{this.h} must both be owned by \texttt{this}. In addition, the elements of \texttt{this.path} need to be owned by \texttt{this}, since \texttt{this.path.Contains} will read them. Note that \texttt{this.h.pos} also needs

\textsuperscript{23} There is a the describes a simple solution to this problem: assign ownership of the object to a special predefined ‘freezer’ object, which never gives up mutation permission \textsuperscript{[7]}, however this does not appear to have been implemented; this would provide similar flexibility to the TM system we use, which allows an initially mutable object to be promoted to immutable.
to be owned by \texttt{this}, however since \texttt{pos} is declared as \{Peer\}, if \texttt{this} owns \texttt{this.h}, it also owns \texttt{this.h.pos}. To fix the invariant, we will declare \texttt{h}, \texttt{path}, and the elements of \texttt{path} as \texttt{reps} (i.e. they are owned by the containing object). Finally, since \texttt{Move} modifies \texttt{this.h}, \texttt{this.h} needs to be made consistent, which requires that the owner (\texttt{this}) be made invalid; this can be achieved by using an \texttt{expose(this)} statement. \texttt{expose(this)}\{\texttt{body}\} marks \texttt{this} as invalid, executes \texttt{body}, checks that the invariant of \texttt{this} holds, and then marks \texttt{this} valid again. As we did with the \texttt{Hamster}, we will simply take unowned \texttt{h} and \texttt{path} values, however we also need the elements of \texttt{path} to be unowned; since \texttt{Spec#} has no [\texttt{ElementsCaptured}] annotation, we will require \texttt{path} to be unowned, and its elements (denoted by \texttt{Owner.ElementProxy(path)}) to be owned by the same owner as \texttt{path} (which is \texttt{null}).

```csharp
class Cage {
    [Rep] public Hamster h;
    [Rep, ElementsRep] List<Point> path;

    Cage([Captured] Hamster h, [Captured] List<Point> path)
    requires Owner.Same(Owner.ElementProxy(path), path);
    { this.h = h; this.path = path; }

    invariant this.path.Contains(this.h.pos);
    void Move() {
        int index = this.path.IndexOf(this.h.pos);
        expose(this){this.h.pos=this.path[index%this.path.Count]; }
    }
}
```

The above constructor now fails to verify, since Boogie is unconvinced that its precondition actually holds when we initialise \texttt{this.path}. This is because the constructor for \texttt{Object} (the default base class if none is provided) is not marked as [\texttt{Pure}]; since it is (implicitly) called upon entry to \texttt{Cage}'s constructor, Boogie has no idea as to what memory could've mutated, and so it doesn’t know whether the precondition still holds. The solution is to explicitly call it, but at the end of the constructor: \{\texttt{this.h = h; this.path = path; base();}\}.

The above \texttt{Cage} code however does not work, since \texttt{List} operations, such as \texttt{Contains} and \texttt{IndexOf}, will call the virtual \texttt{Object.Equals} method to compute equality of \texttt{Points}. However \texttt{Object.Equals} implements \textit{reference} equality, whereas we want \textit{value} equality.

\textbf{Defining Equality of Points}

The obvious solution in C# is to just override \texttt{Object.Equals} accordingly, and let dynamic dispatch handle the rest:

```csharp
class Point {
    .. // as before
    override bool Equals(Object? o) {
        Point? that = o as Point;
        return that !=null && this.x == that.x && this.y == that.y;
    }
}
```

However this fails in \texttt{Spec#} since \texttt{Object.Equals} is annotated with [\texttt{Pure} [\texttt{Reads(ReadsAttribute.Reads.Nothing)}]], and of course every overload of it must also satisfy this. The \texttt{Reads} annotations specifies that the method cannot read fields of any object, not even the receiver, this makes overloading the method useless.
We resorted to making our own Equal method. Since it is called in Cage's invariant, Spec# requires it to be annotated as [Pure], and either annotated with [Reads(ReadsAttribute.Reads.Nothing)] or [Reads(ReadsAttribute.Reads.Owned)] (the default, if the method is [Pure]). The latter annotation means it can only read fields of objects owned by the receiver of the method, so a [Pure] bool Equal(Point that) method can read the fields of this, but not the fields of that. Of course this would make the method unusable in Cage since the Points we are comparing equality against do not own each other. As such, the simplest solution is to pass the fields of the other point to the method:

```
[Pure] bool Equal(double x, double y) {
  return x == this.x && y == this.y;
}
```

Sadly however this mean we can no longer use List's Contains and IndexOf methods, rather we have to expand out their code manually; making these changes takes us to the version we presented in Section 1.

C More Case Studies

Family

The following test case was designed to produce a worst case in the number of invariant checks. We have a Family that (indirectly) contains a list of parents and children. The parents and children are of type Person. Both Family and Person have an invariant, the invariant of Family depends on its contained Persons.

```csharp
class Person {
  final String name;
  Int daysLived;
  final Int birthday;
  Person(String name, Int daysLived, Int birthday) { .. }
  mut method Void processDay(Int dayOfYear) {
    this.daysLived += 1;
    if (this.birthday == dayOfYear) {
      Console.print("Happy birthday "+this.name+"!"); }
  }
  read method Bool invariant() {
    return !this.name.equals("") && this.daysLived >= 0 &&
    this.birthday >= 0 && this.birthday < 365; }
}
class Family {
  static class Box {
    mut List<Person> parents;
    mut List<Person> children;
    Box(mut List<Person> parents, mut List<Person> children){..}
    mut method Void processDay(Int dayOfYear) {
      for(Person c : this.children) { c.processDay(dayOfYear); }
      for(Person p : this.parents) { p.processDay(dayOfYear); }
    }
  }
  capsule Box box;
  Family(capsule List<Person> ps,capsule List<Person> cs) {
    this.box = new Box(ps, cs); }
```
Note how we created a `Box` class to hold the `parents` and `children`. Thanks to this pattern, the invariant only needs to hold at the end of `Family.processDay`, after all the `parents` and `children` have been updated. Thus `Family.processDay` is atomic: it updates all its contained `Persons` together. Had we instead made the `parents` and `children` `capsule` fields of `Family`, the invariant would be required to also hold between modifying the two lists. This could cause problems if, for example, a child was updated before their parent.

We have a simple test case that calls `processDay` on a `Family 1,095 (3 × 365)` times.

```java
// 2 parents (one 32, the other 34), and no children
var fam = new Family(List.of(new Person("Bob", 11720, 40),
                            new Person("Alice", 12497, 87)), List.of());

for (Int day = 0; day < 365; day++) { // Run for 1 year
    fam.processDay(day);
}
for (Int day = 0; day < 365; day++) { // The next year
    fam.processDay(day);
    if (day == 45) {
        fam.addChild(new Person("Tim", 0, day));
    }
}
for (Int day = 0; r < 365; day++) { // The 3rd year
    fam.processDay(day);
    if (day == 340) {
        fam.addChild(new Person("Diana", 0, day));
    }
}
```

The idea is that everything we do with the `Family` is a mutation; the `fam.processDay` calls also mutate the contained `Persons`.

This is a worst case scenario for our approach compared to visible state semantics since it reduces our advantages: our approach avoids invariant checks when objects are not mutated but in this example most operations are mutations; similarly, our approach prevents the exponential explosion of nested invariant checks when deep object graphs are involved, but in this example the object graph of `fam` is very shallow.

We ran this test case using several different languages: L42 (using our protocol) performs 4,000 checks, D and Eiffel perform 7,995, and finally, Spec# performs only 1,104.

---

As happened in our GUI case study, see Section 7.
Sound Invariant Checking Using Type Modifiers and Object Capabilities.

Our protocol performs a single invariant check at the end of each constructor, processDay and addChild call (for both Person and Family).

The visible state semantics of both D and Eiffel perform additional invariant checks at the beginning of each call to processDay and addChild.

The results for Spec# are very interesting, since it performs less checks than L42. This is the case since processDay in Person just does a simple field update, which in Spec# do not invoke runtime invariant checks. Instead, Spec# tries to statically verify that the update cannot break the invariant; if it is unable to verify this, it requires that the update be wrapped in an expose block.

Spec# relies on the absence of arithmetic overflow, and performs runtime checks to ensure this as such the verifier concludes that the field increment in processDay cannot break the invariant. Spec# is able to avoid some invariant checks in this case by relying on all arithmetic operations performing runtime overflow checks; whereas integer arithmetic in L42 has the common wrap around semantics.

The annotations we had to add in the Spec# version were similar to our previous examples, however since the fields of Person all have immutable classes/types, we only needed to add the invariant itself. The Family class was similar to our Cage example (see section 1), however in order to implement the addChild method we were forced to do a shallow clone of the new child (this also caused a couple of extra runtime invariant checks). Unlike L42 however, we did not need to create a box to hold the parents and children fields, instead we wrapped the body of the Family.processDay method in an expose (this) block. In total we needed 16 annotations, worth a total of 45 tokens, this is worse than the code following our approach that we showed above, which has 14 annotations and 14 tokens.

Spec# Papers
Their are many published papers about the pack/unpack methodology used by Spec#.

Veriﬁcation of Object-Oriented Programs with Invariants: [5] this paper introduces their methodology. In their examples section (pages 41–47), they show how their methodology would work in a class hierarchy with Reader and ArrayReader classes. The former represents something that reads characters, whereas the latter is a concrete implementation that reads from an owned array. They extend this further with a Lexer that owns a Reader, which it uses to read characters and parse them into tokens. They also show an example of a FileList class that owns an array of filenames, and a DirFileList class that extends it with a stronger invariant. All of these examples can be represented in L42.

Their ArrayReader class has a relinquishReader method that ‘unpacks’ the ArrayReader and returns its owned array. The returned array can then be freely mutated and passed around by other code. However, afterwards the ArrayReader will be ‘invalid’, and so one can only call methods on it that do not require its invariant to hold. However, it may later be ‘packed’ again (after its invariant is checked). In contrast, our approach requires the invariant of all usable objects to hold. We can still relinquish the array, but at the cost of making the ArrayReader forever unreachable. This can be done by declaring relinquishReader as a capsule method, this works since our type

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25 Runtime checks are enabled by a compilation option; when they fail, unchecked exceptions are thrown.
26 The Spec# code is in the artifact.
27 Our encodings are in the artifact.
modifier system guarantees that the receiver of such a method is not aliased, and hence cannot be used again. Note that Spec# itself cannot represent the `relinquishReader` method at all, since it does not provide explicit pack and unpack operations, rather its `expose` statement performs both an unpack and a pack, thus we cannot unpack an `ArrayReader` without repacking it in the same method.

Their `DirFileList` example inherits from a `FileList` which has an invariant, and a final method, this is something their approach was specifically designed to handle. As L42 does not have traditional subclassing, we are unable to express this concept fully, but L42 does have code reuse via trait composition, in which case `DirFileList` can essentially copy and paste the methods from `FileList`, and they will automatically enforce the invariant of `DirFileList`.

Object Invariants in Dynamic Contexts: this paper shows how one can specify an invariant for a doubly linked list of `int`s (which is an immutable value type). Unlike our protocol however, it allows the invariant of `Node` to refer to sibling `Nodes` which are not owned/encapsulated by itself, but rather the enclosing `List`. Our protocol can verify such a linked list (since its elements are immutable), however we have to specify the invariant inside the `List` class. We do not see this as a problem, as the `Node` type is only supposed to be used as part of a `List`, thus this restriction does not impact users of `List`.

Friends Need a Bit More: Maintaining Invariants Over Shared State: this paper shows how one can verify invariants over interacting objects, where neither owns/contains the other. They have multiple examples which utilise the ‘subject/observer’ pattern, where a ‘subject’ has some state that an ‘observer’ wants to keep track of. In their `Subject/View` example, `Views` are created with references to `Subjects`, and copies of their state. When a `Subject`’s state is modified, it calls a method on its attached `Views`, notifying them of this update. The invariant is that a `View`’s copy of its `Subject`’s state is up to date. Their `Master/Clock` example is similar, a `Clock` contains a reference to a `Master`, and saves a copy of the `Master`’s time. The `Master` has a `Tick` method that increases its time, but unlike the `Subject/View` example, the `Clock` is not notified. The invariant is that the `Clock`’s time is never ahead of its `Master`’s. Our protocol is unable to verify these interactions, because the interacting objects are not immutable or encapsulated by each other.

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28 Our protocol allows for encoding this example, but to express the invariant we would need to use reference equality, which the L42 language does not support.