

Space-Efficient Latent Contracts

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Abstract. Standard higher-order contract monitoring breaks tail recursion and leads to space leaks that can change a program’s asymptotic complexity; space-efficiency restores tail recursion and bounds the amount of space used by contracts. Space-efficient contract monitoring for contracts enforcing simple type disciplines (a/k/a gradual typing) is well studied. Prior work establishes a space-efficient semantics for manifest contracts without dependency [10]; we adapt that work to a latent calculus with dependency. We guarantee space efficiency when no dependency is used; we cannot *generally* guarantee space efficiency when dependency is used, but instead offer a framework for making such programs space efficient on a case-by-case basis.

1 Introduction

Findler and Felleisen [6] brought design-by-contract [16] into the higher-order world, allowing programmers to write pre- and post-conditions on functions to be checked at runtime. Pre- and post-conditions are easy in first-order languages, where it’s very clear who is to blame when a contract is violated: if the pre-condition fails, blame the caller; if the post-condition fails, blame the callee. In higher-order languages, however, it’s harder to tell who calls whom! Who should be to blame when a pre-condition on a higher-order function fails? For example, consider the following contract:

$$(\text{pred}(\lambda x:\text{Int}. x > 0) \mapsto \text{pred}(\lambda y:\text{Int}. y \geq 0)) \mapsto \text{pred}(\lambda z:\text{Int}. z \bmod 2 = 0)$$

This contract applies to a function (call it f , with type $(\text{Int} \rightarrow \text{Int}) \rightarrow \text{Int}$) that takes another function (call it g , with type $\text{Int} \rightarrow \text{Int}$) as input. The contract says that g will only be called with positives and only return naturals; f must return an even number. If f returns an odd number, f is to blame; if g returns a negative number, then it, too is to blame. But what if g is *called* with a non-positive number, say, -1 ? Who is to blame then? Findler and Felleisen’s insight was that even in a higher-order setting, there are only two parties to blame. Here, g was given to f , so any bad values given to g here are due to some nefarious action on f ’s part—blame f ! That is, the higher-order case generalizes pre- and post-conditions so that the negative positions of a contract all blame the caller while the positive positions all blame the callee.

Dependent contracts—where the codomain contract can refer to the function’s argument—are particularly useful. For example, the square root function,

`sqrt`, satisfies the contract: $x:\text{pred}(\lambda y:\text{Real}. y \geq 0) \mapsto \text{pred}(\lambda z:\text{Real}. \text{abs}(x - z * z) < \epsilon)$ That is, `sqrt` takes a non-negative real, x , and returns a non-negative real z that’s within ϵ of the square root of x . (The dependent variable x is bound in the codomain; the variable y is local to the domain predicate.)

1.1 Contracts leak space

While implementations of contracts have proven quite successful (particularly so in Racket [8,18]), there is a problem: contracts leak space. Why?

The default implementation of contracts works by wrapping a function in a *function proxy*. For example, to check that $f = \lambda x:\text{Int}. x + 1$ satisfies the contract $C = \text{pred}(\lambda z:\text{Int}. z \bmod 2 = 0) \mapsto \text{pred}(\lambda z:\text{Int}. z \bmod 2 = 0)$, we monitor the function by wrapping it in a function proxy $\text{mon}^l(C, f)$. When this proxy is called with an input v , we first check that v satisfies C ’s domain contract (i.e., that v is even), then we run f on v to get some result v' , and then check that v' satisfies C ’s codomain contract (that the result is even). Here the contract will always fail blaming l : one of v and v' will always be odd.

Contracts leak space in two ways. First, there is no bound on the number of function proxies that can appear on a given function. More grievously, contracts break tail recursion. To demonstrate the issue with with tail calls, we’ll use the simplest example of mutual recursion: detecting parity.

```
let odd  =  $\lambda x:\text{Int}. \text{if}(x = 0) \text{false} (\text{even}(x - 1))$ 
and even =  $\lambda x:\text{Int}. \text{if}(x = 0) \text{true}  (\text{odd}(x - 1))$ 
```

Functional programmers will expect this program to run in constant space, because it is *tail recursive*. Adding a contract breaks the tail recursion. If we add a contract to `odd` and call `odd 5`, what contract checks accumulate (Fig. 1)?¹ Notice how the checks accumulate in the codomain? Even though the mutually recursive calls to `even` and `odd` are syntactically tail calls, we can’t bound the number of codomain checks that occur. That is, we can’t bound the size of the stack, and tail recursion is broken! Even though there’s only one function proxy on `odd`, our contracts create a space leak.

1.2 Overview and contributions

Space efficiency for gradual types [23] (a/k/a contracts constrained to type tests) is well studied [13,14,24,9,22]; Greenberg [10] developed a space-efficient semantics for general, non-dependent contracts. He used a manifest calculus, conflating contracts and types; however, contracts are typically implemented in latent calculi, where contracts are distinct from whatever types may exist. Greenberg “believe[s] it would be easy to design a latent version of eidetic λ_H , following the translations in Greenberg, Pierce, and Weirich (GPW)” [11]; in this paper,

¹ Readers may observe that the contract betrays a deeper knowledge of numbers than the functions themselves. We offer this example as minimal, not naturally occurring.

```

let odd  = monlodd(x:pred(λx:Int. x ≥ 0) ↦ pred(λb:Bool. b or (x mod 2 = 0)),
                  λx:Int. if (x = 0) false (even (x - 1)))
and even = λx:Int. if (x = 0) true  (odd (x - 1))

      odd 5
→C* monlodd(pred[x→5](...), even 4)
→C* monlodd(pred[x→5](...), monlodd(pred[x→3](...),
      odd monlodd(pred(λx:Int. x ≥ 0), 3)))
→C* monlodd(pred[x→5](...), monlodd(pred[x→3](...), even 2))
→C* monlodd(pred[x→5](...), monlodd(pred[x→3](...), monlodd(pred[x→1](...),
      odd monlodd(pred(λx:Int. x ≥ 0), 1))))
→C* monlodd(pred[x→5](...), monlodd(pred[x→3](...), monlodd(pred[x→1](...), even 0)))

```

Fig. 1. Contracts break tail recursion

we show that belief to be well founded by giving a space-efficient semantics for a (dependent!) variant of contract PCF (CPCF) [3,4].

The rest of this paper discusses a formulation of contracts that enjoys sound space efficiency; that is, where we slightly change the implementation of contracts so that (a) programs are observationally equivalent to the standard semantics, but (b) contracts consume a bounded amount of space. In this paper, we’ve omitted some of the more detailed examples and motivation—we refer curious readers to Greenberg [10], though we intend the paper to be self-contained.

We follow Greenberg’s general structure, defining two forms of dependent CPCF: CPCF_C is the *classic* semantics; CPCF_E follows the space-efficient eideitic semantics. We are able to prove space efficiency without dependency, bounding the amount of space consumed by contracts; we are unable to prove space efficiency in general with dependency, but instead offer a framework that allows for dependent contracts to be *made* space efficient.

We offer two primary contributions: adapting Greenberg’s work to a latent calculus and extending the possibility of space efficiency to dependent contracts.

There are some other, smaller, contributions as well. First, adding in nontermination moves beyond Greenberg’s strongly normalizing calculi, showing that the POPL 2015 paper’s result isn’t an artifact of strong normalization (where we can, in theory, bound the size of the any term’s evaluation in advance, not just contracts). Second, the simpler type system here makes it clear which type system invariants are necessary for space-efficiency and which are bookkeeping for proving that the more complicated manifest type system is sound. Third, by separating contracts and types, we can give tighter space bounds. Finally, we explore how space efficiency can be attained in dependent contracts. While we can’t give a guarantee for dependent contracts, we show that it’s possible to achieve and discuss different ways to do so.

Types	$B ::= \text{Bool} \mid \text{Int} \mid \dots$
	$T ::= B \mid T_1 \rightarrow T_2$
Terms	$e ::= x \mid k \mid e_1 \text{ op } e_2 \mid e_1 e_2 \mid \lambda x:T. e \mid \mu(x:T). e \mid \text{if } e_1 e_2 e_3 \mid$ $\text{err}^l \mid \text{mon}^l(C, e) \mid \text{mon}(c, e)$
	$\text{op} ::= \text{add1} \mid \text{sub1} \mid \dots$
	$k ::= \text{true} \mid \text{false} \mid 0 \mid 1 \mid \dots$
	$w ::= v \mid \text{err}^l$
	$v ::= k \mid \lambda x:T. e \mid \text{mon}^l(x:C_1 \mapsto C_2, v) \mid \text{mon}(x:c_1 \mapsto c_2, \lambda x:T. e)$
	$C ::= \text{pred}_\sigma(e) \mid x:C_1 \mapsto C_2$
	$c ::= r \mid x:c_1 \mapsto c_2$
	$r ::= \text{nil} \mid \text{pred}_\sigma^l(e); r$

Fig. 2. Syntax of classic and space-efficient CPCF

2 Classic and space-efficient Contract PCF

We present classic and space-efficient CPCF as separate calculi sharing syntax and some typing rules (Fig. 2 and Fig. 3), and a single, parameterized operational semantics with some rules held completely in common (omitted to save space) and others specialized to each system (Fig. 4). The formal presentation is modal, with two modes: C for classic and E for space-efficient. While much is shared between the two modes—types, T ; the core syntax of expressions, e ; most of the typing rules—we use colors to highlight parts that belong to only one system. Classic CPCF is typeset in salmon while space-efficient CPCF is in periwinkle.

2.1 Contract PCF (CPCF)

Plain CPCF is an extension of Plotkin’s 1977 PCF [17], developed first by Dimoulas and Felleisen [3,4] (our syntax is in Fig. 2). It is a simply typed language with recursion. The typing rules are straightforward (Fig. 3). The operational semantics for the generic fragment also uses conventional rules (omitted to save space). Dimoulas and Felleisen use evaluation contexts to offer a concise description of their system; we write out our relation in full, giving congruence rules (E*L, E*R, EIF) and error propagating rules (E*RAISE) explicitly—we will need to restrict congruence for casts, and our methods are more transparent written with explicit congruence rules than using the subtly nested evaluation contexts of Herman et al. [13,14].

Contracts are CPCF’s distinguishing feature. Contracts, C , are installed via monitors, written $\text{mon}^l(C, e)$; such a monitor says “ensure that e satisfies the contract C ; if not, the blame lies with label l ”. Monitors only apply to appropriate types (TMON). There are two kinds of contracts in CPCF: *predicate contracts* over base type, written $\text{pred}_\sigma(e)$, and *function contracts*, written $x:C_1 \mapsto C_2$.

Predicate contracts $\text{pred}_\sigma(e)$ have two parts: a predicate on base types, e , which identifies which values satisfy the contract; and a closing substitution σ

which keeps track of values substituted into the contract. For example, if ι is the identity substitution mapping variables to themselves:

- $\text{pred}_\iota(\lambda x:\text{Int}. x > 0)$ identifies the positives;
- $\text{pred}_\iota(\lambda x:\text{Int}. x > y)$ identifies numbers greater than an unspecified number y ; and,
- $\text{pred}_{[y \mapsto 47]}(\lambda x:\text{Int}. x > y)$ identifies numbers greater than 47.

When the closing substitution σ is the identity mapping ι , we write $\text{pred}(e)$ instead of $\text{pred}_\iota(e)$. In CPCF_C , closing substitutions will map each variable to either (a) itself or (b) a value. Substitution into contracts is a non-issue without dependency: each contract is just closed. Having introduced dependency, we use explicit closing substitutions rather than direct substitution for three reasons: first, it simplifies our space efficiency proof for simple contracts (Sec. 4.1); second, explicitness lets us distinguish the contract $\text{pred}_{[x \mapsto 0]}(\lambda x:\text{Int}. x = 0)$ from $\text{pred}_{[x \mapsto 0]}(\lambda x:\text{Int}. 0 = 0)$; third, it emphasizes that contracts are just another form of closure. Predicates are solely over *base types*, not functions.

Function contracts $x:C_1 \mapsto C_2$ are satisfied by functions satisfying their parts: functions whose inputs all satisfy C_1 and whose outputs all satisfy C_2 . Function contracts are dependent: the codomain contract C_2 can refer back to the input to the function. For example, the contract $x:\text{pred}(\lambda z:\text{Int}. z > 0) \mapsto \text{pred}(\lambda y:\text{Int}. y > x)$ is satisfied by increasing functions on the positives. Note that x is bound in the codomain, but z is not.² When function contracts aren't dependent, we omit the binder at the front, e.g., $\text{pred}(\lambda x:\text{Int}. x > 0) \mapsto \text{pred}(\lambda x:\text{Int}. x > 0)$ means operators on positives. We check that contracts are satisfied at runtime.

We use explicit, delayed substitutions to keep track of which values are substituted into predicate contracts. To help with our proof of space efficiency, we don't track variables that don't appear in the predicate:

$$\text{pred}_\sigma(e)[v/x] = \begin{cases} \text{pred}_{\sigma[x \mapsto v]}(e) & x \in \text{fv}(\sigma(e)) \\ \text{pred}_\sigma(e) & \text{otherwise} \end{cases}$$

Alpha equivalence allows us to give fresh names to variables in the domain of σ by consistently renaming those variables inside of the predicate e . Only holding on to substitutions that close up free variables in e is a way of modeling closures. A dependent predicate closes over some finite number of variables; a compiled representation would generate a closure with a corresponding number of slots in the closing environment. Restricting substitutions to exactly those variables appearing free in the predicate serves another purpose: we can easily recover space-efficiency bounds for programs without dependent contracts (Sec. 4.1).

2.2 Classic Contract PCF (CPCF_C)

Classic CPCF gives a straightforward semantics to contracts (Fig. 4), largely following the seminal work by Findler and Felleisen [6]. To check a predicate

² Concrete syntax for such predicates can be written much more nicely, but we ignore such concerns here.

Typing rules $\boxed{\Gamma \vdash e : T}$

$$\frac{x:T \in \Gamma}{\Gamma \vdash x : T} \text{TVAR} \quad \frac{}{\Gamma \vdash k : \text{ty}(k)} \text{TCONST} \quad \frac{}{\Gamma \vdash \text{err}^l : T} \text{TBLAME}$$

$$\frac{\Gamma, x:T_1 \vdash e_{12} : T_2}{\Gamma \vdash \lambda x:T_1. e_{12} : T_1 \rightarrow T_2} \text{TAbs} \quad \frac{\Gamma, x:T \vdash e : T}{\Gamma \vdash \mu(x:T). e : T} \text{TREC}$$

$$\frac{\text{ty}(op) = T_1 \rightarrow T_2 \rightarrow T \quad \Gamma \vdash e_1 : T_1 \quad \Gamma \vdash e_2 : T_2}{\Gamma \vdash e_1 \text{ op } e_2 : T} \text{TOP} \quad \frac{\Gamma \vdash e_1 : T_1 \rightarrow T_2 \quad \Gamma \vdash e_2 : T_1}{\Gamma \vdash e_1 e_2 : T_2} \text{TAPP}$$

$$\frac{\Gamma \vdash e_1 : \text{Bool} \quad \Gamma \vdash e_2 : T \quad \Gamma \vdash e_3 : T}{\Gamma \vdash \text{if } e_1 e_2 e_3 : T} \text{TIF}$$

$$\frac{\Gamma \vdash e : T \quad \Gamma \vdash C : T}{\Gamma \vdash \text{mon}^l(C, e) : T} \text{TMON} \quad \frac{\Gamma \vdash e : T \quad \Gamma \vdash c : T}{\Gamma \vdash \text{mon}(c, e) : T} \text{TMONC}$$

Contract typing $\boxed{\Gamma \vdash C : T}$ $\boxed{\Gamma \vdash c : T}$

$$\frac{\Gamma, \Gamma' \vdash e : B \rightarrow \text{Bool} \quad \Gamma' \vdash \sigma}{\Gamma \vdash \text{pred}_\sigma(e) : B} \text{TPRED} \quad \frac{\Gamma \vdash C_1 : T_1 \quad \Gamma, x:T_1 \vdash C_2 : T_2}{\Gamma \vdash x:C_1 \mapsto C_2 : T_1 \rightarrow T_2} \text{TFUN}$$

$$\frac{}{\Gamma \vdash \text{nil} : B} \text{TCNIL} \quad \frac{\Gamma \vdash \text{pred}_\sigma(e) : B \quad \Gamma \vdash r : B}{\Gamma \vdash \text{pred}_\sigma^l(e); r : B} \text{TCPRED}$$

$$\frac{\Gamma \vdash c_1 : T_1 \quad \Gamma, x:T_1 \vdash c_2 : T_2}{\Gamma \vdash x:c_1 \mapsto c_2 : T_1 \rightarrow T_2} \text{TCFUN}$$

Closing substitutions $\boxed{\Gamma \vdash \sigma}$

$$\frac{}{\emptyset \vdash \iota} \text{TID} \quad \frac{\Gamma \vdash \sigma \quad x:T \vdash v : T}{\Gamma, x:T \vdash \sigma[x \mapsto v]} \text{TMAP}$$

Fig. 3. Typing rules of classic and space-efficient CPCF

$$\begin{array}{c}
\frac{}{\text{mon}^l(\text{pred}_\sigma(e_1), v_2) \rightarrow_C \text{if } (\sigma(e_1) \ v_2) \ v_2 \ \text{err}^l} \text{EMONPRED} \\
\frac{}{\text{mon}^l(x: C_1 \mapsto C_2, v_1) \ v_2 \rightarrow_C \text{mon}^l(C_2[v_2/x], v_1 \ \text{mon}^l(C_1, v_2))} \text{EMONAPP} \\
\frac{e \rightarrow_C e'}{\text{mon}^l(C, e) \rightarrow_C \text{mon}^l(C, e')} \text{EMON} \quad \frac{}{\text{mon}^l(C, \text{err}^{l'}) \rightarrow_C \text{err}^{l'}} \text{EMONRAISE} \\
\frac{}{\text{mon}^l(C, e) \rightarrow_E \text{mon}(\text{label}^l(C), e)} \text{EMONLABEL} \\
\frac{}{\text{mon}(\text{nil}, v_1) \rightarrow_E v_1} \text{EMONCNIL} \\
\frac{}{\text{mon}(\text{pred}_\sigma^l(e); r, v_1) \rightarrow_E \text{if } (\sigma(e) \ v_1) \ \text{mon}(r, v_1) \ \text{err}^l} \text{EMONCPRED} \\
\frac{}{\text{mon}(x: c_1 \mapsto c_2, v_1) \ v_2 \rightarrow_E \text{mon}(c_2[v_2/x], v_1 \ \text{mon}(c_1, v_2))} \text{EMONCAPP} \\
\frac{e \neq \text{mon}(c', e'') \quad e \rightarrow_E e'}{\text{mon}(c, e) \rightarrow_E \text{mon}(c, e')} \text{EMONC} \quad \frac{}{\text{mon}(c, \text{err}^l) \rightarrow_E \text{err}^l} \text{EMONCRAISE} \\
\frac{}{\text{mon}(c_2, \text{mon}(c_1, e)) \rightarrow_E \text{mon}(\text{join}(c_1, c_2), e)} \text{EMONCJOIN}
\end{array}$$

Fig. 4. Operational semantics of classic and space-efficient CPCF

contract, we simply test it (EMONPRED), returning either the value or an appropriately labeled error. Function contracts are deferred: $\text{mon}^l(x: C_1 \mapsto C_2, v)$ is a *value*, called a *function proxy*. When a function proxy is applied, it unwraps the proxy, monitoring the argument with the domain contract, running the function, and then monitoring the return value with the codomain contract (EMONAPP).

Our semantics may seem to be *lax*, where no monitor is applied to dependent uses of the argument in the codomain monitor [11]. In fact, it is agnostic: we could be *picky* by requiring that function contract monitors $\text{mon}^l(x: C_1 \mapsto C_2, e)$ have the substitution $[x \mapsto \text{mon}^l(C_1, x)]$ throughout C_2 ; we could be *indy* by having $[x \mapsto \text{mon}^l(C_1, x)]$ throughout C_2 [4]. We default to a lax rule to make our proof of soundness easier, but we'll have as a corollary that classic and space-efficient semantics yield the same result regardless of what the closing substitutions do in the codomain (Sec. 3).

Standard congruence rules allow for evaluation inside of monitors (EMON) and the propagation of errors (EMONRAISE).

$$\begin{aligned}
& \text{label}^l(\text{pred}_\sigma(e_1)) = \text{pred}_\sigma^l(e_1) \\
& \text{label}^l(x:C_1 \mapsto C_2) = x:\text{label}^l(C_1) \mapsto \text{label}^l(C_2) \\
& \text{join}(\text{nil}, r_2) = r_2 \\
& \text{join}(\text{pred}_\sigma^l(e); r_1, r_2) = \text{pred}_\sigma^l(e); \text{drop}(\text{join}(r_1, r_2), \text{pred}_\sigma(e)) \\
& \text{join}(x:c_{11} \mapsto c_{12}, x:c_{21} \mapsto c_{22}) = x:\text{join}(c_{21}, c_{11}) \mapsto \text{join}(\text{wrap}(c_{12}, x, c_{22}), c_{22}) \\
& \text{drop}(\text{nil}, \text{pred}_\sigma(e)) = \text{nil} \\
& \text{drop}(\text{pred}_{\sigma_2}^l(e_2); r, \text{pred}_{\sigma_1}^l(e_1)) = \\
& \quad \begin{cases} \text{drop}(r, \text{pred}_{\sigma_2}(e_2)) & \text{pred}_{\sigma_1}(e_1) \supset \text{pred}_{\sigma_2}(e_2) \\ \text{pred}_{\sigma_2}^l(e_2); \text{drop}(r, \text{pred}_{\sigma_1}(e_1)) & \text{pred}_{\sigma_1}(e_1) \not\supset \text{pred}_{\sigma_2}(e_2) \end{cases} \\
& \text{wrap}(\text{pred}_\sigma^l(e), x, c) = \begin{cases} \text{pred}_\sigma^l(e) & x \notin \text{fv}(\sigma(e)) \\ \text{pred}_{\sigma[x \mapsto \text{mon}(\text{join}(c', c), e)]}^l(e) & \sigma(x) = \text{mon}(c', e) \\ \text{pred}_{\sigma[x \mapsto \text{mon}(c, \sigma(x))]}^l(e) & \text{otherwise} \end{cases} \\
& \text{wrap}(\text{nil}, x, c) = \text{nil} \\
& \text{wrap}(\text{pred}_\sigma^l(e); r, x, c) = \text{wrap}(\text{pred}_\sigma^l(e), x, c); \text{wrap}(r, x, c) \\
& \text{wrap}(y:c_1 \mapsto c_2, x, c) = y:\text{wrap}(c_1, x, c) \mapsto \text{wrap}(c_2, x, c)
\end{aligned}$$

Fig. 5. Contract labeling and predicate stack management

2.3 Space-efficient Contract PCF (CPCF_E)

How can we recover tail calls in CPCF? CPCF_C will happily wrap arbitrarily many function proxies around a value, and there’s no bound on the number of codomain contract checks that can accumulate. The key idea is *joining* contracts. We’ll make two changes to the language: we’ll bound function proxies so each function has at most one, and we’ll bound stacks to avoid redundant checking. We ultimately show that contracts without dependency use constant space, but that the story for dependent functions is more complex (Sec. 4).

Fortuitously, our notion of join solves both of our problems, working identically for both simple and dependent contracts. To ensure a function value can have only one proxy, we change the semantics of monitoring: when monitoring a proxied value, we join the new monitor and the old one. To bound the size of stacks contract checks, we join pending contracts to avoid redundancy.

The join operation works on *labeled contracts*, which (a) move the label from the monitor into the contract and (b) allow us to keep track of many predicates at once (Fig. 5). Concretely, labeled contracts use the metavariable c (as opposed to C), comprising function contracts as usual ($x:c_1 \mapsto c_2$) and *predicate stacks*, r (Fig. 2). A predicate stack r is a list of *labeled predicates* $\text{pred}^l(e)$, where nil is the empty stack.

The join operation takes two labeled contracts and combines them, eliminating redundant contracts as it goes. To join a new and an old predicate stack, we keep new contracts and eliminate redundant old ones; only more “re-

cent” contracts are kept. Joining functions works contravariantly, being careful to maintain correct substitution behavior using `wrap`.

Finally, we establish what we mean by “redundant” using *predicate implication*: when does one contract imply another?

Definition 1 (Predicate implication). *Let $\text{pred}_{\sigma_1}(e_1) \supset \text{pred}_{\sigma_2}(e_2)$ be a relation on predicates such that:*

- (**Reflexivity**) *If $\emptyset \vdash \text{pred}_{\sigma}(e) : B$ then $\text{pred}_{\sigma}(e) \supset \text{pred}_{\sigma}(e)$.*
- (**Transitivity**) *If $\text{pred}_{\sigma_1}(e_1) \supset \text{pred}_{\sigma_2}(e_2)$ and $\text{pred}_{\sigma_2}(e_2) \supset \text{pred}_{\sigma_3}(e_3)$, then $\text{pred}_{\sigma_1}(e_1) \supset \text{pred}_{\sigma_3}(e_3)$.*
- (**Substitutivity**) *When $\Gamma_{i_1}, x : T', \Gamma_{i_2} \vdash \text{pred}_{\sigma_i}(e_i) : T$ and $\emptyset \vdash v : T'$, if $\text{pred}_{\sigma_1}(e_1) \supset \text{pred}_{\sigma_2}(e_2)$ then $\text{pred}_{\sigma_1}(e_1)[v/x] \supset \text{pred}_{\sigma_2}(e_2)[v/x]$.*
- (**Adequacy**) *If $\emptyset \vdash \text{pred}_{\sigma_1}(e_1) : T$ and $\text{pred}_{\sigma_1}(e_1) \supset \text{pred}_{\sigma_2}(e_2)$ then $\forall k \in \mathcal{K}_B. \sigma_1(e_1) k \rightarrow_m \text{true}$ implies $\sigma_2(e_2) k \rightarrow_m \text{true}$.*
- (**Decidability**) *For all $\emptyset \vdash \text{pred}_{\sigma_1}(e_1) : B$ and $\emptyset \vdash \text{pred}_{\sigma_2}(e_2) : B$, it is decidable whether $\text{pred}_{\sigma_1}(e_1) \supset \text{pred}_{\sigma_2}(e_2)$ or $\text{pred}_{\sigma_1}(e_1) \not\supset \text{pred}_{\sigma_2}(e_2)$.*

The entire development of space-efficiency is parameterized over this implication relation, \supset , characterizing when one first-order contract subsumes another. We write $\not\supset$ for the negation of \supset . The \supset relation is a *total pre-order* (a/k/a a *preference relation*)—it would be a total order, but it may not necessarily enjoy anti-symmetry. For example, we could have $\text{pred}_l(\lambda x:\text{Int}. x \geq 0) \supset \text{pred}_l(\lambda x:\text{Int}. x + 1 > 0)$ and vice versa, even though the two predicates aren’t equal. You can also view \supset as a total order *up-to contextual equivalence*.

There is at least one workable implication relation: syntactic equality. We say $\text{pred}_{\sigma_1}(e_1) \supset \text{pred}_{\sigma_2}(e_2)$ iff $e_1 = e_2$ and $\sigma_1 = \sigma_2$. Since we’ve been careful to store only those values that are actually referenced in the closure of the predicate, the steps to determine these equalities are finite and computable at runtime. For example, suppose we wish to show that $\text{pred}_{[y \mapsto 47]}(\lambda x:\text{Int}. x > y) \supset \text{pred}_{[y \mapsto 47]}(\lambda x:\text{Int}. x > y)$. The code part—the predicate $\lambda x:\text{Int}. x > y$ —is the same; an implementation might observe that the function pointers are equal. The environment has only one slot, for y , with the value 47 in it; an implementation might compare the two environments slot-by-slot.

Substitution in the codomain: lax, picky, and indy We extend Greenberg’s notion of join to account for dependency with a new function, `wrap`. Greenberg, Pierce, and Weirich identified two variants of latent contracts in the literature, differing in their treatment of the dependent substitution of arguments in the codomain: *picky*, where we monitor the value substituted in the codomain with the domain contract; and *lax*, where the actual parameter value substituted into the codomain is unmonitored [11]. There is a third variant, *indy*, which applies a monitor to the argument value but uses a different blame label [4]. These different models of substitution all exhibit different behavior for *abusive* contracts, where the codomain contract violates the domain contract.

There is another source of substitutions in the codomain: multiple function proxies. How do the monitors unfold when we have two function proxies? In the

$$\begin{aligned}
C_1 &= f:(\text{pred}(\lambda x:\text{Int}. x > 0) \mapsto \text{pred}(\lambda x:\text{Int}. x > 0)) \mapsto \text{pred}(\lambda x:\text{Int}. x > 0) \\
C_2 &= f:(\text{pred}(\lambda x:\text{Int}. \text{true}) \mapsto \text{pred}(\lambda x:\text{Int}. \text{true})) \mapsto \text{pred}(\lambda x:\text{Int}. f \ 0 = 0) \\
&\text{mon}^1(C_1, \text{mon}^2(C_2, \lambda f:(\text{Int} \rightarrow \text{Int}). f \ 5)) (\lambda x:\text{Int}. x) \\
&\rightarrow_{\mathcal{C}} \text{mon}^1(C_{12}[(\lambda x:\text{Int}. x)/f], \\
&\quad \text{mon}^2(C_2, \lambda f:(\text{Int} \rightarrow \text{Int}). f \ 5) \text{mon}^1(C_{11}, (\lambda x:\text{Int}. x))) \\
&\rightarrow_{\mathcal{C}}^* \text{mon}^1(C_{12}[(\lambda x:\text{Int}. x)/f], \text{mon}^2(C_{22}[\text{mon}^1(C_{11}, \lambda x:\text{Int}. x)/f], 5)) \\
&\rightarrow_{\mathcal{C}} \text{mon}^1(C_{12}[(\lambda x:\text{Int}. x)/f], \\
&\quad \text{if } ((\lambda x:\text{Int}. \text{mon}^1(C_{11}, \lambda x:\text{Int}. x) \ 0 = 0) \ 5) \ 5 \ \text{err}^{l_2}) \\
&\rightarrow_{\mathcal{C}} \text{mon}^1(C_{12}[(\lambda x:\text{Int}. x)/f], \text{if } (\text{mon}^1(C_{11}, \lambda x:\text{Int}. x) \ 0 = 0) \ 5 \ \text{err}^{l_2}) \\
&\rightarrow_{\mathcal{C}}^* \text{err}^{l_2}
\end{aligned}$$

Fig. 6. Abusive function proxies in $\text{CPCF}_{\mathcal{C}}$

classic lax semantics, we find (leaving the domain check unevaluated):

$$\begin{aligned}
&\text{mon}(x:c_{11} \mapsto c_{12}, \text{mon}(x:c_{21} \mapsto c_{22}, f)) \ v \\
&\rightarrow_{\mathcal{C}} \text{mon}(c_{12}[v/x], \text{mon}(x:c_{21} \mapsto c_{22}, f) \ \text{mon}(c_{11}, v)) \\
&\rightarrow_{\mathcal{C}} \text{mon}(c_{12}[v/x], \text{mon}(c_{22}[\text{mon}(c_{11}, v)/x], f \ \text{mon}(c_{21}, \text{mon}(c_{11}, v))))
\end{aligned}$$

Even though we’re using the lax semantics, we substitute contracts into the codomain. For the space-efficient semantics to be sound, it must behave *exactly* like the classic semantics: no matter what joins happen, $\text{CPCF}_{\mathcal{E}}$ must replicate the contract substitutions done in $\text{CPCF}_{\mathcal{C}}$. We can construct an abusive contract in $\text{CPCF}_{\mathcal{C}}$ —even though it has lax semantics—by having the inner function proxy abuse the outer one (Fig. 6). Why was blame raised? Because c_2 ’s codomain contract *abused* c_1 ’s domain contract. Even though $\text{CPCF}_{\mathcal{C}}$ has a lax semantics, wrapping multiple function proxies leads to monitoring domains from one contract in the codomain of another—a situation ripe for abuse.

Space-efficiency means joining contracts, so how can we emulate this classic-semantics substitution behavior? We use the `wrap` function, forcing a substitution when two function contracts are joined. By keeping track of these substitutions at every join, any joins that happen in the future will be working on contracts which already have appropriate substitutions.

$\text{CPCF}_{\mathcal{E}}$ uses *labeled contracts* (Fig. 2); substitution for labeled predicate contracts is explicit and delayed, as for ordinary contracts:

$$\begin{aligned}
\text{pred}_{\sigma}^l(e)[v/x] &= \begin{cases} \text{pred}_{\sigma[x \mapsto v]}^l(e) & x \in \text{fv}(\sigma(e)) \\ \text{pred}_{\sigma}^l(e) & \text{otherwise} \end{cases} \\
\text{nil}[v/x] &= \text{nil} \\
(\text{pred}_{\sigma}^l(e); r)[v/x] &= \text{pred}_{\sigma}^l(e)[v/x]; r[v/x]
\end{aligned}$$

We do *not* do any joining when a substitution occurs (but see Sec. 6). In $\text{CPCF}_{\mathcal{E}}$, closing substitutions map each variable to (a) itself ($[x \mapsto x]$), (b) a monitor on itself ($[x \mapsto \text{mon}(c, x)]$), or (c) a value. We add an evaluation rule taking ordinary

contract monitors $\text{mon}^l(C, e)$ to labeled-contract monitors $\text{mon}(c, e)$ by means of the labeling function `label` (`EMONLABEL`).

Space-efficiency comes by restricting congruence to only apply when there are abutting monitors (cf. `EMONC` here in `CPCFE` to `EMON` in `CPCFC`). When two monitors collide, we *join* them (`EMONCJOIN`). Checking function contracts is as usual (`EMONCAPP` is the same as `EMONAPP`, only the latter works over labeled contracts); checking predicate stacks proceeds straightforwardly predicate-by-predicate (`EMONCNIL` and `EMONCPRED`).

3 Soundness for space efficiency

`CPCFC` and `CPCFE` are operationally equivalent, even though their cast semantics differ. We can make this connection formal by proving that every `CPCF` term either: (a) diverges in both `CPCFC` and `CPCFE` or (b) reduces to equivalent terms in both `CPCFC` and `CPCFE`.

One minor technicality: some of the forms in our language are necessary only for runtime or only appear in one of the two calculi. We characterize *source programs* as those which omit runtime terms.

Definition 2 (Source program). *A well typed source program does not use `TBLAME` or `TMONC` (and so `TCNIL`, `TCPRED`, and `TCFUN` cannot be used).*

Greenberg identified the key property for proving soundness of a space efficient semantics: to be sound, the space-efficient semantics must recover a notion of congruence for checking.

Lemma 3 (Monitor congruence (single step)). *If $\emptyset \vdash e_1 : T$ and $\emptyset \vdash c : T$ and $e_1 \rightarrow_E e_2$, then $\text{mon}(c, e_1) \rightarrow_E^* w$ iff $\text{mon}(c, e_2) \rightarrow_E^* w$.*

Proof. By cases on the step taken to find $e_1 \rightarrow_E e_2$. In the easy case, there's no joining of coercions and the same rule can apply in both derivations. In the more interesting case, two contract monitors join. In either case, it suffices to show that the terms are ultimately confluent, since determinism will do the rest.

It is particularly satisfying that the key property for showing soundness of space efficiency can be proved independently of the inefficient semantics. Implementors can work entirely in the context of the space-efficient semantics, knowing that congruence ensures soundness. We show the observational equivalence of `CPCFC` and `CPCFE` by logical relations (Fig. 7), which gives us contextual equivalence—the strongest equivalence we could ask for.

Lemma 4 (Similar contracts are logically related). *If $\Gamma \vdash C_1 \sim C_2 : T$ and $\Gamma \vdash v_1 \simeq v_2 : T$ then $\Gamma \vdash \text{mon}^l(C_1, v_1) \simeq \text{mon}^l(C_2, v_2) : T$.*

Proof. By induction on the type index of the invariant relation $\Gamma \vdash C_1 \sim C_2 : T$.

Lemma 5 (Unwinding). *If $\emptyset \vdash \mu(x:T). e : T$, then $\mu(x:T). e \rightarrow_m^* w$ iff there exists an n such that unrolling the fixpoint only n times converges to the same value, i.e., $e[\mu(x:T). \dots e[\mu(x:T). e/x] \dots /x] \rightarrow_m^* w$.*

Result rules

$$\boxed{e_1 \sim e_2 : T}$$

$$\begin{aligned} k \sim k : B &\iff \text{ty}(k) = B \\ v_{11} \sim v_{21} : T_1 \rightarrow T_2 &\iff \forall e_{12} \sim e_{22} : T_1. v_{11} \ e_{12} \simeq v_{21} \ e_{22} : T_2 \\ \text{err}^l \sim \text{err}^l : T & \end{aligned}$$

Term rules

$$\boxed{e_1 \simeq e_2 : T}$$

$$e_1 \simeq e_2 : T \iff (e_1 \text{ diverges} \wedge e_2 \text{ diverges}) \vee (e_1 \rightarrow_{\mathbb{C}}^* w_1 \wedge e_2 \rightarrow_{\mathbb{E}}^* w_2 \wedge w_1 \sim w_2 : T)$$

Contract rules (invariant relation)

$$\boxed{\Gamma \vdash C_1 \sim C_2 : T}$$

$$\begin{aligned} \Gamma \vdash \text{pred}_{\sigma_1}(e_1) \sim \text{pred}_{\sigma_2}(e_2) : B &\iff \Gamma \vdash \sigma_1 \ (e_1) \simeq \sigma_2(e_2) : B \rightarrow \text{Bool} \\ \Gamma \vdash x : C_{11} \mapsto C_{12} \sim x : C_{21} \mapsto C_{22} : T_1 \rightarrow T_2 &\iff \\ \Gamma \vdash C_{11} \sim C_{21} : T_1 \wedge \Gamma, x : T_1 \vdash C_{12} \sim C_{22} : T_2 & \end{aligned}$$

Closing substitutions and open terms

$$\boxed{\Gamma \models \delta}$$

$$\boxed{\Gamma \vdash e_1 \simeq e_2 : T}$$

$$\begin{aligned} \Gamma \models \delta &\iff \forall x \in \text{dom}(\Gamma). \delta_1(x) \sim \delta_2(x) : \Gamma(x) \\ \Gamma \vdash e_1 \simeq e_2 : T &\iff \forall \Gamma \models \delta. \delta_1(e_1) \simeq \delta_2(e_2) : T \end{aligned}$$

Fig. 7. Logical relation between classic and space-efficient CPCF**Theorem 6 (CPCF_C and CPCF_E terms are logically related).**

1. If $\Gamma \vdash e : T$ as a source program then $\Gamma \vdash e \simeq e : T$.
2. If $\Gamma \vdash C : T$ as a source program then $\Gamma \vdash C \sim C : T$.

Proof. By mutual induction on the typing relations.

4 Bounds for space efficiency

We have seen that CPCF_E behaves the same as CPCF_C (Theorem 6), but is CPCF_E actually space efficient? For programs that don't use dependency, yes! With dependency, the story is more complicated.

4.1 The simple case

Greenberg showed that for *simple* contracts—without dependency—we can put a bounds on space [10]. We can recover his result in our more general framework. Observe that a given source program e starts with a finite number of predicate contracts in it. As e runs, no new predicates appear (because dependent substitutions have no effect), but predicates may accumulate in stacks. In the worst case, a predicate stack could contain every predicate contract from the original program e exactly once... but no more than that, because joins remove redundancy! Function contracts are also bounded: e starts out with function contracts

Predicate extraction	$\text{preds}(e), \text{preds}(C), \text{preds}(c) : \mathcal{P}(e \times (\text{Var} \rightarrow e))$
$\text{preds}(x) = \emptyset$	$\text{preds}(\text{pred}_\sigma(e)) =$
$\text{preds}(k) = \emptyset$	$\{(e, \sigma)\} \cup \text{preds}(e) \cup \bigcup_{[x \mapsto v] \in \sigma} \text{preds}(v)$
$\text{preds}(\lambda x:T. e) = \text{preds}(e)$	$\text{preds}(x:C_1 \mapsto C_2) = \text{preds}(C_1) \cup \text{preds}(C_2)$
$\text{preds}(\text{mon}^l(C, e)) = \text{preds}(C) \cup \text{preds}(e)$	$\text{preds}(\text{nil}) = \emptyset$
$\text{preds}(\text{mon}(c, e)) = \text{preds}(c) \cup \text{preds}(e)$	$\text{preds}(\text{pred}_\sigma^l(e); r) = \{(e, \sigma)\} \cup \text{preds}(e) \cup$
$\text{preds}(e_1 \ e_2) = \text{preds}(e_1) \cup \text{preds}(e_2)$	$\bigcup_{[x \mapsto e'] \in \sigma} \text{preds}(e') \cup \text{preds}(r)$
$\text{preds}(e_1 \ \text{op} \ e_2) = \text{preds}(e_1) \cup \text{preds}(e_2)$	$\text{preds}(x:c_1 \mapsto c_2) = \text{preds}(c_1) \cup \text{preds}(c_2)$
$\text{preds}(\text{if } e_1 \ e_2 \ e_3) =$	
$\text{preds}(e_1) \cup \text{preds}(e_2) \cup \text{preds}(e_3)$	
$\text{preds}(\text{err}^l) = \emptyset$	

Contract size	$P_B : \mathbb{N}$	$S_B : \mathbb{N}$	$\text{size}(C) : \mathbb{N}$
	$P_B = \{e \in \text{preds}(e) \mid \Gamma \vdash \text{pred}_\sigma(e) : B\} \quad S_B = L \cdot P_B \cdot \log_2 P_B$		
	$\text{size}(\text{pred}_\sigma(e)) = S_B \text{ when } \emptyset \vdash \text{pred}_\sigma(e) : B \quad \text{size}(x:C_1 \mapsto C_2) = \text{size}(C_1) + \text{size}(C_2)$		

Fig. 8. Predicate extraction and contract size

of a certain height, and evaluation can only shrink that height. The leaves of function contracts are labeled with predicate stacks, so the largest contract we could ever see is of maximum height with maximal predicate stacks at every leaf. As the program runs, abutting monitors are joined, giving us a bound on the total number of monitors in a program (one per non-monitor AST node).

We can make these ideas formal by first defining what we mean by “all the predicates in a program”, and then showing that evaluation doesn’t introduce predicates (Lemma 9). We let $\text{preds}(e)$ be the set of predicates in a term, where a predicate is represented as a pair of term and a closing substitution.

We say program e *uses simple contracts* when all predicates in e are closed and every predicate stack has no redundancies. As such a program reduces, no new contracts appear (and contracts may disappear).

Lemma 7. $\text{preds}(e[e'/x]) \subseteq \text{preds}(e) \cup \text{preds}(e')$

Proof. By induction on e . If e is a predicate contract, it has no free variables (by assumption), so the substitution doesn’t hold on to anything.

Lemma 8. If $\emptyset \vdash c_1 : T$ and $\emptyset \vdash c_2 : T$ then $\text{preds}(\text{join}(c_1, c_2)) \subseteq \text{preds}(c_1) \cup \text{preds}(c_2)$.

Proof. By induction on c_1 , ignoring wrap’s substitution by Lemma 7.

Lemma 9 (Reduction is non-increasing in simple predicates). If $\emptyset \vdash e : T$ and $e \rightarrow_m e'$ then $\text{preds}(e') \subseteq \text{preds}(e)$.

Proof. By induction on the step taken.

To compute the concrete bounds, we define P_B as the number of distinct predicates at the base type B . We can represent a predicate stack at type B in S_B bits, where L is the number of bits needed to represent a blame label. A given well typed contract $\emptyset \vdash C : T$ can then be represented in $\text{size}(C)$ bits, where each predicate stacks are represented is S_B bits and function types are represented as trees of predicate stacks. Finally, since reduction is non-increasing (Lemma 9), we can bound the amount of space used by any contract by looking at the source program, e : we can represent all contracts in our program in at most $s = \max_{C \in e} \text{size}(C)$ space—constant for a fixed source program.

Readers familiar with Greenberg’s paper (and earlier work, like Herman et al. [13]) will notice that our bounds are more precise, tracking the number of holes in contracts per type ($\text{size}(C)$) rather than simply computing the largest conceivable type ($2^{\text{height}(T)}$).

4.2 The dependent case

In the dependent case, we can’t *generally* bound the number of contracts by the size of contracts used in the program. Consider the following term, where $n \in \mathbb{N}$:

```

let downTo =  $\mu(f:\text{Int} \rightarrow \text{Int}).$ 
  monl( $x:\text{pred}(\lambda x:\text{Int}. x \geq 0) \mapsto \text{pred}(\lambda y:\text{Int}. x \geq y),$ 
     $\lambda x:\text{Int}. \text{if } (x = 0) 0 (f (x - 1))$ ) in
downTo n

```

How many different contracts will appear in a run of this program? As `downTo` runs, we’ll see n different forms of the predicate $\text{pred}_{\sigma_i}^l(\lambda y:\text{Int}. x \geq y)$. We’ll have one $\sigma_n = [x \mapsto n]$ on the first call, $\sigma_{n-1} = [x \mapsto n - 1]$ on the second call, and so on. But n ’s magnitude doesn’t affect our measure of the size of source program’s contracts. The number of distinct contracts that appear will be effectively unbounded.

In the simple case, we get bounds automatically, using the smallest possible implication relation—syntactic equality. In the dependent case, it’s up to the programmer to identify implications that recover space efficiency. We can recover space efficiency for `downTo` by saying $\text{pred}_{\sigma_1}(\lambda y:\text{Int}. x \geq y) \supset \text{pred}_{\sigma_2}(\lambda y:\text{Int}. x \geq y)$ iff $\sigma_1(x) \leq \sigma_2(x)$. Then the codomain checks from recursive calls will be able to join:

$$\begin{aligned}
\text{downTo } n &\longrightarrow_{\mathbb{E}}^* \text{mon}^l(\text{pred}_{[x \mapsto n]}(\dots), \dots) \\
&\longrightarrow_{\mathbb{E}}^* \text{mon}^l(\text{pred}_{[x \mapsto n]}(\dots), \text{mon}^l(\text{pred}_{[x \mapsto n-1]}(\dots), \dots)) \\
&\longrightarrow_{\mathbb{E}}^* \text{mon}^l(\text{pred}_{[x \mapsto n-1]}(\dots), \dots)
\end{aligned}$$

Why are we able to recover space efficiency in this case? Because we can come up with an easily decidable implication rule for our specific predicates matching how our function checks narrower and narrower properties as it recurses.

Recall the mutually recursive `even/odd` example (Fig. 1). We can make this example space-efficient by adding the implication that:

$$\text{pred}_{\sigma_1}(\lambda b:\text{Bool}. b \text{ or } (x \bmod 2 = 0)) \supset \text{pred}_{\sigma_2}(\lambda b:\text{Bool}. b \text{ or } (x \bmod 2 = 0))$$

iff $\sigma_1(x) + 2 = \sigma_2(x)$. Suppose we put contracts on both **even** and **odd**:

```
let odd = monlodd(x:pred(λx:Int. x ≥ 0) ↦ pred(λb:Bool. b or (x mod 2 = 0)),
                  λx:Int. if (x = 0) false (even (x - 1)))
and even =
  monleven(x:pred(λx:Int. x ≥ 0) ↦ pred(λb:Bool. b or ((x + 1) mod 2 = 0)),
            λx:Int. if (x = 0) true (odd (x - 1)))
```

Now our trace of contracts won't be homogeneous; eliding domain contracts:

```
odd 5 →C* monlodd(pred[x↦5](...), even 4)
      →C* monlodd(pred[x↦5](...), monleven(pred[x↦4](...),
      monlodd(pred[x↦3](...), monleven(pred[x↦2](...),
      monlodd(pred[x↦1](...), even 0))))
```

To make these checks space efficient, we'd need several implications; we write odd_p for $\lambda b:\text{Bool}. b \text{ or } (x \bmod 2 = 0)$ and even_p for $\lambda b:\text{Bool}. b \text{ or } ((x + 1) \bmod 2 = 0)$. The following table gives conditions on the implication relation for the row predicate to imply the column predicate:

\supset	$\text{pred}_{\sigma_2}(\text{odd}_p)$	$\text{pred}_{\sigma_2}(\text{even}_p)$
$\text{pred}_{\sigma_1}(\text{odd}_p)$	$\sigma_1(x) + 2 = \sigma_2(x)$	$\sigma_1(x) + 1 = \sigma_2(x)$
$\text{pred}_{\sigma_1}(\text{even}_p)$	$\sigma_1(x) + 1 = \sigma_2(x)$	$\sigma_1(x) + 2 = \sigma_2(x)$

Having all four of these implications allows us to eliminate any pair of checks generated by the recursive calls in **odd** and **even**, reducing the codomain checking to constant space—here, just one check. We could define a different implication relation, where, say, $\text{pred}_{\sigma_1}(\text{odd}_p) \supset \text{pred}_{\sigma_2}(\text{odd}_p)$ iff $\sigma_1(x) \bmod 2 = \sigma_2(x) \bmod 2$. Such an implication would apply more generally than those in the table above.

As usual, there is a trade-off between time and space. It's possible to write contracts where the necessary implication relation for space efficiency amounts to checking both contracts. Consider the following tail-recursive factorial function:

```
let any = λz:Int. true
let fact = μ(f:Int→Int→Int).
  monl(x:pred(any) ↦ acc:pred(any) ↦ pred(λy:Int. x ≥ 0),
      λx:Int. λacc:Int. if (x = 0) acc (f (x - 1) (x * acc)))
```

This contract isn't *wrong*, just strange: if you call **fact** with a negative number, the program diverges and you indeed won't get a value back out (contracts enforce partial correctness). A call of **fact** 3 yields monitors that check, from outside to inside, that $3 \geq 0$ and $2 \geq 0$ and $1 \geq 0$ and $0 \geq 0$. When should we say that $\text{pred}_{\sigma_1}(\lambda y:\text{Int}. x \geq 0) \supset \text{pred}_{\sigma_1}(\lambda y:\text{Int}. x \geq 0)$? We could check that $\sigma_1(x) \geq \sigma_2(x)$... but the time cost is just like checking the original contract.

5 Where should the implication relation come from?

The simplest option is to punt: derive the implication relation as the reflexive transitive closure of a programmer's rules. A programmer might specify how several different predicates interrelate as follows:

```

1 y: Int{x1 >= y} implies y: Int{x2 >= y} when x1 <= x2
2 y: Int{x1 > y} implies y: Int{x2 >= y} when x1 <= x2 + 1
3 y: Int{x1 > y} implies y: Int{x2 > y} when x1 <= x2

```

A default collection of such implications might come with the language; library programmers should be able to write their own, as well. But it is probably unwise to allow programmers to write arbitrary implications: what if they’re wrong? A good implementation would only accept verified implications, using a theorem prover or an SMT solver to avoid bogus implications.

Rather than having programmers write their own implications, we could try to *automatically* derive the implications. Given a program, a fixed number of predicates occur, even if an unbounded number of predicate/closing substitution pairings might occur at runtime. Collect all possible predicates from the source program, and consider each pair of predicates over the same base type, $\text{pred}(e_1)$ and $\text{pred}(e_2)$ such that $\Gamma \vdash e_i : B \rightarrow \text{Bool}$. We can derive from the typing derivation the shapes of the respective closing substitutions, σ_1 and σ_2 . What are the conditions on σ_1 and σ_2 such that $\text{pred}_{\sigma_1}(e_1) \supset \text{pred}_{\sigma_2}(e_2)$? We are looking for a property $P(\sigma_1, \sigma_2)$ such that:

$$\forall k \in \mathcal{K}_B, P(\sigma_1, \sigma_2) \wedge \sigma_1(e_1) k \longrightarrow_E^* \text{true} \Rightarrow \sigma_2(e_2) k \longrightarrow_E^* \text{true}$$

Ideally, P is also efficiently decidable—at least more efficiently than deciding both predicates. The problem of finding P can be reduced to that of finding the weakest precondition for the safety of the following function:

```

1 fun x:B =>
2   let y0 = v10, ..., yn = v1n (* σ1's bindings *)
3       z0 = v20, ..., zn = v2m (* σ2's bindings *) in
4   if e1 x then (if e2 x then () else error) else ()

```

Since P would be the *weakest* precondition, we would know that we had found the most general condition for the implication relation. Whether or not the most general condition is the *best* condition depends on context. We should also consider a cost model for P ; programmers may want to occasionally trade space for time, not bothering to join predicates that would be expensive to test.

Finding implication conditions resembles liquid type inference [19,27,15]: programmers get a small control knob (which expressions can go in P) and then an SMT solver does the rest. The settings are different, though: liquid types are about verifying programs, while we’re executing checks at runtime.

5.1 Implementation

Implementation issues abound. How should the free variables in terms be represented? What kind of refactorings and optimizations can the compiler do, and how might they interfere with the set of contracts that appear in a program? When is the right moment in compilation to fix the implication relation? More generally, what’s the design space of closure representations and calling conventions for languages with contracts?

6 Extensions

Generalizing our space-efficient semantics to sums and products does not seem particularly hard: we'd need contracts with corresponding shapes, and the join operation would push through such shapes. Recursive types and datatypes are more interesting [21]. Findler et al.'s lazy contract checking keeps contracts from changing the asymptotic time complexity of the program [7]; we may be able to adapt their work to avoid changes in asymptotic space complexity, as well.

The predicates here range over base types, but we could also allow predicates over other types. If we allow predicates over higher types, how should the adequacy constraint on predicate implication (Definition 1) change?

Impredicative polymorphism in the style of System F would require even more technical changes. The introduction of type variables would make our reasoning about names and binders trickier. In order to support predicates over type variables, we'd need to allow predicates over higher types—and so our notion of adequacy of \supset would change. In order to support predicates over quantified types, we'd need to change adequacy again. Adequacy would end up looking like the logical relation used to show relational parametricity: when would we have $\forall\alpha.T_1 \supset \forall\alpha.T_2$? If we substitute T'_1 for α on the left and T'_2 for α on the right (and T'_1 and T'_2 are somehow related), then $T_1[T'_1/\alpha] \supset T_2[T'_2/\alpha]$. Not only would the technicalities be tricky, implementations would need to be careful to manage closure representations correctly (e.g., what happens if polymorphic code differs for boxed and unboxed types?).

We don't treat blame as an interesting algebraic structure—it's enough for our proofs to show that we always produce the same answer. Changing our calculus to have a more interesting notion of blame, like *indy* semantics [4] or involutive blame labels [29,28], would be a matter of pushing a shallow change in the semantics through the proofs.

Finally, it would make sense to have substitution on predicate stacks perform joins, saying $(\text{pred}_\sigma^l(e); r)[v/x] = \text{join}(\text{pred}_\sigma^l(e)[v/x]; \text{nil}, r[v/x])$, so that substituting a value into a predicate stack checks for newly revealed redundancies. We haven't proved that this change would be sound, which would require changes to both type and space-efficiency soundness.

7 Related work

For the technique of space efficiency itself, we refer the reader to Greenberg [10] for a full description of related work. We have striven to use Greenberg's notation exactly, but we made some changes in adapting to dependent contracts: the cons operator for predicate stacks is a semi-colon, to avoid ambiguity; there were formerly two things named *join*, but one has been folded into the other; and our predicates have closing substitutions to account for dependency. We place one more requirement on the implication relation than Greenberg did: monotonicity under substitution, which we call *substitutivity*. Substitutions weren't an issue in his non-dependent system, but we must require that if a join can happen without having a value for x , then the same join happens when we know x 's value.

CPCF was first introduced in several papers by Dimoulas et al. in 2011 [3,4], and has later been the subject of studies of blame for dependent function contracts [5] and static analysis [26]. Our exact behavioral equivalence means we could use results from Tobin-Hochstadt et al.’s static analysis in terms of $\text{CPCF}_{\mathcal{C}}$ to optimize space efficient programs in $\text{CPCF}_{\mathcal{E}}$. More interestingly, the predicate implication relation \supset seems to be doing some of the work that their static analysis does, so there may be a deeper relationship.

Thiemann introduces a manifest calculus where the compiler optimizes casts for time efficiency: a theorem prover uses the “delta” between types to synthesize more efficient checks [25]. His deltas and our predicate implication relation are similar. He uses a separate logical language for predicates and restricts dependency (codomains can only depend on base values, avoiding abusive contracts).

Sekiyama et al. [20] also use delayed substitutions in their polymorphic manifest contract calculus, but for different technical reasons. While delayed substitutions resemble explicit substitutions [1] or explicit bindings [12,2], we use delayed substitutions more selectively and to resolve issues with dependency.

The manifest type system in Greenberg’s work is somewhat disappointing compared to the type system given here. Greenberg works much harder than we do to prove a stronger type soundness theorem... but that theorem isn’t enough to help materially in proving the soundness of space efficiency. Developing the approach to dependency used here was much easier in a latent calculus, though several bugs along the way would have been caught early by a stronger type system. Type system complexity trade-offs of this sort are an old story.

7.1 Racket’s implementation

If contracts leak space, how is it that they are used so effectively throughout PLT Racket? Racket is designed to avoid using contracts in leaky ways. In Racket, contracts tend to go on module boundaries. Calls inside of a module then don’t trigger contract checks—including recursive calls, like in the `even/odd` example. Racket *will* monitor recursive calls across module boundaries, and these checks can indeed lead to space leaks. In our terms, Racket tends to implement contract checks on recursive functions as follows:

$$\begin{aligned} \text{downTo} &= \text{mon}^l(x:\text{pred}(\lambda x:\text{Int}. x \geq 0) \mapsto \text{pred}(\lambda y:\text{Int}. x \geq y), \\ &\quad \mu(f:\text{Int} \rightarrow \text{Int}). \lambda x:\text{Int}. \text{if } (x = 0) \ 0 \ (f \ (x - 1))) \end{aligned}$$

Note that calling `downTo n` will merely check that the final result is less than n —none of the intermediate values. Our version of `downTo` above puts the contract *inside* the recursive knot, forcing checks every time (Sec. 4.2).

Racket also offers a less thorough form of space efficiency. We can construct a program where Racket will avoid redundant checks, but wrapping the underlying function with the same contract twice leads to a space leak (Figure 9).³

Finally, contracts are first-class in Racket. Computing new contracts at runtime breaks our framing of space-efficiency: if we can’t predetermine which contracts arise at runtime, we can’t fix an implication relation in advance. We hope

³ Robby Findler, personal correspondence, 2016-05-19.

```

1 (define (count-em-integer? x)
2   (printf "checking ~s\n" x)
3   (integer? x))
4 (letrec
5   ([f (contract (-> any/c count-em-integer?)
6     (lambda (x) (if (zero? x) x (f (- x 1))))
7     'pos 'neg)])
8   (f 3))

```

Fig. 9. Space-efficiency in Racket

that CPCF_E is close enough to Racket’s internal model to provide insight into how to achieve space efficiency for at least some contracts in Racket.

8 Conclusion

We have translated Greenberg’s original result [10] from a manifest calculus [11] to a latent one [3,4]. In so doing, we have: offered a simpler explanation of the original result; isolated the parts of the type system required for space bounds; and, extended the original result, by covering more features (dependency and nontermination) and more precisely bounding non-dependent programs.

Acknowledgments. The existence of this paper is due to comments from Sam Tobin-Hochstadt and David Van Horn that I chose to interpret as encouragement. Robby Findler provided the Racket example and helped correct and clarify a draft; Sam Tobin-Hochstadt also offered corrections and suggestions. The reviews offered helpful comments, too.

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