

Contracts Made Manifest

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Abstract

Since Findler & Felleisen (2002) introduced *higher-order contracts*, many variants have been proposed. Broadly, these fall into two groups: some follow Findler and Felleisen in using *latent* contracts, purely dynamic checks that are transparent to the type system; others use *manifest* contracts, where *refinement types* record the most recent check that has been applied to each value. These two approaches are commonly assumed to be equivalent—different ways of implementing the same idea, one retaining a simple type system, and the other providing more static information. Our goal is to formalize and clarify this folklore understanding.

Our work extends that of Gronski & Flanagan (2007), who defined a latent calculus λ_C and a manifest calculus λ_H , gave a translation ϕ from λ_C to λ_H , and proved that, if a λ_C term reduces to a constant, then so does its ϕ -image. We enrich their account with a translation ψ from λ_H to λ_C and prove an analogous theorem.

We then generalize the whole framework to *dependent contracts*, whose predicates can mention free variables. This extension is both pragmatically crucial, supporting a much more interesting range of contracts, and theoretically challenging. We define dependent versions of λ_H and two dialects (“lax” and “picky”) of λ_C , establish type soundness—a substantial result in itself, for λ_H —and extend ϕ and ψ accordingly. Surprisingly, the intuition that the latent and manifest systems are equivalent now breaks down: the extended translations preserve behavior in one direction but, in the other, sometimes yield terms that blame more.

This is a longer version of a POPL 2010 paper (Greenberg et al., 2010), with proofs and extended discussion.

1 Introduction

The idea of contracts—arbitrary program predicates acting as dynamic pre- and post-conditions—was popularized by Eiffel (Meyer, 1992). More recently, Findler & Felleisen (2002) introduced a λ -calculus with *higher-order contracts*. This calculus includes terms like $\langle \{x:\text{Int} \mid \text{pos } x\} \rangle^{l,l'} 1$, in which a boolean predicate, *pos*, is applied to a run-time value, 1. This term evaluates to 1, since *pos* 1 returns true. On the other hand, the term $\langle \{x:\text{Int} \mid \text{pos } x\} \rangle^{l,l'} 0$ evaluates to *blame*, written $\uparrow l$, signaling that a contract with label *l* has been violated. The other label on the contract, *l'*, comes into play with *function contracts*, $c_1 \mapsto c_2$. For example, the term

$$\langle \{x:\text{Int} \mid \text{nonzero } x\} \mapsto \{x:\text{Int} \mid \text{pos } x\} \rangle^{l,l'} (\lambda x:\text{Int}. x - 1)$$

“wraps” the function $\lambda x:\text{Int}. x - 1$ in a pair of checks: whenever the wrapped function is called, the argument is checked to see whether it is nonzero; if not, the blame term $\uparrow l'$ is

produced, signaling that the *context* of the contracted term violated the expectations of the contract. If the argument check succeeds, then the function is run and its result is checked against the contract $\text{pos } x$, raising $\uparrow l$ if this fails (e.g., if the wrapped function is applied to 1).

Findler & Felleisen’s work sparked a resurgence of interest in contracts, and in the intervening years a bewildering variety of related systems have been studied. Broadly, these come in two different sorts. In systems with *latent* contracts, types and contracts are orthogonal features. Examples of this style include Findler and Felleisen’s original system, Hinze *et al.* (2006), Blume & McAllester (2006), Chitil & Huch (2007), Guha *et al.* (2007), and Tobin-Hochstadt & Felleisen (2008). By contrast, *manifest* contracts are integrated into the type system, which tracks, for each value, the most recently checked contract. *Hybrid types* (Flanagan, 2006) are a well-known example in this style; others include the work of Ou *et al.* (2004), Wadler & Findler (2009), and Knowles *et al.* (2006).

The key feature of manifest systems is that descriptions like $\{x:\text{Int} \mid \text{nonzero } x\}$ are incorporated into the type system as *refinement types*. Values of refinement type are introduced via *casts* like $\langle\{x:\text{Int} \mid \text{true}\} \Rightarrow \{x:\text{Int} \mid \text{nonzero } x\}\rangle^l n$, which has static type $\{x:\text{Int} \mid \text{nonzero } x\}$ and checks, dynamically, that n really is nonzero, raising $\uparrow l$ otherwise. Similarly, $\langle\{x:\text{Int} \mid \text{nonzero } x\} \Rightarrow \{x:\text{Int} \mid \text{pos } x\}\rangle^l n$ casts an integer that is statically known to be nonzero to one that is statically known to be positive.

The manifest analogue of function contracts is casts between function types. For example, consider:

$$f = \langle[\text{Int}] \rightarrow [\text{Int}] \Rightarrow \{x:\text{Int} \mid \text{pos } x\} \rightarrow \{x:\text{Int} \mid \text{pos } x\}\rangle^l (\lambda x:I. x - 1),$$

where $[\text{Int}] = \{x:\text{Int} \mid \text{true}\}$. The sequence of events when f is applied to some argument n (of type P) is similar to what we saw before:

$$f n \longrightarrow_h \langle[\text{Int}] \Rightarrow \{x:\text{Int} \mid \text{pos } x\}\rangle^l ((\lambda x:[\text{Int}]. x - 1) (\langle\{x:\text{Int} \mid \text{pos } x\} \Rightarrow [\text{Int}]\rangle^l n))$$

First, n is cast from $\{x:\text{Int} \mid \text{pos } x\}$ to $[\text{Int}]$ (it happens that in this case the cast cannot fail, since the target predicate is just true, but if it did, it would raise $\uparrow l$); then the function body is evaluated; and finally its result is cast from $[\text{Int}]$ to $\{x:\text{Int} \mid \text{pos } x\}$, raising $\uparrow l$ if this fails. The domain cast is contravariant and the codomain cast is covariant.

One point to note here is that casts in the manifest system have just one label, while contract checks in the latent system have two. This difference is not fundamental, but rather a question of the pragmatics of assigning responsibility: both latent and manifest systems can be given more or less rich algebras of blame. Informally, a function contract check $\langle c_1 \mapsto c_2 \rangle^{l,l'} f$ divides responsibility for f ’s behavior between its body and its environment: the programmer is saying “If f is ever applied to an argument that doesn’t pass c_1 , I refuse responsibility ($\uparrow l'$), whereas if f ’s result for good arguments doesn’t satisfy c_2 , I accept responsibility ($\uparrow l$).” In a manifest system, the programmer who writes $\langle R_1 \rightarrow R_2 \Rightarrow S_1 \rightarrow S_2 \rangle^l f$ is saying “Although all I know statically about f is that its results satisfy R_2 when it is applied to arguments satisfying R_1 , I assert that it’s okay to use it on arguments satisfying S_1 [because I believe that S_1 implies R_1] and that its results will always satisfy S_2 [because R_2 implies S_2].” In the latter case, the programmer is taking responsibility for *both* assertions (so $\uparrow l$ makes sense in both cases), while the additional responsibility for checking that arguments satisfy S_1 will be discharged elsewhere (by another cast, with a different label).

While contract checks in latent systems seem intuitively to be much the same thing as casts in manifest systems, the formal correspondence is not immediate. How do the contravariant function casts of λ_H relate to the invariant checks of λ_C ? How does λ_H model λ_C 's pair of polarized blame labels? This has led to some confusion in the community about the nature of contracts. Indeed, as we will see, matters become yet murkier in richer languages with features such as dependency.

Gronski & Flanagan (2007) initiated a formal investigation of the connection between the latent and manifest worlds. They defined a core calculus, λ_C , capturing the essence of latent contracts in a simply typed lambda-calculus, and an analogous manifest calculus λ_H . To compare these systems, they introduced a type-preserving translation ϕ from λ_C to λ_H . What makes ϕ interesting is that it relates the languages feature for feature: contracts over base types are mapped to casts at base type, and function contracts are mapped to function casts. The main result is that ϕ preserves behavior, in the sense that if a term t in λ_C evaluates to a constant k , then so does its translation $\phi(t)$.

Our work extends theirs in two directions. First, we strengthen their main result by introducing a new feature-for-feature translation ψ from λ_H to λ_C and proving a similar correspondence theorem for ψ . (We also give a new, more detailed, proof of the correspondence theorem for ϕ .) This shows that the manifest and latent approaches are effectively equivalent in the nondependent case.

Second, and more significantly, we extend the whole story to allow dependent function contracts in λ_C and dependent arrow types in λ_H . Dependency is extremely handy in contracts, as it allows for precise specifications of how the results of functions depend on their arguments. For example, here is a contract that we might use with an implementation of vector concatenation:

$$z_1:\text{Vec} \mapsto z_2:\text{Vec} \mapsto \{z_3:\text{Vec} \mid \text{vlen } z_3 = \text{vlen } z_1 + \text{vlen } z_2\}$$

Adding dependent contracts to λ_C is easy: the dependency is all in the contracts and the types stay simple. We have just one significant design choice: should domain contracts be rechecked when the bound variable appears the codomain contract? This leads to two dialects of λ_C , one which does recheck (*picky* λ_C) and one which does not (*lax* λ_C). The choice is not clear—dependent contract systems have typically used the lax rule, while the picky one is arguably more correct—so we consider both. The question of which blame labels belong on this extra check is discussed at length in Dimoulas *et al.* (2011). In λ_H , on the other hand, dependency significantly complicates the metatheory, requiring the addition of a denotational semantics for types and kinds to break a potential circularity in the definitions, plus an intricate sequence of technical lemmas involving parallel reduction to establish type soundness. (Although Gronski & Flanagan worked only with nondependent λ_C and λ_H , Knowles & Flanagan (2010) showed soundness for a variant of dependent λ_H in which the order of evaluation is nondeterministic and failed casts get stuck instead of raising blame. See Section 7.)

Surprisingly, the tight correspondence between λ_C and λ_H breaks down in the dependent case: the natural generalization of the translations does not preserve blame exactly. Indeed, we can place λ_H between the two variants of λ_C on an “axis of blame” (Figure 1), where behavior is preserved exactly when moving left on the axis (from picky λ_C to λ_H to lax

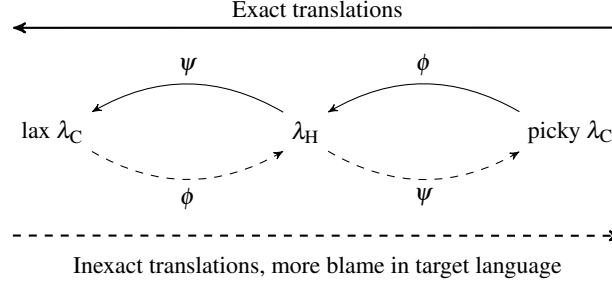


Fig. 1. The axis of blame

λ_C), but translated terms can blame more than their pre-images when moving right.¹ The discrepancy arises in the case of “abusive” contracts, such as

$$f: (\{x:\text{Int} \mid \text{nonzero } x\} \mapsto \{y:\text{Int} \mid \text{true}\}) \mapsto \{z:\text{Int} \mid f\ 0 = 0\}$$

This rather strange contract has the form $f:c_1 \mapsto c_2$, where c_2 uses f in a way that violates c_1 ! In particular, if we apply it (in $\text{lax } \lambda_C$) to $\lambda f:\text{Int} \rightarrow \text{Int}. 0$ and then apply the result to $\lambda x:\text{Int}. x$ and 5, the final result will be 5, since $\lambda x:\text{Int}. x$ does satisfy the contract $\{x:\text{Int} \mid \text{nonzero } x\} \mapsto \{y:\text{Int} \mid \text{true}\}$ and 5 satisfies the contract $\{z:\text{Int} \mid (\lambda x:\text{Int}. x)\ 0 = 0\}$. However, running the translation of f in λ_H yields extra check, wrapping the occurrence of f in the codomain contract with a cast from $\{x:\text{Int} \mid \text{nonzero } x\} \mapsto \{y:\text{Int} \mid \text{true}\}$ to $\{x:\text{Int} \mid \text{true}\} \mapsto \{y:\text{Int} \mid \text{true}\}$, which fails when the wrapped function is applied to 0. We discuss this phenomenon in greater detail in Section 4.

We should note at the outset that, like Gronski & Flanagan (2007), we’re interested in translations that relate λ_C and λ_H *feature for feature*. The dependent version of the wrap operator from Findler & Felleisen (2002) works fine as a translation, giving us an exact treatment of blame.

$$\begin{aligned} \phi(\langle \{x:B \mid t\} \rangle^{l,l'}) &= \langle [B] \Rightarrow \{x:B \mid \phi(t)\} \rangle^l \\ \phi(\langle x:c_1 \mapsto c_2 \rangle^{l,l'}) &= \lambda f: [x:c_1 \mapsto c_2]. \\ &\quad \lambda x: [c_1]. \\ &\quad \phi(\langle c_2 \rangle^{l,l'}) (f (\phi(\langle c_1 \rangle^{l',l}) x)) \end{aligned}$$

We can define a similar mapping function that implements λ_H ’s semantics as base type contracts in lax or $\text{picky } \lambda_C$. This is unsurprising: λ_C and λ_H are lambda calculi that feature, among other things, a way to conditionally raise exceptions. That these languages are interencodable is completely unsurprising. But translations like these don’t relate function contracts to function casts at all, so they don’t do much to tell us about how λ_C and λ_H semantics relate.

In summary, our main contributions are (a) the translation ψ and a symmetric version of Gronski & Flanagan’s behavioral correspondence theorem, (b) the basic metatheory of

¹ There might, in principle, be some other way of defining ϕ and ψ that (a) preserves types, (b) maps “feature for feature”, and (c) induces an exact behavioral equivalence. After considering a number of alternatives, we conjecture that no such ϕ and ψ exist.

$$\begin{array}{ll}
B ::= \text{Bool} \mid \dots & \text{base types} \\
k ::= \text{true} \mid \text{false} \mid \dots & \text{first-order constants}
\end{array}$$
Fig. 2. Base types and constants for λ_C and λ_H

(CBV, blame-sensitive) dependent λ_H , (c) dependent versions of ϕ and ψ and their properties with regard to λ_H and both dialects of λ_C , and (d) a weaker behavioral correspondence in the dependent case. This paper extends the discussion of Greenberg *et al.* (2010), giving the more interesting proofs.

2 The nondependent languages

We begin in this section by defining the nondependent versions of λ_C and λ_H and continue in Section 3 with the translations between them. The dependent languages, dependent translations, and their properties are developed in Sections 4, 5, and 6. Throughout the paper, rules prefixed with an E or a F are operational rules for λ_C and λ_H , respectively. An initial T is used for λ_C typing rules; typing rules beginning with an S belong to λ_H .

All of our languages will share a set of base types and first-order constants, given in Figure 2. Let the set \mathcal{K}_B contain constants of base type B . We assume that Bool is among the base types, with $\mathcal{K}_{\text{Bool}} = \{\text{true}, \text{false}\}$.

The language λ_C

The language λ_C is the simply typed lambda calculus straightforwardly augmented with contracts. The most interesting feature is the *contract* term $\langle c \rangle^{l,l'}$, which, when applied to a term t , dynamically ensures that t and its surrounding context satisfy c .² If t doesn't satisfy c , then the *positive* label l will be blamed and the whole term will reduce to $\uparrow l$; on the other hand, if the context doesn't treat $\langle c \rangle^{l,l'} t$ as c demands, then the negative label l' will be blamed and the term will reduce to $\uparrow l'$. Contracts come in two forms: base contracts $\{x:B \mid t\}$ over a base type B and higher-order contracts $c_1 \mapsto c_2$, which check the arguments and results of functions.

The syntax and semantics of λ_C appears in Figure 3, with some common definitions (shared with λ_H) in Figure 2. Besides the contract term $\langle c \rangle^{l,l'}$, λ_C includes first-order constants k , blame, and *active checks* $\langle \{x:B \mid t_1\}, t_2, k \rangle^l$. Active checks do not appear in source programs; they are a technical artifact of the small-step operational semantics, as we explain below. Also, note that we only allow contracts over base types B : we have function contracts, like $\{x:\text{Int} \mid \text{pos } x\} \mapsto \{x:\text{Int} \mid \text{nonzero } x\}$, but not contracts over functions, like $\{f:\text{Bool} \rightarrow \text{Bool} \mid f \text{ true} = f \text{ false}\}$.

Values v include constants, abstractions, contracts, and function contracts applied to values (more on these later); a *result* r is either a value or $\uparrow l$ for some l . We interpret constants using two constructions: the type-assignment function ty_c , which maps constants

² Our presentation differs slightly from that of Gronski & Flanagan (2007), since we use first-class contracts rather than forcing all contracts to be applied; details of how ϕ changes can be found in Sections 5.1 and 6.1.

Syntax for λ_C

$T ::= B \mid T_1 \rightarrow T_2$	types
$c ::= \{x:B \mid t\} \mid c_1 \mapsto c_2$	contracts
$t ::= x \mid k \mid \lambda x:T_1. t_2 \mid t_1 t_2 \mid \uparrow l \mid \langle c \rangle^{l,l'} \mid \langle \{x:B \mid t_1\}, t_2, k \rangle^l$	terms
$v ::= k \mid \lambda x:T_1. t_2 \mid \langle c \rangle^{l,l'} \mid \langle c_1 \mapsto c_2 \rangle^{l,l'} v$	values
$r ::= v \mid \uparrow l$	results
$E ::= [] t \mid v [] \mid \langle \{x:B \mid t\}, [], k \rangle^l$	evaluation contexts

Operational semantics for λ_C

$(\lambda x:T_1. t_2) v$	\longrightarrow_c	$t_2 \{x := v\}$	E_BETA
$k v$	\longrightarrow_c	$\llbracket k \rrbracket(v)$	E_CONST
$\langle \{x:B \mid t\} \rangle^{l,l'} k$	\longrightarrow_c	$\langle \{x:B \mid t\}, t \{x := k\}, k \rangle^l$	E_CCHECK
$\langle \{x:B \mid t\}, \text{true}, k \rangle^l$	\longrightarrow_c	k	E_OK
$\langle \{x:B \mid t\}, \text{false}, k \rangle^l$	\longrightarrow_c	$\uparrow l$	E_FAIL
$\langle \langle c_1 \mapsto c_2 \rangle^{l,l'} v \rangle v'$	\longrightarrow_c	$\langle c_2 \rangle^{l,l'} (v (\langle c_1 \rangle^{l,l'} v'))$	E_CDECOMP
$E[\uparrow l]$	\longrightarrow_c	$\uparrow l$	E_BLAZE
$E[t_1]$	\longrightarrow_c	$E[t_2]$ when $t_1 \longrightarrow_c t_2$	E_COMPAT

Typing rules for λ_C

$\boxed{\Gamma \vdash t : T}$		
$\frac{x:T \in \Gamma}{\Gamma \vdash x : T}$	T_VAR	$\frac{}{\Gamma \vdash k : \text{ty}_c(k)}$ T_CONST
$\frac{\Gamma, x:T_1 \vdash t_2 : T_2}{\Gamma \vdash \lambda x:T_1. t_2 : T_1 \rightarrow T_2}$	T_LAM	$\frac{\Gamma \vdash t_1 : T_1 \rightarrow T_2 \quad \Gamma \vdash t_2 : T_1}{\Gamma \vdash t_1 t_2 : T_2}$ T_APP
$\frac{\vdash c : T}{\Gamma \vdash \langle c \rangle^{l,l'} : T \rightarrow T}$	T_CONTRACT	$\frac{}{\Gamma \vdash \uparrow l : T}$ T_BLAZE
$\frac{\emptyset \vdash k : B \quad \emptyset \vdash t_2 : \text{Bool} \quad \vdash \{x:B \mid t_1\} : B \quad t_2 \longrightarrow_c^* \text{true} \text{ implies } t_1 \{x := k\} \longrightarrow_c^* \text{true}}{\emptyset \vdash \langle \{x:B \mid t_1\}, t_2, k \rangle^l : B}$		
$\boxed{\vdash c : T}$		
$\frac{x:B \vdash t : \text{Bool}}{\vdash \{x:B \mid t\} : B}$	T_BASEC	$\frac{\vdash c_1 : T_1 \quad \vdash c_2 : T_2}{\vdash c_1 \mapsto c_2 : T_1 \rightarrow T_2}$ T_FUNC

Fig. 3. Syntax and semantics for λ_C

to first-order types of the form $B_1 \rightarrow B_2 \rightarrow \dots \rightarrow B_n$ (and which is assumed to agree with \mathcal{K}_B); and the denotation function $\llbracket - \rrbracket$ which maps constants to functions from constants to constants (or blame, to allow for partiality). Denotations must agree with ty_c , i.e., if $\text{ty}_c(k) = B_1 \rightarrow B_2$, then $\llbracket k \rrbracket(k_1) \in \mathcal{K}_{B_2}$ if $k_1 \in \mathcal{K}_{B_1}$.

The operational semantics is given in Figure 3. It includes six rules for basic (small-step, call-by-value) reductions, plus two rules that involve evaluation contexts E (Figure 3). The evaluation contexts implement left-to-right evaluation for function application. If $\uparrow l$

appears in the active position of an evaluation context, it is propagated to the top level, like an uncatchable exception. As usual, values (and results) do not step.

The first two basic rules are standard, implementing primitive reductions and β -reductions for abstractions. In these rules, arguments must be values v . Since constants are first-order, we know that when E_CONST reduces a well-typed application, the argument is not just a value, but a constant.

The rules E_CCHECK, E_OK, E_FAIL and E_CDECOMP, describe the semantics of contracts. In E_CCHECK, base-type contracts applied to constants step to an active check. Active checks include the original contract, the current state of the check, the constant being checked, and a label to blame if necessary. If the check evaluates to true, then E_OK returns the initial constant. If false, the check has failed and a contract has been violated, so E_FAIL steps the term to $\uparrow l$. Higher-order contracts on a value v wait to be applied to an additional argument. This is why function contracts applied to values are values. There is no substantial difference between this approach and expanding function contracts into new lambdas. When that argument has also been reduced to a value v' , E_CDECOMP decomposes the function cast: the argument value is checked with the argument part of the contract (switching positive and negative blame, since the context is responsible for the argument), and the result of the application is checked with the result contract.

The typing rules for λ_C (Figure 3) are mostly standard. We give types to constants using the type-assignment function ty_c . Blame expressions have all types. Contracts are checked for well-formedness using the judgment $\vdash c : T$, comprising the rules T_BASEC, which requires that the checking term in a base contract return a boolean value when supplied with a term of the right type, and T_FUNC. Note that the predicate t in a contract $\{x:B \mid t\}$ can contain at most x free, since we are considering only nondependent contracts for now. Contract application, like function application, is checked using T_APP.

The T_CHECKING rule only applies in the empty context (active checks are only created at the top level during evaluation). The rule ensures that the contract $\{x:B \mid t_1\}$ has the right base type for the constant k , that the check expression t_2 has a boolean type, and that the check is actually checking the right contract. The latter condition is formalized by the implication: $t_2 \rightarrow_c^* \text{true}$ implies $t_1\{x := k\} \rightarrow_c^* \text{true}$ asserts that if t_2 evaluates to true, then the original check $t_1\{x := k\}$ must also evaluate to true. This requirement is needed for two reasons: first, nonsensical terms like $\langle \{x:\text{Int} \mid \text{pos } x\}, \text{true}, 0 \rangle^l$ should not be well typed; and second, we use this property in showing that the translations are type preserving (see Section 5). This rule obviously makes typechecking for the full “internal language” with checks undecidable, but excluding checks recovers decidability. We could give a more precise condition—for example, that $t_1\{x := k\} \rightarrow_c^* t_2$ —but there is no need.

The language enjoys standard preservation and progress theorems. Together, these ensure that evaluating a well-typed term to a normal form always yields a result r , which is either blame or a value.

The language λ_H

Our second core calculus, nondependent λ_H , extends the simply typed lambda-calculus with *refinement types* and *cast expressions*. The definitions appear in Figure 4. Unlike λ_C , which separates contracts from types, λ_H combines them into refined base types $\{x:B \mid s_1\}$

Syntax for λ_H

$S ::= \{x:B \mid s_1\} \mid S_1 \rightarrow S_2$	types/contracts
$s ::= x \mid k \mid \lambda x:S_1. s_2 \mid s_1 s_2 \mid \uparrow l \mid \langle S_1 \Rightarrow S_2 \rangle^l \mid \langle \{x:B \mid s_1\}, s_2, k \rangle^l$	terms
$w ::= k \mid \lambda x:S_1. s_2 \mid \langle S_1 \Rightarrow S_2 \rangle^l \mid \langle S_{11} \rightarrow S_{12} \Rightarrow S_{21} \rightarrow S_{22} \rangle^l w$	values
$q ::= w \mid \uparrow l$	results
$F ::= []s \mid w[] \mid \langle \{x:B \mid s\}, [], k \rangle^l$	evaluation contexts

Operational semantics for λ_H

$(\lambda x:S_1. s_2) w_2 \longrightarrow_h s_2 \{x := w_2\}$	F_BETA
$k w \longrightarrow_h \llbracket k \rrbracket(w)$	F_CONST
$\langle \{x:B \mid s_1\} \Rightarrow \{x:B \mid s_2\} \rangle^l k \longrightarrow_h \langle \{x:B \mid s_2\}, s_2 \{x := k\}, k \rangle^l$	F_CCHECK
$\langle \{x:B \mid s\}, \text{true}, k \rangle^l \longrightarrow_h k$	F_OK
$\langle \{x:B \mid s\}, \text{false}, k \rangle^l \longrightarrow_h \uparrow l$	F_FAIL
$\langle \langle S_{11} \rightarrow S_{12} \Rightarrow S_{21} \rightarrow S_{22} \rangle^l w \rangle w' \longrightarrow_h \langle S_{12} \Rightarrow S_{22} \rangle^l (w (\langle S_{21} \Rightarrow S_{11} \rangle^l w'))$	F_CDECOMP
$F[\uparrow l] \longrightarrow_h \uparrow l$	F_BLAME
$F[s_1] \longrightarrow_h F[s_2] \text{ when } s_1 \longrightarrow_h s_2$	F_COMPAT

Typing rules for λ_H

$\boxed{\Delta \vdash s : S}$	
$\frac{x:S \in \Delta}{\Delta \vdash x : S}$	S_VAR
$\frac{}{\Delta \vdash k : \text{ty}_h(k)}$	S_CONST
$\frac{\vdash S_1 \quad \Delta, x:S_1 \vdash s_2 : S_2}{\Delta \vdash \lambda x:S_1. s_2 : S_1 \rightarrow S_2}$	S_LAM
$\frac{\Delta \vdash s_1 : S_1 \rightarrow S_2 \quad \Delta \vdash s_2 : S_1}{\Delta \vdash s_1 s_2 : S_2}$	S_APP
$\frac{\vdash S_1 \quad \vdash S_2 \quad [S_1] = [S_2]}{\Delta \vdash \langle S_1 \Rightarrow S_2 \rangle^l : S_1 \rightarrow S_2}$	S_CAST
$\frac{\vdash S}{\Delta \vdash \uparrow l : S}$	S_BLAME
$\frac{\Delta \vdash s : S_1 \quad \vdash S_2 \quad \vdash S_1 <: S_2}{\Delta \vdash s : S_2}$	S_SUB
$\frac{\emptyset \vdash k : \{x:B \mid \text{true}\} \quad \emptyset \vdash s_2 : \{x:\text{Bool} \mid \text{true}\} \quad \vdash \{x:B \mid s_1\} \quad s_2 \longrightarrow_h^* \text{true} \text{ implies } s_1 \{x := k\} \longrightarrow_h^* \text{true}}{\emptyset \vdash \langle \{x:B \mid s_1\}, s_2, k \rangle^l : \{x:B \mid s_1\}}$	S_CHECKING
$\boxed{\vdash S_1 <: S_2}$	
$\frac{\forall k \in \mathcal{K}_B. (s_1 \{x := k\} \longrightarrow_h^* \text{true} \text{ implies } s_2 \{x := k\} \longrightarrow_h^* \text{true})}{\vdash \{x:B \mid s_1\} <: \{x:B \mid s_2\}}$	SSUB_REFINE
$\frac{\vdash S_{21} <: S_{11} \quad \vdash S_{12} <: S_{22}}{\vdash S_{11} \rightarrow S_{12} <: S_{21} \rightarrow S_{22}}$	SSUB_FUN
$\boxed{\vdash S}$	
$\frac{}{\vdash \{x:B \mid \text{true}\}}$	SWF_RAW
$\frac{x:\{x:B \mid \text{true}\} \vdash s : \{x:\text{Bool} \mid \text{true}\}}{\vdash \{x:B \mid s\}}$	SWF_REFINE
$\frac{\vdash S_1 \quad \vdash S_2}{\vdash S_1 \rightarrow S_2}$	SWF_FUN

Fig. 4. Syntax and semantics for λ_H

and function types $S_1 \rightarrow S_2$. As for λ_C , we do not allow refinement types over functions, nor do we allow refinements of refinements. (Belo *et al.* (2011) add these features to a dependent λ_H .) Unrefined base types B are *not* valid types; they must be wrapped in a trivial refinement, as the *raw* type $\{x:B \mid \text{true}\}$. The terms of the language are mostly standard, including variables, the same first-order constants as λ_C , blame, abstractions, and applications. The cast expression $\langle S_1 \Rightarrow S_2 \rangle^l$ dynamically checks that a term of type S_1 can be given type S_2 . Like λ_C , active checks are used to give a small-step semantics to cast expressions.

The values of λ_H include constants, abstractions, casts, and function casts applied to values. Results are either values or blame. We give meaning to constants as we did in λ_C , reusing $\llbracket - \rrbracket$. Type assignment is via ty_h , which we assume produces well-formed types (defined in Figure 4). To keep the languages in sync, we require that ty_h and ty_c agree on “type skeletons”: if $\text{ty}_c(k) = B_1 \rightarrow B_2$, then $\text{ty}_h(k) = \{x:B_1 \mid s_1\} \rightarrow \{x:B_2 \mid s_2\}$.

The small-step, call-by-value semantics in Figure 4 comprises six basic rules and two rules involving evaluation contexts F . Each rule corresponds closely to its λ_C counterpart.

Notice how the decomposition rules compare. In λ_C , the term $(\langle c_1 \mapsto c_2 \rangle^{l,l'} v) v'$ decomposes into two contract checks: c_1 checks the argument v' and c_2 checks the result of the application. In λ_H the term $(\langle S_{11} \rightarrow S_{12} \Rightarrow S_{21} \rightarrow S_{22} \rangle^l w) w'$ decomposes into two casts, but the behavior is a bit more subtle. The contravariant check $\langle S_{21} \Rightarrow S_{11} \rangle^l w'$ makes w' a suitable input for w , while $\langle S_{12} \Rightarrow S_{22} \rangle^l$ checks the result from w applied to (the cast) w' . Suppose $S_{21} = \{x:\text{Int} \mid \text{pos } x\}$ and $S_{11} = \{x:B \mid \text{nonzero } x\}$. Then the check on the argument ensures that $\text{nonzero } x \rightarrow_h^* \text{true}$ —not, as one might expect, that $\text{pos } w' \rightarrow_h^* \text{true}$. While it is easy to read off from a λ_C contract exactly which checks will occur at runtime, a λ_H cast must be dissected carefully to see exactly which checks will take place. On the other hand, which label will be blamed is clearer with casts—there’s only one!

The typing rules for λ_H (Figure 4) are also similar to those of λ_C . Just as the λ_C rule T_CONTRACT checks to make sure that the contract has the right form, the λ_H rule S_CAST ensures that the two types in a cast are well-formed and have the same simple-type skeleton, defined as $\lfloor - \rfloor : S \rightarrow T$ (pronounced “erase S ”):

$$\begin{aligned} \lfloor \{x:B \mid s\} \rfloor &= B \\ \lfloor S_1 \rightarrow S_2 \rfloor &= \lfloor S_1 \rfloor \rightarrow \lfloor S_2 \rfloor \end{aligned}$$

This prevents “stupid” casts, like $\langle \lceil \text{Int} \rceil \Rightarrow \lceil \text{Bool} \rceil \rangle^l$. We define a similar operator, $\lceil - \rceil : S \rightarrow S$ (pronounced “raw” S), which trivializes all refinements:

$$\begin{aligned} \lceil \{x:B \mid s\} \rceil &= \{x:B \mid \text{true}\} \\ \lceil S_1 \rightarrow S_2 \rceil &= \lceil S_1 \rceil \rightarrow \lceil S_2 \rceil \end{aligned}$$

These operations apply to λ_C contracts and types in the natural way. Type well-formedness in λ_H is similar to contract well-formedness in λ_C , though the SWF_RAW case needs to be added to get things off the ground.

The active check rule S_CHECKING plays a role analogous to the T_CHECKING rule in λ_C , again using an implication to guarantee that we only have sensible terms in the predicate position.

An important difference is that λ_H has subtyping. The S_SUB rule allows an expression to be promoted to any well-formed supertype. Refinement types are supertypes if, for all constants of the base type, their condition evaluates to true whenever the subtype’s condition evaluates to true. For function types, we use the standard contravariant subtyping rule. We do not consider source programs with subtyping, since this makes type checking undecidable³; subtyping is just a technical device for ensuring type preservation. Consider the following reduction:

$$\langle \{x:\text{Int} \mid \text{true}\} \Rightarrow \{x:\text{Int} \mid \text{pos } x\} \rangle^l 1 \longrightarrow_h^* 1$$

The source term is well-typed at $\{x:\text{Int} \mid \text{pos } x\}$. Since it evaluates to 1, we would like to have $\Delta \vdash 1 : \{x:\text{Int} \mid \text{pos } x\}$. To have type preservation in general, though, $\text{ty}_h(1)$ must be a subtype of $\{x:\text{Int} \mid s\}$ *whenever* $s\{x := 1\} \longrightarrow_h^* \text{true}$. That is, constants of base type must have “most-specific” types. One way to satisfy this requirement is to set $\text{ty}_h(k) = \{x:B \mid x = k\}$ for $k \in \mathcal{K}_B$; then if $s\{x := k\} \longrightarrow_h^* \text{true}$, we have $\vdash \text{ty}_h(k) <: \{x:B \mid s\}$.

Standard progress and preservation theorems hold for λ_H . We can also obtain a semantic type soundness theorem as a restriction of the one for dependent λ_H (Theorem 4.12).

3 The nondependent translations

The latent and manifest calculi differ in a few respects. Obviously, λ_C uses contract application and λ_H uses casts. Second, λ_C contracts have two labels—positive and negative—where λ_H contracts have a single label. Finally, λ_H has a much richer type system than λ_C . Our ψ from λ_H to λ_C and Gronski and Flanagan’s ϕ from λ_C to λ_H must account for these differences while carefully mapping “feature for feature”.

The interesting parts of the translations deal with contracts and casts. Everything else is translated homomorphically, though the type annotation on lambdas must be chosen carefully. The full definitions of these translations are in Section 5; the nondependent definitions are a straightforward restriction.

For ψ , translating from λ_H ’s rich types to λ_C ’s simple types is easy: we just erase the types to their simple skeletons. The interesting case is translating the cast $\langle S_1 \Rightarrow S_2 \rangle^l$ to a contract by translating the pair of types together, $\langle \psi(S_1, S_2) \rangle^{l,l}$. We define ψ as two mutually recursive functions: $\psi(s)$ translates λ_H terms to λ_C terms; $\psi(S_1, S_2)$ translates a pair of λ_H types—effectively, a cast—to a λ_C contract. The latter function is defined as follows:

$$\begin{aligned} \psi(\{x:B \mid s_1\}, \{x:B \mid s_2\}) &= \{x:B \mid \psi(s_2)\} \\ \psi(S_{11} \rightarrow S_{12}, S_{21} \rightarrow S_{22}) &= \psi(S_{21}, S_{11}) \mapsto \psi(S_{12}, S_{22}) \end{aligned}$$

We use the single label on the cast in both the positive and negative positions of the resulting contract, i.e.:

$$\psi(\langle S_1 \Rightarrow S_2 \rangle^l) = \langle \psi(S_1, S_2) \rangle^{l,l}.$$

³ Flanagan (2006) and Knowles & Flanagan (2010) discuss trade-offs between static and dynamic checking that allow for decidable type systems and subtyping.

When we translate a pair of refinement types, we produce a contract that will check the predicate of the target type (like `F_CCHECK`); when translating a pair of function types, we translate the domain contravariantly (like `F_CDECOMP`). For example,

$$\langle \{x:\text{Int} \mid \text{nonzero } x\} \rightarrow [\text{Int}] \Rightarrow [\text{Int}] \rightarrow \{y:\text{Int} \mid \text{pos } y\} \rangle^l$$

translates to $\langle \{x:\text{Int} \mid \text{nonzero } x\} \mapsto \{y:\text{Int} \mid \text{pos } y\} \rangle^{l,l}$.

Translating from λ_C to λ_H , we are moving from a simple type system to a rich one. The translation ϕ (essentially the same as Gronski & Flanagan’s) generates terms in λ_H with *raw* types— λ_H types with trivial refinements, corresponding to λ_C ’s simple types. Since the translation targets raw types, the type preservation theorem is stated as “if $\Gamma \vdash t : T$ then $[\Gamma] \vdash \phi(t) : [T]$ ” (see Section 6.1).

Whereas the difficulty with ψ is ensuring that the checks match up, the difficulty with ϕ is ensuring that the terms in λ_C and λ_H will blame the same labels. We deal with this problem by translating a single contract with two blame labels into two separate casts. Intuitively, the cast carrying the negative blame label will run all of the checks in negative positions in the contract, while the cast with the positive blame label will run the positive checks. We let

$$\phi(\langle c \rangle^{l,l'}) = \lambda x: [c]. \langle \phi(c) \Rightarrow [c] \rangle^{l'} (\langle [c] \Rightarrow \phi(c) \rangle^l x),$$

where the translation of contracts to refined types is:

$$\begin{aligned} \phi(\{x:B \mid t\}) &= \{x:B \mid \phi(t)\} \\ \phi(c_1 \mapsto c_2) &= \phi(c_1) \rightarrow \phi(c_2) \end{aligned}$$

The operation of casting into and out of raw types is a kind of “bulletproofing.” Bulletproofing maintains the raw-type invariant: the positive cast takes x out of $[c]$ and the negative cast returns it there. For example,

$$\langle \{x:\text{Int} \mid \text{nonzero } x\} \mapsto \{y:\text{Int} \mid \text{pos } y\} \rangle^{l,l'}$$

translates to the λ_H term

$$\begin{aligned} &\lambda f: [\text{Int} \rightarrow \text{Int}]. \\ &\langle \{x:\text{Int} \mid \text{nonzero } x\} \rightarrow \{y:\text{Int} \mid \text{pos } y\} \Rightarrow [\text{Int} \rightarrow \text{Int}] \rangle^{l'} \\ &(\langle [\text{Int} \rightarrow \text{Int}] \Rightarrow \{x:\text{Int} \mid \text{nonzero } x\} \rightarrow \{y:\text{Int} \mid \text{pos } y\} \rangle^l f). \end{aligned}$$

Unfolding the domain parts of the casts on f , the domain of the negative cast ensures that f ’s argument is nonzero with $\langle [\text{Int}] \Rightarrow \{x:\text{Int} \mid \text{nonzero } x\} \rangle^{l'}$; the domain of the positive cast does nothing, since $\langle \{x:\text{Int} \mid \text{nonzero } x\} \Rightarrow [\text{Int}] \rangle^l$ has no effect. Similarly, the codomain of the negative cast does nothing while the codomain of the positive cast checks that the result is positive. Separating the checks allows λ_H to keep track of blame labels, mimicking λ_C . This embodies the idea of contracts as pairs of projections (Findler, 2006). Note that bulletproofing is “overkill” at base type: for example, $\langle \{x:\text{Int} \mid \text{nonzero } x\} \rangle^{l,l'}$ translates to

$$\begin{aligned} &\lambda x: [\text{Int}]. \\ &\langle \{x:\text{Int} \mid \text{nonzero } x\} \Rightarrow [\text{Int}] \rangle^{l'} \\ &(\langle [\text{Int}] \Rightarrow \{x:\text{Int} \mid \text{nonzero } x\} \rangle^l x). \end{aligned}$$

Only the positive cast does anything—the negative cast into [Int] always succeeds. This asymmetry is consistent with λ_C , where base-type contracts also ignore the negative label.

Both ϕ and ψ preserve behavior in a strong sense: if $\Gamma \vdash t : B$, then either t and $\phi(t)$ both evaluate to the same constant k or they both raise $\uparrow l$ for the same l ; and conversely for ψ . Interestingly, we need to set up this behavioral correspondence *before* we can prove that the translations preserve well-typedness, because of the T_CHECKING and S_CHECKING rules.

4 The dependent languages

We now extend λ_C to dependent function contracts and λ_H to dependent functions. Very little needs to be changed in λ_C , since contracts and types barely interact; the changes to E_CDECOMP and T_FUNC are the important ones. Adding dependency to λ_H is more involved. In particular, adding contexts to the subtyping judgment entails adding contexts to SSUB_REFINE. To avoid a dangerous circularity, we define closing substitutions in terms of a separate type semantics. Additionally, the new F_CDECOMP rule has a slightly tricky (but necessary) asymmetry, explained below.

Dependent λ_C

Dependent λ_C has been studied since Findler & Felleisen (2002); it received a very thorough treatment (with an untyped host language) in Blume & McAllester (2006), was ported to Haskell by Hinze *et al.* (2006) and Chitil & Huch (2007), and was used as a specification language in Xu *et al.* (2009). Type soundness is not particularly difficult, since types and contracts are kept separate. Our formulation follows Findler & Felleisen (2002), with a few technical changes to make the proofs for ϕ easier.

The new T_REFINEC, T_FUNC, and E_CDECOMP rules in Figure 5 suffice to add dependency to λ_C . To help us work with the translations, we also make some small changes to the bindings in contexts, adding a new binding form to track the labels on a contract check throughout the contract well-formedness judgment. Note that T_FUNC adds $x:c_1^{l,l'}$ to the context when checking the codomain of a function contract, swapping blame labels. We add a new variable rule, T_VARC, that treats $x:c^{l,l'}$ as if it were its skeleton, $x:[c]$. While unnecessary for λ_C 's metatheory, this new binding form helps ϕ preserve types when translating from λ_H to picky λ_C ; see Section 6.1.

Two different variants of the E_CDECOMP rule can be found in the literature: they are *lax* and *picky*.⁴ The original rule in Findler & Felleisen (2002) is *lax* (like most other contract calculi): it does not recheck c_1 when substituting v' into c_2 . Hinze *et al.* (2006) choose instead to be *picky*, substituting $\langle c_1 \rangle^{l,l'} v'$ into c_2 because it makes their conjunction contract idempotent. We can show (straightforwardly) that both enjoy standard progress and preservation properties. Below, we consider translations to and from both dialects of λ_C : picky λ_C using only E_CDECOMP_PICKY in Sections 5.1 and 6.2, and lax λ_C using only E_CDECOMP_LAX in Sections 5.2 and 6.1.

⁴ *Indy blame* is a variant of picky (Dimoulas *et al.*, 2011). We do not consider it here, since it doesn't affect *whether or not* blame is raised, only *which* blame.

Syntax for dependent λ_C

$T ::= B \mid T_1 \rightarrow T_2$	types
$c ::= \{x:B \mid t\} \mid x:c_1 \mapsto c_2$	contracts
$\Gamma ::= \emptyset \mid \Gamma, x:T \mid \Gamma, x:c^{l,l'}$	typing contexts
$t ::= x \mid k \mid \lambda x:T_1. t_2 \mid t_1 t_2 \mid \uparrow l \mid \langle c \rangle^{l,l'} \mid \langle \{x:B \mid t_1\}, t_2, k \rangle^l$	terms
$v ::= k \mid \lambda x:T_1. t_2 \mid \langle c \rangle^{l,l'} \mid \langle x:c_1 \mapsto c_2 \rangle^{l,l'} v$	values
$r ::= v \mid \uparrow l$	results
$E ::= [] t \mid v [] \mid \langle \{x:B \mid t\}, [], k \rangle^l$	evaluation contexts

Operational semantics for λ_C

$(\lambda x:T_1. t_2) v \longrightarrow_c t_2 \{x := v\}$	E_BETA
$k v \longrightarrow_c \llbracket k \rrbracket(v)$	E_CONST
$\langle \{x:B \mid t\} \rangle^{l,l'} k \longrightarrow_c \langle \{x:B \mid t\}, t \{x := k\}, k \rangle^l$	E_CCHECK
$\langle x:c_1 \mapsto c_2 \rangle^{l,l'} v' \longrightarrow_c \langle c_2 \{x := v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l,l'} v'))$	E_CDECOMPLAX
$\langle x:c_1 \mapsto c_2 \rangle^{l,l'} v' \longrightarrow_c \langle c_2 \{x := \langle c_1 \rangle^{l,l'} v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l,l'} v'))$	E_CDECOMPICKY
$\langle \{x:B \mid t\}, \text{true}, k \rangle^l \longrightarrow_c k$	E_OK
$\langle \{x:B \mid t\}, \text{false}, k \rangle^l \longrightarrow_c \uparrow l$	E_FAIL
$E[\uparrow l] \longrightarrow_c \uparrow l$	E_BLAZE
$E[t_1] \longrightarrow_c E[t_2] \text{ when } t_1 \longrightarrow_c t_2$	E_COMPAT

Contract erasure

$$\llbracket \{x:B \mid t\} \rrbracket = B \quad \llbracket x:c_1 \mapsto c_2 \rrbracket = \llbracket c_1 \rrbracket \rightarrow \llbracket c_2 \rrbracket$$

Typing rules for dependent λ_C

$\boxed{\vdash \Gamma}$		
$\frac{}{\vdash \emptyset} \text{ T_EMPTY}$	$\frac{\vdash \Gamma}{\vdash \Gamma, x:T} \text{ T_EXTVART}$	$\frac{\vdash \Gamma \quad \Gamma \vdash^{l,l'} c : [c]}{\vdash \Gamma, x:c^{l,l'}} \text{ T_EXTVARC}$
$\boxed{\Gamma \vdash t : T}$		
$\frac{x:T \in \Gamma}{\Gamma \vdash x : T} \text{ T_VART}$	$\frac{x:c^{l,l'} \in \Gamma}{\Gamma \vdash x : [c]} \text{ T_VARC}$	$\frac{}{\Gamma \vdash k : \text{ty}_c(k)} \text{ T_CONST}$
$\frac{\Gamma, x:T_1 \vdash t_2 : T_2}{\Gamma \vdash \lambda x.T_1. t_2 : T_1 \rightarrow T_2} \text{ T_LAM}$	$\frac{\Gamma \vdash t_1 : T_1 \rightarrow T_2 \quad \Gamma \vdash t_2 : T_1}{\Gamma \vdash t_1 t_2 : T_2} \text{ T_APP}$	
$\frac{\Gamma \vdash^{l,l'} c : T}{\Gamma \vdash \langle c \rangle^{l,l'} : T \rightarrow T} \text{ T_CONTRACT}$	$\frac{}{\Gamma \vdash \uparrow l : T} \text{ T_BLAME}$	
$\frac{\vdash \Gamma \quad \emptyset \vdash k : B \quad \emptyset \vdash t_2 : \text{Bool} \quad \emptyset \vdash^{l,l'} \{x:B \mid t_1\} : B \quad t_2 \longrightarrow_c^* \text{true} \text{ implies } t_1 \{x := k\} \longrightarrow_c^* \text{true}}{\Gamma \vdash \langle \{x:B \mid t_1\}, t_2, k \rangle^l : B} \text{ T_CHECKING}$		
$\boxed{\Gamma \vdash^{l,l'} c : T}$		
$\frac{\Gamma, x:B \vdash t : \text{Bool}}{\Gamma \vdash^{l,l'} \{x:B \mid t\} : B} \text{ T_REFINEC}$	$\frac{\Gamma \vdash^{l,l'} c_1 : T_1 \quad \Gamma, x:c_1^{l,l'} \vdash^{l,l'} c_2 : T_2}{\Gamma \vdash^{l,l'} x:c_1 \mapsto c_2 : T_1 \rightarrow T_2} \text{ T_FUNC}$	

Fig. 5. Syntax and semantics for dependent λ_C

We make a standard assumption about constant denotations being well typed: if $\Gamma \vdash k v : T$ then $\Gamma \vdash \llbracket k \rrbracket(v) : T$.

4.1 Theorem [Progress]: If $\emptyset \vdash t : T$ then either $t \longrightarrow_c t'$ or $t = r$ (i.e., $t = v$ or $t = \uparrow l$).

Proof

By induction on the typing derivation. \square

For preservation, we prove a confluence lemma, along with weakening and substitution. Note that our substitution lemma must now also cover contracts, since they are no longer closed.

4.2 Lemma [Determinacy]: If $t \longrightarrow_c t'$ and $t \longrightarrow_c t''$, then $t' = t''$.

4.3 Corollary [Coevaluation]: If $t \longrightarrow_c^* r$ and $t \longrightarrow_c^* t'$, then $t' \longrightarrow_c^* r$.

4.4 Lemma [Term and contract substitution]: If $\emptyset \vdash v : T'$, then

1. if $\Gamma, x:T', \Gamma' \vdash t : T$ then $\Gamma, \Gamma' \{x := v\} \vdash t \{x := v\} : T$, and
2. if $\Gamma, x:T', \Gamma' \vdash^{l,l'} c : T$ then $\Gamma, \Gamma' \{x := v\} \vdash^{l,l'} c \{x := v\} : T$.

Proof

By mutual induction on the typing derivations for t and c . \square

We omit the proof for $x:c^{l,l'}$ bindings, which is similar.

4.5 Theorem [Preservation]: If $\emptyset \vdash t : T$ and $t \longrightarrow_c t'$ then $\emptyset \vdash t' : T$.

Proof

By induction on the typing derivation. This proof is straightforward because typing and contracts hardly interact. \square

Dependent λ_H

Now we come to the challenging part: dependent λ_H and its proof of type soundness.⁵ These results require the most complex metatheory in the paper, because we need some strong properties about call-by-value evaluation⁶ The full definitions are in Figure 7;

First, we enrich the type system with dependent function types, $x:S_1 \rightarrow S_2$, where x may appear in S_2 . A new dependent application rule, S_APP, substitutes the argument

⁵ The proof of type soundness for this system is significantly different from the soundness proof in Knowles & Flanagan (2010), where the operational semantics of λ_H is full, nondeterministic β -reduction. At first glance, it might seem that our theorems follow directly from the results for Knowles and Flanagan's language, since CBV is a restriction of full β -reduction. However, the reduction relation is used in the type system (in rule S_IMP), so the type systems for the two languages are not the same. For example, suppose the term *bad* contains a cast that fails. In our system $\{y:B \mid \text{true}\}$ is not a subtype of $\{y:B \mid (\lambda x:S. \text{true}) \text{ bad}\}$ because the contract evaluates to blame. However, the subtyping does hold in the Knowles and Flanagan system because the predicate reduces to true.

⁶ The benefit of a CBV semantics is a better treatment of blame. By contrast, Knowles & Flanagan (2010) cannot treat failed casts as exceptions because that would destroy confluence. They treat them as stuck terms.

Syntax for dependent λ_H

$S ::= \{x:B \mid s_1\} \mid x:S_1 \rightarrow S_2$	types/contracts
$\Delta ::= \emptyset \mid \Delta, x:S$	typing contexts
$s ::= x \mid k \mid \lambda x:S_1. s_2 \mid s_1 s_2 \mid \uparrow l \mid \langle S_1 \Rightarrow S_2 \rangle^l \mid \langle \{x:B \mid s_1\}, s_2, k \rangle^l$	terms
$w ::= k \mid \lambda x:S_1. s_2 \mid \langle S_1 \Rightarrow S_2 \rangle^l \mid \langle x:S_{11} \rightarrow S_{12} \Rightarrow x:S_{21} \rightarrow S_{22} \rangle^l w$	values
$q ::= w \mid \uparrow l$	results
$F ::= [] s \mid w [] \mid \langle \{x:B \mid s\}, [], k \rangle^l$	evaluation contexts

Operational semantics for dependent λ_H

$s_1 \rightsquigarrow_h s_2$	
$(\lambda x:S_1. s_2) w_2 \rightsquigarrow_h s_2 \{x := w_2\}$	F_BETA
$k w \rightsquigarrow_h \llbracket k \rrbracket(w)$	F_CONST
$\langle \{x:B \mid s_1\} \Rightarrow \{x:B \mid s_2\} \rangle^l k \rightsquigarrow_h \langle \{x:B \mid s_2\}, s_2 \{x := k\}, k \rangle^l$	F_CCHECK
$\langle x:S_{11} \rightarrow S_{12} \Rightarrow x:S_{21} \rightarrow S_{22} \rangle^l w \rightsquigarrow_h \langle S_{12} \{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w' \} \Rightarrow S_{22} \{x := w'\} \rangle^l (w (\langle S_{21} \Rightarrow S_{11} \rangle^l w'))$	F_CDECOMP
$\langle \{x:B \mid s\}, \text{true}, k \rangle^l \rightsquigarrow_h k$	F_OK
$\langle \{x:B \mid s\}, \text{false}, k \rangle^l \rightsquigarrow_h \uparrow l$	F_FAIL
$s_1 \longrightarrow_h s_2$	
$\frac{s_1 \rightsquigarrow_h s_2}{s_1 \longrightarrow_h s_2} \text{ F_REDUCE}$	
$\frac{s_1 \longrightarrow_h s_2}{F[s_1] \longrightarrow_h F[s_2]} \text{ F_COMPAT}$	
$\frac{}{F[\uparrow l] \longrightarrow_h \uparrow l} \text{ F_BLAME}$	

Fig. 6. Syntax and operational semantics for dependent λ_H

into the result type of the application. We generalize SWF_REFINE to allow refinement-type predicates that use variables from the enclosing context. SWF_FUN adds the bound variable to the context when checking the codomain of function types. In SSUB_FUN, subtyping for dependent function types remains contravariant, but we also add the argument variable to the context with the smaller type. This is similar to the function subtyping rule of $F_{<}$ (Cardelli *et al.*, 1991).

We need to be careful when implementing higher-order dependent casts in the rule F_CDECOMP. As the cast decomposes, the variables in the codomain types of such a cast must be replaced. However, this substitution is asymmetric; on one side, we cast the argument and on the other we do not. This behavior is required for type soundness. For suppose we have $\Delta \vdash x:S_{11} \rightarrow S_{12}$ and $\Delta \vdash x:S_{21} \rightarrow S_{22}$ with equal skeletons, and values $\Delta \vdash w : (x:S_{11} \rightarrow S_{12})$ and $\Delta \vdash w' : S_{21}$. Then $\Delta \vdash (\langle x:S_{11} \rightarrow S_{12} \Rightarrow x:S_{21} \rightarrow S_{22} \rangle^l w) w' : S_{22} \{x := w'\}$. When we decompose the cast, we must make *some* substitution into S_{12} and S_{22} , but which? It is clear that we must substitute w' into S_{22} , since the original application has type $S_{22} \{x := w'\}$. Decomposing the cast will produce the inner application $\Delta \vdash w (\langle S_{21} \Rightarrow S_{11} \rangle^l w') : S_{12} \{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w'\}$; in order to apply the codomain cast to this term, we must substitute $\langle S_{21} \Rightarrow S_{11} \rangle^l w'$ into S_{12} . This calculation determines the form of F_CDECOMP.

While the operational semantics change only in F_CDECOMP, we've split the evaluation relation into two parts: reductions \rightsquigarrow_h and steps \longrightarrow_h . This is a technical change that allows us to factor our proofs more cleanly (particularly for the parallel reduction proofs).

Type erasure and size

$$\begin{array}{ll}
[-] : S \rightarrow T & | - | : S \rightarrow \mathbb{N} \\
[\{x:B \mid s\}] = B & |\{x:B \mid s\}| = 1 \\
[x:S_1 \rightarrow S_2] = [S_1] \rightarrow [S_2] & |x:S_1 \rightarrow S_2| = 1 + |S_1| + |S_2|
\end{array}$$

Typing rules

$$\begin{array}{c}
\boxed{\vdash \Delta} \\
\\
\frac{}{\vdash \emptyset} \quad \text{S_EMPTY} \qquad \frac{\vdash \Delta \quad \Delta \vdash S}{\vdash \Delta, x:S} \quad \text{S_EXTVAR} \\
\\
\boxed{\Delta \vdash s : S} \\
\\
\frac{x:S \in \Delta}{\Delta \vdash x : S} \quad \text{S_VAR} \qquad \frac{}{\Delta \vdash k : \text{ty}_h(k)} \quad \text{S_CONST} \\
\\
\frac{\Delta \vdash S_1 \quad \Delta, x:S_1 \vdash s_2 : S_2}{\Delta \vdash \lambda x:S_1. s_2 : (x:S_1 \rightarrow S_2)} \quad \text{S_LAM} \qquad \frac{\Delta \vdash s_1 : (x:S_1 \rightarrow S_2) \quad \Delta \vdash s_2 : S_1}{\Delta \vdash s_1 s_2 : S_2 \{x := s_2\}} \quad \text{S_APP} \\
\\
\frac{\Delta \vdash S_1 \quad \Delta \vdash S_2 \quad [S_1] = [S_2]}{\Delta \vdash \langle S_1 \Rightarrow S_2 \rangle^l : S_1 \rightarrow S_2} \quad \text{S_CAST} \qquad \frac{\Delta \vdash s : S_1 \quad \Delta \vdash S_2 \quad \Delta \vdash S_1 <: S_2}{\Delta \vdash s : S_2} \quad \text{S_SUB} \\
\\
\frac{\vdash \Delta \quad \emptyset \vdash k : \{x:B \mid \text{true}\} \quad \emptyset \vdash s_2 : \{x:\text{Bool} \mid \text{true}\} \quad \emptyset \vdash \{x:B \mid s_1\} \quad \emptyset \vdash s_2 \supset s_1 \{x := k\}}{\Delta \vdash \langle \{x:B \mid s_1\}, s_2, k \rangle^l : \{x:B \mid s_1\}} \quad \text{S_CHECKING} \\
\\
\boxed{\Delta \vdash S} \\
\\
\frac{}{\Delta \vdash \{x:B \mid \text{true}\}} \quad \text{SWF_RAW} \qquad \frac{\Delta, x:\{x:B \mid \text{true}\} \vdash s : \{x:\text{Bool} \mid \text{true}\}}{\Delta \vdash \{x:B \mid s\}} \quad \text{SWF_REFINE} \qquad \frac{\Delta \vdash S_1 \quad \Delta, x:S_1 \vdash S_2}{\Delta \vdash x:S_1 \rightarrow S_2} \quad \text{SWF_FUN} \\
\\
\boxed{\Delta \vdash S_1 <: S_2} \\
\\
\frac{\Delta, x:\{x:B \mid \text{true}\} \vdash s_1 \supset s_2}{\Delta \vdash \{x:B \mid s_1\} <: \{x:B \mid s_2\}} \quad \text{SSUB_REFINE} \qquad \frac{\Delta \vdash S_{21} <: S_{11} \quad \Delta, x:S_{21} \vdash S_{12} <: S_{22}}{\Delta \vdash x:S_{11} \rightarrow S_{12} <: x:S_{21} \rightarrow S_{22}} \quad \text{SSUB_FUN} \\
\\
\boxed{\Delta \vdash s_1 \supset s_2} \\
\\
\frac{\forall \sigma. (\Delta \models \sigma \wedge \sigma(s_1) \longrightarrow_h^* \text{true}) \text{ implies } \sigma(s_2) \longrightarrow_h^* \text{true}}{\Delta \vdash s_1 \supset s_2} \quad \text{S_IMP}
\end{array}$$

$$\Delta \models \sigma \iff \forall x \in \text{dom}(\Delta). \sigma(x) \in \llbracket \sigma(\Delta(x)) \rrbracket$$

Fig. 7. Typing rules for dependent λ_H

Denotations of types and kinds

$$\begin{aligned}
s \in \llbracket \{x:B \mid s_0\} \rrbracket &\iff s \longrightarrow_h^* \uparrow l \vee (\exists k \in \mathcal{K}_B. s \longrightarrow_h^* k \wedge s_0\{x:=k\} \longrightarrow_h^* \text{true}) \\
s \in \llbracket x:S_1 \rightarrow S_2 \rrbracket &\iff \forall q \in \llbracket S_1 \rrbracket. s \ q \in \llbracket S_2\{x:=q\} \rrbracket \\
\{x:B \mid s\} \in \llbracket \star \rrbracket &\iff \forall k \in \mathcal{K}_B. s\{x:=k\} \in \llbracket \{x:\text{Bool} \mid \text{true}\} \rrbracket \\
x:S_1 \rightarrow S_2 \in \llbracket \star \rrbracket &\iff S_1 \in \llbracket \star \rrbracket \wedge \forall q \in \llbracket S_1 \rrbracket. S_2\{x:=q\} \in \llbracket \star \rrbracket
\end{aligned}$$

Semantic judgments

$$\begin{aligned}
\Delta \models S_1 <: S_2 &\iff \forall \sigma \text{ s.t. } \Delta \models \sigma, \llbracket \sigma(S_1) \rrbracket \subseteq \llbracket \sigma(S_2) \rrbracket \\
\Delta \models s : S &\iff \forall \sigma \text{ s.t. } \Delta \models \sigma, \sigma(s) \in \llbracket \sigma(S) \rrbracket \\
\Delta \models S &\iff \forall \sigma \text{ s.t. } \Delta \models \sigma, \sigma(S) \in \llbracket \star \rrbracket
\end{aligned}$$

Fig. 8. Type and kind semantics for dependent λ_H

The final change generalizes `SSUB_REFINE` to open terms. We must close these terms before we can compare their behavior, using closing substitutions σ and reading $\Delta \models \sigma$ as “ σ satisfies Δ ”.

Care is needed here to prevent the typing rules from becoming circular: the typing rule `S_SUB` references the subtyping judgment, the subtyping rule `SSUB_REFINE` references the implication judgment, and the single implication rule `S_IMP` has $\Delta \models \sigma$ in a negative position. To avoid circularity, $\Delta \models \sigma$ must not refer back to the other judgments. (The reader may wonder why this wasn’t a problem in λ_C , but notice that in λ_C , implication is only used in `T_CHECKING`—which has no (real) context. If we only needed implication in the `S_CHECKING` rule, we wouldn’t need contexts here, either—we can ensure that active checks only occur at the top-level, with an empty context. But the `SSUB_REFINE` subtyping rule refers to `S_IMP`, and subtyping may be used in arbitrary contexts.)

We can achieve this by building the syntactic rules on top of a denotational semantics for λ_H ’s types.⁷ The idea is that the semantics of a type is a set of closed terms defined independently of the syntactic typing relation, but that turns out to contain all closed well-typed terms of that type. Thus, in the definition of $\Delta \models \sigma$, we quantify over a somewhat larger set than strictly necessary—not just the syntactically well-typed terms of appropriate type (which are all the ones that will ever appear in programs), but all semantically well-typed ones.

The type semantics appears in Figure 8. It is defined by induction on type skeletons. For refinement types, terms must either go to blame or produce a constant that satisfies (all instances of) the given predicate. For function types, well typed arguments must yield well-typed results. By construction, these sets include only terminating terms that do not get stuck.

⁷ Knowles & Flanagan (2010) also introduce a type semantics, but it differs from ours in two ways. First, because they cannot treat blame as an exception (because their semantics is nondeterministic) they must restrict the terms in the semantics to be those that only get stuck at failed casts. They do so by requiring the terms to be well-typed in the simply typed lambda calculus after all casts have been erased. Secondly, their type semantics does not require strong normalization. However, it is not clear whether their language actually admits nontermination—they include a `fix` constant, but their semantic type soundness proof appears to break down in that case.

We must again make the assumption that constants have most-specific types: if $\llbracket \text{ty}_h(k) \rrbracket = B$ and $s\{x := k\} \longrightarrow_h^* \text{true}$ then $\emptyset \vdash \text{ty}_h(k) <: \{x:B \mid s\}$. We make some other, more standard assumptions, as well. Constants must have closed, well-formed types, and the types assigned must be well-formed. We require that constants are semantically well typed: $k \in \llbracket \text{ty}_h(k) \rrbracket$; this requirement is true by our “most-specific type” assumption at base types, but must be assumed at (first-order) functions types. Note that this rules out including fix as a constant, since our type semantics is inhabited only by strongly normalizing terms.⁸

We introduce a few facts about the type semantics before proving semantic type soundness.

4.6 Lemma [Determinacy]: If $s \longrightarrow_h s'$ and $s \longrightarrow_h s''$, then $s' = s''$.

4.7 Corollary [Coevaluation]: If $s \longrightarrow_h^* s'$ and $s \longrightarrow_h^* q$, then $s' \longrightarrow_h^* q$.

4.8 Lemma [Expansion and contraction of $\llbracket S \rrbracket$]: If $s \longrightarrow_h^* s'$, then $s' \in \llbracket S \rrbracket$ iff $s \in \llbracket S \rrbracket$.

Proof

By induction on $|S|$. \square

4.9 Lemma [Blame inhabits all types]: For all S , $\uparrow l \in \llbracket S \rrbracket$.

Proof

By induction on $|S|$. \square

4.10 Corollary [Nonemptiness]: For all S , there exists some q such that $q \in \llbracket S \rrbracket$.

The normal forms of \longrightarrow_h^* are of the form $q = w$ or $\uparrow l$.

4.11 Lemma [Strong normalization]: If $s \in \llbracket S \rrbracket$, then there exists a q such that $s \longrightarrow_h^* q$ —i.e., either $s \longrightarrow_h^* w$ or $s \longrightarrow_h^* \uparrow l$.

Proof

By induction on $|S|$.

$S = \{x:B \mid s_0\}$: Suppose $s \in \llbracket \{x:B \mid s_0\} \rrbracket$. By definition, either $s \longrightarrow_h^* w$ or $s \longrightarrow_h^* \uparrow l$, so s normalizes.

$S = x:S_1 \rightarrow S_2$: Suppose $s \in \llbracket x:S_1 \rightarrow S_2 \rrbracket$. We know that for any $q \in \llbracket S_1 \rrbracket$ that $s q \in \llbracket S_2 \{x := q\} \rrbracket$. Since $\llbracket S_1 \rrbracket$ is nonempty (by Lemma 4.10), let $q \in \llbracket S_1 \rrbracket$. By the IH, $s q \longrightarrow_h^* w$ or $s q \longrightarrow_h^* \uparrow l$. By the definition of evaluation contexts and \longrightarrow_h^* , the function position is evaluated first. If the application reduces to a value (i.e., $s q \longrightarrow_h^* w$), then first $s q \longrightarrow_h^* w' q$, and so $s \longrightarrow_h^* w'$. Alternatively, the application could reduce to blame (i.e., $s q \longrightarrow_h^* \uparrow l$). There are two ways for this to happen: either $s \longrightarrow_h^* \uparrow l$, or $s \longrightarrow_h^* w'$ and $q \longrightarrow_h^* \uparrow l$. In both cases, s normalizes. \square

We are interested relating our syntactic type system and the type semantics by *semantic type soundness*: if $\emptyset \vdash s : S$, then $s \in \llbracket S \rrbracket$. However, to prove this, we must generalize it. In the bottom of Figure 8, we define three *semantic judgements* that correspond to each of the three typing judgments. (Note that the third one requires the definition of a

⁸ We conjecture that a step-indexed logical relation (Ahmed, 2006) would allow us to consider nonterminating terms.

kind semantics that picks out well-behaved types—those whose embedded terms belong to the type semantics.) We then show that the typing judgments imply their semantic counterparts.

4.12 Theorem [Semantic type soundness]:

1. If $\Delta \vdash S_1 <: S_2$ then $\Delta \models S_1 <: S_2$.
2. If $\Delta \vdash s : S$ then $\Delta \models s : S$.
3. If $\Delta \vdash S$ then $\Delta \models S$.

Proof

Proof of (1) is in Lemma 4.14. Proof of (2) and (3) is in Lemma 4.21. \square

The first part follows by induction on the subtyping judgment.

4.13 Lemma [Trivial refinements of constants]: If $k \in \mathcal{K}_B$, then $k \in \llbracket \{x:B \mid \text{true}\} \rrbracket$.

4.14 Lemma [Semantic subtype soundness]: If $\Delta \vdash S_1 <: S_2$ then $\Delta \models S_1 <: S_2$.

Proof

By induction on the subtyping derivation.

SSUB_REFINE: We know $\Delta \vdash \{x:B \mid s_1\} <: \{x:B \mid s_2\}$, and must show the corresponding semantic subtyping. Inversion of this derivation gives us $\Delta, x:\{x:B \mid \text{true}\} \vdash s_1 \supset s_2$, which means:

$$\forall \sigma. ((\Delta, x:\{x:B \mid \text{true}\} \models \sigma \wedge \sigma(s_1) \longrightarrow_h^* \text{true}) \text{ implies } \sigma(s_2) \longrightarrow_h^* \text{true}) \quad (*)$$

We must show $\Delta \models \{x:B \mid s_1\} <: \{x:B \mid s_2\}$, i.e., that $\forall \sigma. (\Delta \models \sigma \text{ implies } \llbracket \{x:B \mid s_1\} \rrbracket \subseteq \llbracket \{x:B \mid s_2\} \rrbracket)$. Let σ be given such that $\Delta \models \sigma$. Suppose $s \in \llbracket \sigma(\{x:B \mid s_1\}) \rrbracket$. By definition, either s goes to $\uparrow l$, or it goes to $k \in \mathcal{K}_B$ such that $s_1 \{x := k\} \longrightarrow_h^* \text{true}$. In the former case, $\uparrow l \in \llbracket \{x:B \mid s_2\} \rrbracket$ by definition. So consider the latter case, where $s \longrightarrow_h^* k$.

We already know that $k \in \mathcal{K}_B$, so it remains to see that: $\sigma(s_2) \{x := k\} \longrightarrow_h^* \text{true}$. We know by assumption that $\sigma(s_1) \{x := k\} \longrightarrow_h^* \text{true}$. By Lemma 4.13, $k \in \llbracket \{x:B \mid \text{true}\} \rrbracket$.

Now observe that $\Delta, x:\{x:B \mid \text{true}\} \models \sigma \{x := k\}$. Since $\sigma'(s_1) \longrightarrow_h^* \text{true}$, we can conclude that $\sigma'(s_2) \longrightarrow_h^* \text{true}$ by our assumption (*). This completes this case.

SSUB_FUN: $\Delta \vdash (x:S_{11} \rightarrow S_{12}) <: (x:S_{21} \rightarrow S_{22})$; by the IH, we have $\Delta \models S_{21} <: S_{11}$ and $\Delta, x:S_{21} \models S_{12} <: S_{22}$. We must show that $\Delta \models (x:S_{11} \rightarrow S_{12}) <: (x:S_{21} \rightarrow S_{22})$.

Let $\Delta \models \sigma$ and $s \in \llbracket \sigma(x:S_{11} \rightarrow S_{12}) \rrbracket$, for some σ . We must show, for all q , that if $q \in \llbracket \sigma(S_{21}) \rrbracket$ then $s q \in \llbracket \sigma(S_{22}) \{x := q\} \rrbracket$.

Let $q \in \llbracket \sigma(S_{21}) \rrbracket$. Then $q \in \llbracket \sigma(S_{11}) \rrbracket$. Since $s \in \llbracket \sigma(x:S_{11} \rightarrow S_{12}) \rrbracket$, we know that $s q \in \llbracket \sigma(S_{12}) \{x := q\} \rrbracket$. Finally, since $\Delta, x:S_{21} \models S_{12} <: S_{22}$ and $\Delta, x:S_{21} \models \sigma \{x := q\}$, we can conclude that $s q \in \llbracket \sigma(S_{22}) \{x := q\} \rrbracket$, and so $s \in \llbracket \sigma(x:S_{21} \rightarrow S_{22}) \rrbracket$. \square

However, we run into some complications with the second and third parts (which must be proven together). The crux of the difficulty lies with the S_APP rule. Suppose the application $s_1 s_2$ was well typed and $s_1 \in \llbracket x:S_1 \rightarrow S_2 \rrbracket$ and $s_2 \in \llbracket S_1 \rrbracket$. According to S_APP , the application's type is $S_2 \{x := s_2\}$. By the type semantics defined in Figure 8, if $s_1 \in \llbracket x:S_1 \rightarrow S_2 \rrbracket$, then $s_1 q \in \llbracket S_2 \{x := q\} \rrbracket$ for any $q \in \llbracket S_1 \rrbracket$. Sadly, s_2 isn't necessarily a result! We do know, however, that $s_2 \in \llbracket S_1 \rrbracket$, so $s_2 \longrightarrow_h^* q_2$ by strong normalization (Lemma 4.11). We need to ask, then, how the type semantics of $S_2 \{x := s_2\}$ and $S_2 \{x := q_2\}$ relate. (One

might think that we can solve this by changing the type semantics to quantify over terms, not results. But this just pushes the problem to the S_LAM case.)

We can show that the two type semantics are in fact equal using a parallel reduction technique. We define a parallel reduction relation \Rightarrow on terms and types in Figure 9 that allows redexes in different parts of a term (or type) to be reduced in the same step, and we prove that types that parallel-reduce to each other—like $S_2\{x := s_2\}$ and $S_2\{x := q_2\}$ —have the same semantics. The definition of parallel reduction is standard, though we need to be careful to make it respect our call-by-value reduction order: the β -redex $(\lambda x:S_1. s_1) s_2$ should not be contracted unless s_2 is a value, since doing so can change the order of effects. (Other redexes within s_1 and s_2 can safely reduce.)⁹ The proof requires a longish sequence of technical lemmas that essentially show that \Rightarrow commutes with \longrightarrow_h^* . Since the proofs require fussy symbol manipulation, we’ve done these proofs in Coq. Our development is available at http://www.cis.upenn.edu/~mgree/papers/lambdah_parred.tgz. We restate the critical results here.

Lemma [Substitution of parallel-reducing terms, Lemma A3 in `thy.v`]:

If $w \Rightarrow w'$, then

1. if $s \Rightarrow s'$ then $s\{x := w\} \Rightarrow s'\{x := w'\}$, and
2. if $S \Rightarrow S'$ then $S\{x := w\} \Rightarrow S'\{x := w'\}$.

Lemma [Parallel reduction implies coevaluation, Lemma A20 in `thy.v`]:

If $s_1 \Rightarrow s_2$ then $s_1 \longrightarrow_h^* k$ iff $s_2 \longrightarrow_h^* k$. Similarly, $s_1 \longrightarrow_h^* \uparrow l$ iff $s_2 \longrightarrow_h^* \uparrow l$.

An alternative strategy would be to use \Rightarrow in the typing rules and \longrightarrow_h in the operational semantics. This would simplify some of our metatheory, but it would complicate the specification of the language. Using \longrightarrow_h in the typing rules gives a clearer intuition and keeps the core system small.

4.17 Lemma [Single parallel reduction preserves type semantics]:

If $S_1 \Rightarrow S_2$ then $\llbracket S_1 \rrbracket = \llbracket S_2 \rrbracket$.

Proof

By induction on $|S_1|$ (which is equal to $|S_2|$), with a case analysis on the final rule used to show $S_1 \Rightarrow S_2$. \square

4.18 Corollary: [Parallel reduction preserves type semantics] If $S_1 \Rightarrow^* S_2$ then $\llbracket S_1 \rrbracket = \llbracket S_2 \rrbracket$.

4.19 Lemma [Partial semantic substitution]: If $\Delta_1, x:S', \Delta_2 \models s : S$, and $\Delta_1, x:S', \Delta_2 \models S$, and $\Delta_1 \models s' : S'$ then $\Delta_1, \Delta_2\{x := s'\} \models s\{x := s'\} : S\{x := s'\}$ and $\Delta_1, \Delta_2\{x := s'\} \models S\{x := s'\}$.

Proof

By the definition of $\Delta \models \sigma$. \square

⁹ We conjecture that the reflexive transitive closure of a similar “CBV-respecting” variant of full β -reduction could be used in place of our parallel reduction. It is not clear whether it would lead to shorter proofs.

Parallel reduction for dependent λ_H

$$\begin{array}{c}
\boxed{s_1 \Rightarrow s_2} \\
\\
\frac{}{s \Rightarrow s} \text{FP_REFL} \quad \frac{w \Rightarrow w'}{k \ w \Rightarrow \llbracket k \rrbracket(w')} \text{FP_RCONST} \quad \frac{s_{12} \Rightarrow s'_{12} \quad w_2 \Rightarrow w'_2}{(\lambda x:S. s_{12}) \ w_2 \Rightarrow s'_{12}\{x := w'_2\}} \text{FP_RBETA} \\
\\
\frac{s_2 \Rightarrow s'_2}{\langle \{x:B \mid s_1\} \Rightarrow \{x:B \mid s_2\} \rangle^l \ k \Rightarrow \langle \{x:B \mid s'_2\}, s'_2\{x := k\}, k \rangle^l} \text{FP_RCCHECK} \\
\\
\frac{}{\langle \{x:B \mid s\}, \text{true}, k \rangle^l \Rightarrow k} \text{FP_ROK} \quad \frac{}{\langle \{x:B \mid s\}, \text{false}, k \rangle^l \Rightarrow \uparrow l} \text{FP_RFAIL} \\
\\
\frac{S_{11} \Rightarrow S'_{11} \quad S_{12} \Rightarrow S'_{12} \quad S_{21} \Rightarrow S'_{21} \quad S_{22} \Rightarrow S'_{22} \quad w_1 \Rightarrow w'_1 \quad w_2 \Rightarrow w'_2}{\langle \{x:S_{11} \rightarrow S_{12} \Rightarrow x:S_{21} \rightarrow S_{22}\}^l \ w_1 \rangle \ w_2 \Rightarrow \langle S'_{12}\{x := \langle S'_{21} \Rightarrow S'_{11} \rangle^l \ w'_2\} \Rightarrow S'_{22}\{x := w'_2\}\rangle^l \ (w'_1 \ (\langle S'_{21} \Rightarrow S'_{11} \rangle^l \ w'_2))} \text{FP_RCDECOMP} \\
\\
\frac{S_1 \Rightarrow S'_1 \quad s_{12} \Rightarrow s'_{12}}{\lambda x:S_1. s_{12} \Rightarrow \lambda x:S'_1. s'_{12}} \text{FP_LAM} \quad \frac{s_1 \Rightarrow s'_1 \quad s_2 \Rightarrow s'_2}{s_1 \ s_2 \Rightarrow s'_1 \ s'_2} \text{FP_APP} \quad \frac{S_1 \Rightarrow S'_1 \quad S_2 \Rightarrow S'_2}{\langle S_1 \Rightarrow S_2 \rangle^l \Rightarrow \langle S'_1 \Rightarrow S'_2 \rangle^l} \text{FP_CAST} \\
\\
\frac{S \Rightarrow S' \quad s \Rightarrow s'}{\langle S, s, k \rangle^l \Rightarrow \langle S', s', k \rangle^l} \text{FP_CHECK} \quad \frac{}{F[\uparrow l] \Rightarrow \uparrow l} \text{FP_BLAME} \\
\\
\boxed{S_1 \Rightarrow S_2} \\
\\
\frac{}{S \Rightarrow S} \text{FP_SREFL} \quad \frac{s \Rightarrow s'}{\{x:B \mid s\} \Rightarrow \{x:B \mid s'\}} \text{FP_SREFINE} \quad \frac{S_1 \Rightarrow S'_1 \quad S_2 \Rightarrow S'_2}{x:S_1 \rightarrow S_2 \Rightarrow x:S'_1 \rightarrow S'_2} \text{FP_SFUN}
\end{array}$$

Fig. 9. Parallel reduction for dependent λ_H

4.20 Lemma [Semantic typing for casts]: If $\Delta \models S_1$ and $\Delta \models S_2$ and $\lfloor S_1 \rfloor = \lfloor S_2 \rfloor$, then $\Delta \models \langle S_1 \Rightarrow S_2 \rangle^l : x:S_1 \rightarrow S_2$ for fresh x .

Proof

By induction on $|S_1| = |S_2|$, going by cases on the shape of S_2 . Let $\Delta \models \sigma$; we show that $\sigma(\langle S_1 \Rightarrow S_2 \rangle^l) \in \llbracket \sigma(S_1 \rightarrow S_2) \rrbracket$.

$S_2 = \{x:B \mid \sigma(s_2)\}$: Let $q \in \llbracket \sigma(S_1) \rrbracket$. If $q = \uparrow l'$, then the applied cast goes to $\uparrow l'$, and we are done by Lemma 4.9. So $q = k \in \mathcal{K}_B$. By F_CCHECK $\langle S_1 \Rightarrow \{x:B \mid \sigma(s_2)\} \rangle^l \ k \rightarrow_h \langle \{x:B \mid \sigma(s_2)\}, \sigma(s_2)\{x := k\}, k \rangle^l$. By the well-kinding of S_2 , we know that $\sigma(s_2)\{x := k\} \in \llbracket \{x:\text{Bool} \mid \text{true}\} \rrbracket$, so by strong normalization (Lemma 4.11), the predicate in the active check goes to blame or to a value. If it goes to blame, we are done. If it goes to a value, that value must be true or false. If it goes to false, then the whole term goes to blame and we are done. If it goes to true, then the check will step to k . But $\sigma(s_2)\{x := k\} \rightarrow_h^* \text{true}$, so $k \in \llbracket \sigma(\{x:B \mid s_2\}) \rrbracket$ by definition. Expansion (Lemma 4.8) completes the proof.

$S_2 = x:S_{21} \rightarrow S_{22}$: We must have $S_1 = x:S_{11} \rightarrow S_{12}$. Let $q \in \llbracket \sigma(S_1) \rrbracket$; if it is blame we're done by Lemma 4.9, so let it be a value w . Let $q' \in \llbracket \sigma(S_{21}) \rrbracket$; if it is blame we're done, so let it be a value w' . By F_CDECOMP:

$$\langle \sigma(S_{12})\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l \ w'\} \Rightarrow \sigma(S_{22})\{x := w'\} \rangle^l \ (w \ (\langle \sigma(S_{21}) \Rightarrow \sigma(S_{11}) \rangle^l \ w'))$$

By the IH, $\langle \sigma(S_{21}) \Rightarrow \sigma(S_{11}) \rangle^l$ is semantically well-typed, so $\langle \sigma(S_{21}) \Rightarrow \sigma(S_{11}) \rangle^l w' \in \llbracket \sigma(S_{11}) \rrbracket$. By strong normalization (Lemma 4.11), this term reduces (and therefore parallel reduces, by Lemma A4) to some q'' .

We know that $w q'' \in \llbracket \sigma(S_{12})\{x := q''\} \rrbracket$ by assumption. Using parallel reduction (Corollary 4.18), we have $\llbracket \sigma(S_{12})\{x := q''\} \rrbracket = \llbracket \sigma(S_{12})\{x := \langle \sigma(S_{21}) \Rightarrow \sigma(S_{11}) \rangle^l w'\} \rrbracket$.

Before applying the IH, we note that $\Delta \models S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w'\}$ and $\Delta \models S_{22}\{x := w'\}$ by Lemma 4.19. Then, by the IH we see that $\langle \sigma(S_{12})\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w'\} \Rightarrow \sigma(S_{22})\{x := w'\} \rangle^l$ is semantically well-kinded, so

$$\langle \sigma(S_{12})\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w'\} \Rightarrow \sigma(S_{22})\{x := w'\} \rangle^l (w w'') \in \llbracket \sigma(S_{22})\{x := w'\} \rrbracket \quad \square$$

4.21 Lemma [Semantic type soundness]:

1. If $\Delta \vdash s : S$ then $\Delta \models s : S$.
2. If $\Delta \vdash S$ then $\Delta \models S$.

Proof

By induction on the typing and well-formedness derivations, using Corollary 4.18 in the S_APP case and Lemma 4.14 in the S_SUB case. \square

Theorem 4.12 gives us type soundness, and it combines with Lemma 4.11 for an even stronger result: well-typed programs always evaluate to values of appropriate (semantic) type.

While one can prove progress and preservation theorems, we omit them: we already have type soundness. Our later proofs will require standard weakening and substitution lemmas, though, so we prove them now.

4.22 Lemma [Weakening]: If $\Delta \vdash s : S$ and $\Delta \vdash S$, and $\Delta \cap \Delta' = \emptyset$ with $\vdash \Delta, \Delta'$, then $\Delta, \Delta' \vdash s : S$ and $\Delta, \Delta' \vdash S$.

Proof

By straightforward induction on s and $|S|$; we reuse the (critical) context well-formedness derivation in the S_CHECKING case. \square

The substitution lemma has one complication: the operational judgment S_IMP requires the semantic type soundness theorem to show that a syntactically well-typed term can be used in a closing substitution. It is otherwise straightforward.

4.23 Lemma [Substitution (implication)]: If $\Delta_1, x:S, \Delta_2 \vdash s_1 \supset s_2$ and $\Delta_1 \vdash s : S$, then $\Delta_1, \Delta_2\{x := s\} \vdash s_1\{x := s\} \supset s_2\{x := s\}$.

Proof

Direct, unfolding the closing substitutions. \square

4.24 Lemma [Substitution (subtyping)]: If $\Delta_1, x:S, \Delta_2 \vdash S_1 <: S_2$ and $\Delta_1 \vdash s : S$, then $\Delta_1, \Delta_2\{x := s\} \vdash S_1\{x := s\} <: S_2\{x := s\}$.

Proof

By induction on the subtyping derivation. \square

4.25 Lemma [Substitution (typing and well-formedness)]: If $\Delta_1 \vdash s : S$ then

Result correspondence

$$\boxed{r \approx q : T}$$

$$k \approx k : B \iff k \in \mathcal{K}_B$$

$$v \approx w : T_1 \rightarrow T_2 \iff \forall t \sim s : T_1. v t \sim w s : T_2$$

$$\uparrow l \approx \uparrow l : T$$

Term correspondence

$$\boxed{t \sim s : T}$$

$$t \sim s : T \iff t \longrightarrow_c^* r \wedge s \longrightarrow_h^* q \wedge r \approx q : T$$

Fig. 10. A blame-exact result/term correspondence

1. if $\Delta_1, x:S, \Delta_2 \vdash s_1 : S_1$ then $\Delta_1, \Delta_2 \{x := s\} \vdash s_1 \{x := s\} : S_1 \{x := s\}$,
2. if $\Delta_1, x:S, \Delta_2 \vdash S_1$ then $\Delta_1, \Delta_2 \{x := s\} \vdash S_1 \{x := s\}$, and
3. if $\vdash \Delta_1, x:S, \Delta_2$ then $\vdash \Delta_1, \Delta_2 \{x := s\}$.

Proof

By mutual induction on the typing derivations. \square

5 Exact translations

Translations moving left on the axis of blame—from picky λ_C to λ_H , and from λ_H to lax λ_C —are exact. That is, we can show a tight behavioral correspondence between terms and their translations (see Figure 10). We read $t \sim s : T$ as “ t corresponds with s at type T ”.

Our correspondence is a standard logical relation, defined in two intertwined parts: a relation on results, $r \approx q : T$ and its closure with respect to evaluation, $t \sim s : T$. The term correspondence is defined directly: terms correspond when they reduce to corresponding results. The result correspondence is defined inductively over λ_C ’s simple types. Blame corresponds to itself at any type. At B , constants in \mathcal{K}_B correspond to themselves; results at $T_1 \rightarrow T_2$ correspond when they applying them to corresponding *terms* yields corresponding *terms*. Stratifying the definition this way simplifies some of our proofs later. We call this correspondence *exact* because terms corresponding at base type yield identical results.

Note that we define the correspondence here on closed (or harmlessly open) terms. In the following two sections, we’ll define translation specific extensions of the correspondence to open terms and contracts.

5.1 Translating picky λ_C to λ_H : dependent ϕ

We define the full ϕ for the dependent calculi in Figure 11. In the dependent case, we need to translate *derivations* of well-formedness and well-typing of λ_C contexts, terms, and contracts into λ_H contexts, terms, and types. We translate derivations to ensure type preservation, translating T_VART and T_VARC derivations differently: we leave variables of simple type alone, but we cast variables bound to contracts.

To see why we need this distinction, consider the function contract $f : (x : \{x : \text{Int} \mid \text{pos } x\}) \mapsto \{y : \text{Int} \mid \text{true}\} \mapsto \{z : \text{Int} \mid f \ 0 = 0\}$. Note that this contract is well-formed in λ_C , but that the codomain “abuses” the bound variable. A naive translation will *not* be well-typed in λ_H : $f \ 0$ will not be typeable when f has type $x : \{x : \text{Int} \mid \text{pos } x\} \rightarrow [\text{Int}]$, since f only accepts positive arguments. The problem is that SWF_FUN can add a (possibly refined) type to

$$\begin{array}{ll}
\textbf{Contexts} & \phi : (\vdash \Gamma) \rightarrow \Delta \\
& \phi(\vdash \emptyset) = \emptyset \\
& \phi(\vdash \Gamma, x:T) = \phi(\vdash \Gamma, x:[T]) \\
& \phi(\vdash \Gamma, x:c^{l,l'}) = \phi(\vdash \Gamma, x:\phi(\Gamma \vdash^{l,l'} c : [c])) \\
\\
\textbf{Terms} & \phi : (\Gamma \vdash t : T) \rightarrow s \\
& \phi(\Gamma_1, x:T, \Gamma_2 \vdash x : T) = x \\
& \phi(\Gamma_1, x:c^{l,l'}, \Gamma_2 \vdash x : [c]) = \langle \phi(\Gamma_1 \vdash^{l,l'} c : [c]) \Rightarrow [c] \rangle^{l'} x \\
& \phi(\Gamma \vdash k : T) = k \\
& \phi(\Gamma \vdash \lambda x:T_1. t_2 : T_1 \rightarrow T_2) = \lambda x:[T_1]. \phi(\Gamma, x:T_1 \vdash t_2 : T_2) \\
& \phi(\Gamma \vdash t_1 t_2 : T_2) = \phi(\Gamma \vdash t_1 : T_1 \rightarrow T_2) \phi(\Gamma \vdash t_2 : T_1) \\
& \phi(\Gamma \vdash \uparrow l : T) = \uparrow l \\
& \phi(\emptyset \vdash \langle c, t, k \rangle^l : B) = \langle \phi(\emptyset \vdash^{l,l'} c : B), \phi(\emptyset \vdash t : \text{Bool}), k \rangle^l \\
& \phi(\Gamma \vdash \langle c \rangle^{l,l'} : T) = \lambda x:[c]. \langle \phi(\Gamma \vdash^{l,l'} c : T) \Rightarrow [c] \rangle^{l'} (\langle [c] \Rightarrow \phi(\Gamma \vdash^{l,l'} c : T) \rangle^l x) \\
& \text{where } x \text{ is fresh}
\end{array}$$

$$\begin{array}{ll}
\textbf{Types} & \phi : (\Gamma \vdash^{l,l'} c : T) \rightarrow S \\
& \phi(\Gamma \vdash^{l,l'} \{x:B \mid t\} : B) = \{x:B \mid \phi(\Gamma, x:B \vdash t : \text{Bool})\} \\
& \phi(\Gamma \vdash^{l,l'} x:c_1 \mapsto c_2 : T_1 \rightarrow T_2) = x:\phi(\Gamma \vdash^{l,l'} c_1 : T_1) \rightarrow \phi(\Gamma, x:c_1^{l,l'} \mapsto c_2 : T_2)
\end{array}$$

Fig. 11. The translation ϕ from dependent λ_C to dependent λ_H

the context when checking the codomain, so we need to restore the “variables have raw types” invariant—something we can’t always rely on subtyping to do, since types are not in general subtypes of their raw type. By tracking which variables were bound by contracts in λ_C , we can be sure to cast them to raw types when they’re referenced. We therefore translate the contract above to $f:S \rightarrow \{z:\text{Int} \mid (\langle S \Rightarrow [\text{Int} \rightarrow \text{Int}] \rangle^{l'} f) 0 = 0\}$, where $S = x:\{x:\text{Int} \mid \text{pos } x\} \rightarrow [\text{Int}]$. This (partially) motivates the $x:c^{l,l'}$ binding form in dependent λ_C .

Bulletproofing uses raw types, defined here for the dependent system.

$$\begin{array}{llll}
\llbracket \{x:B \mid s\} \rrbracket & = & \{x:B \mid \text{true}\} & \llbracket x:S_1 \rightarrow S_2 \rrbracket & = & \llbracket S_1 \rrbracket \rightarrow \llbracket S_2 \rrbracket \\
\llbracket B \rrbracket & = & \{x:B \mid \text{true}\} & \llbracket T_1 \rightarrow T_2 \rrbracket & = & \llbracket T_1 \rrbracket \rightarrow \llbracket T_2 \rrbracket \\
\llbracket \{x:B \mid t\} \rrbracket & = & \{x:B \mid \text{true}\} & \llbracket x:c_1 \mapsto c_2 \rrbracket & = & \llbracket c_1 \rrbracket \rightarrow \llbracket c_2 \rrbracket
\end{array}$$

Note that dependency is eliminated.

We could write the translation on terms instead of derivations, defining

$$\phi(x:c_1 \mapsto c_2) = x:\phi(c_1) \rightarrow \phi(c_2) \{x := \langle \phi(c_1) \Rightarrow [c_1] \rangle^{l'} x\}$$

but the proofs are easier if we translate derivations.

Constants translate to themselves. One technical point: to maintain the raw type invariant, we need λ_H ’s higher-order constants to have typings that can be seen as raw by the subtyping relation, i.e., $\Delta \vdash \text{ty}_h(k) <: [\text{ty}_c(k)]$. This can be proven at base types (since we have already assumed that $\text{ty}_h(k)$ is the “most specific type” for each k), but must be assumed for higher-order types. This slightly restricts the types we might assign to our constants, e.g., we cannot say $\text{ty}_h(\text{sqrt}) = x:\{x:\text{Float} \mid x \geq 0\} \rightarrow \{y:\text{Float} \mid (y * y) = x\}$, since it is not the case that $\Delta \vdash \text{ty}_h(\text{sqrt}) <: [\text{Float} \rightarrow \text{Float}]$. Since its domain cannot be refined, $\llbracket \text{sqrt} \rrbracket$ must be defined for all $k \in \mathcal{K}_{\text{Float}}$, e.g., $\llbracket \text{sqrt} \rrbracket(-1)$ must be defined. We’ve already required that denotations be total over their simple types in λ_C , and λ_H

$$\begin{array}{c}
\textbf{Contract/type correspondence} \\
\boxed{c \sim^{l,l'} S : T} \\
\{x:B \mid t\} \sim^{l,l'} \{x:B \mid s\} : B \iff \forall k \in \mathcal{K}_B. t\{x := k\} \sim s\{x := k\} : \text{Bool} \\
x:c_1 \mapsto c_2 \sim^{l,l'} x:S_1 \rightarrow S_2 : T_1 \rightarrow T_2 \iff c_1 \sim^{l',l} S_1 : T_1 \wedge \\
\forall t \sim s : T_1. c_2\{x := \langle c_1 \rangle^{l',l} t\} \sim^{l,l'} S_2\{x := \langle [S_1] \Rightarrow S_1 \rangle^{l'} s\} : T_2 \\
\textbf{Dual closing substitutions} \\
\Gamma \models \delta \iff \begin{cases} \forall x:T \in \Gamma. \delta_1(x) \sim \delta_2(x) : T \\ \forall x:c^{l,l'} \in \Gamma. \delta_1(x) = \langle \delta_1(c) \rangle^{l,l'} t \wedge \delta_2(x) = \langle [c] \Rightarrow \delta_2(S) \rangle^l s \\ \text{where } S = \phi(\Gamma \vdash^{l,l'} c : [c]) \wedge t \sim s : [c] \end{cases} \\
\textbf{Lifted to open terms} \\
\Gamma \vdash t \sim s : T \iff \forall \delta. (\Gamma \models \delta \text{ implies } \delta_1(t) \sim \delta_2(s) : T) \\
\Gamma \vdash c \sim^{l,l'} S : T \iff \forall \delta. (\Gamma \models \delta \text{ implies } \delta_1(c) \sim^{l,l'} \delta_2(S) : T)
\end{array}$$

Fig. 12. Blame-exact correspondence for ϕ from picky λ_C

uses the same denotation function $\llbracket - \rrbracket$, so this requirement does not seem too severe. In any case, we can define it to be equal to $\uparrow l_0$, for some l_0 . We could instead translate k to $\langle \text{ty}_h(k) \Rightarrow [\text{ty}_h(k)] \rangle^{l_0} k$; however, in this case the nondependent fragments of the languages would no longer correspond exactly.

We extend the term correspondence of Figure 10 to contracts and types, lifting the correspondences to open terms using dual closing substitutions. For a binding $x:c^{l,l'} \in \Gamma$, we use ϕ to insert the negative cast (labelled with l') and closing substitutions (in Figure 12) to insert the positive cast (labelled with l). Do not be confused by the label used for function contract correspondence—this definition does, in fact, match up with closing substitutions. A binding $x:c^{l,l'} \in \Gamma$ must have come from the domain of an application of T_FUNC, so the labels on the binding are *already* swapped when ϕ or $\Gamma \models \delta$ sees them. In the definition of function contract correspondence, we swap manually—whence the l' on the inserted cast. It helps to think of polarity in terms of position rather than the presence or absence of a prime.

5.1 Lemma [Expansion and contraction]: If $t \longrightarrow_c^* t'$, and $s \longrightarrow_h^* s'$ then $t \sim s : T$ iff $t' \sim s' : T$.

5.2 Lemma [Constants correspond to themselves]: For all $k, k \sim k : \text{ty}_c(k)$.

5.3 Lemma [Equivalence is closed under parallel reduction]: If $s \Rightarrow s'$ then $t \sim s : T$ iff $t \sim s' : T$. Similarly, if $S \Rightarrow S'$ then $c \sim^{l,l'} S : T$ iff $c \sim^{l,l'} S' : T$.

Proof

In both cases, by induction on T , using the first to prove the second. \square

5.4 Lemma [Trivial casts]: If $t \sim s : B$, then $t \sim \langle S \Rightarrow [B] \rangle^l s : B$ for all S .

5.5 Lemma [Related base casts]: If $\{x:B \mid t\} \sim^{l_0,l_1} \{x:B \mid s\} : B$ and $t' \sim s' : B$, then $\langle \{x:B \mid t\} \rangle^{l,l'} t' \sim \langle S \Rightarrow \{x:B \mid s\} \rangle^l s' : B$ for all S .

Proof

Direct. Note that l_0 and l_1 are entirely irrelevant. \square

5.6 Lemma [Bulletproofing]: If $t \sim s : T$ and $c \sim^{l,l'} S : T$ then $\langle c \rangle^{l,l'} t \sim \langle S \Rightarrow \lceil S \rceil \rangle^{l'} \langle \lceil S \rceil \Rightarrow S \rangle^l s : T$.

Proof

By induction on T . First, observe that either both t and s go to $\uparrow l''$ or both t and s go to values related at T . In the former case, the outer terms also go to blame. So we only consider the case where $t \rightarrow_c^* v$, $s \rightarrow_h^* w$, and $v \approx w : T$.

$T = B$: So $c = \{x:B \mid t_1\}$ and $S = \{x:B \mid s_1\}$ and $S' = \{x:B \mid s_2\}$. By Lemma 5.5 we have $\langle c \rangle^{l,l'} t \sim \langle \lceil S \rceil \Rightarrow S \rangle^l s : B$. By Lemma 5.4 we can add the extra, trivial cast $\langle S \Rightarrow \lceil S \rceil \rangle^{l'}$.

$T = T_1 \rightarrow T_2$: We know that $c = x:c_1 \mapsto c_2$ and $S = x:S_1 \rightarrow S_2$. Let $t' \sim s' : T_1$. We only need to consider the case where $t' \rightarrow_c^* v'$ and $s' \rightarrow_h^* w'$ —if $t' \rightarrow_c^* \uparrow l''$ and $s' \rightarrow_h^* \uparrow l''$ the outer terms correspond because both blame l'' .

On the λ_C side, $(\langle c \rangle^{l,l'} t) t' \rightarrow_c^* \langle c_2 \{x := \langle c_1 \rangle^{l',l} v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l',l} v'))$. In λ_H , we can see

$$\begin{aligned} & (\langle S \Rightarrow \lceil S \rceil \rangle^{l'} \langle \lceil S \rceil \Rightarrow S \rangle^l s) s' \rightarrow_h^* \\ & \langle S_2 \{x := \langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} w'\} \Rightarrow \lceil S_2 \rceil \rangle^{l'} ((\langle \lceil S \rceil \Rightarrow S \rangle^l w) (\langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} w')) \end{aligned}$$

We cannot determine where the redex is until we know the shape of T_1 —does the negative argument cast step to an active check, or do we decompose the positive cast?

— $T_1 = B$. Since $v' \approx w' : B$, we must have $v' = w' = k \in \mathcal{K}_B$. By Lemma 5.5 and $c_1 \sim^{l',l} S_1 : B$, we know that $\langle c_1 \rangle^{l',l} v' \sim \langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} w' : B$. Both terms go to blame or to the same value—which must be k , from inspection of the contract and cast evaluation rules. The former case is immediate, since the outer terms then go to blame. So suppose $\langle c_1 \rangle^{l',l} k \rightarrow_c^* k$ and $\langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k \rightarrow_h^* k$. Now the terms evaluate like so:

$$\begin{aligned} & \langle c_2 \{x := \langle c_1 \rangle^{l',l} v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l',l} v')) \rightarrow_c^* \langle c_2 \{x := \langle c_1 \rangle^{l',l} k\} \rangle^{l,l'} (v k) \\ & \langle S_2 \{x := \langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k\} \Rightarrow \lceil S_2 \rceil \rangle^{l'} ((\langle \lceil S \rceil \Rightarrow S \rangle^l w) (\langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k)) \rightarrow_h^* \\ & \langle S_2 \{x := \langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k\} \Rightarrow \lceil S_2 \rceil \rangle^{l'} \\ & \langle \lceil S_2 \rceil \Rightarrow S_2 \{x := k\} \rangle^l (w (\langle S_1 \rceil \Rightarrow \lceil S_1 \rceil \rangle^{l'} k)) \end{aligned}$$

By Lemma 5.4, $k \sim \langle S_1 \rceil \Rightarrow \lceil S_1 \rceil \rangle^{l'} k : B$, so $v k \sim w (\langle S_1 \rceil \Rightarrow \lceil S_1 \rceil \rangle^{l'} k) : T_2$.

We have by definition (and $k \sim k : B$) that $c_2 \{x := \langle c_1 \rangle^{l',l} k\} \sim^{l,l'} S_2 \{x := \langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k\} : T_2$. Recall that $\langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k \rightarrow_h^* k$. This implies $\langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k \Rightarrow^* k$ (Lemma A4 in the Coq). We can then see that $S_2 \{x := \langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k\} \Rightarrow^* S_2 \{x := k\}$ by Lemma A1 in the Coq. By extension with the congruence rules:

$$\begin{aligned} & \langle S_2 \{x := \langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k\} \Rightarrow \lceil S_2 \rceil \rangle^{l'} \\ & \langle \lceil S_2 \rceil \Rightarrow S_2 \{x := \langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k\} \rangle^l (w (\langle S_1 \rceil \Rightarrow \lceil S_1 \rceil \rangle^{l'} k)) \Rightarrow \\ & \langle S_2 \{x := \langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} k\} \Rightarrow \lceil S_2 \rceil \rangle^{l'} \\ & \langle \lceil S_2 \rceil \Rightarrow S_2 \{x := k\} \rangle^l (w (\langle S_1 \rceil \Rightarrow \lceil S_1 \rceil \rangle^{l'} k)) \end{aligned}$$

By the IH $\langle c_2\{x := \langle c_1\}^{l',l} k\} \rangle^{l,l'} (v\ k)$ corresponds to the former, which means it is related to the latter by Lemma 5.3. We conclude the case with expansion (Lemma 6.1).

— $T_1 = T_{11} \rightarrow T_{12}$. We continue with an application of F_CDECOMP in λ_H :

$$\begin{aligned} & \langle S_2\{x := \langle [S_1] \Rightarrow S_1 \rangle^{l'} w'\} \Rightarrow [S_2] \rangle^{l'} \\ & ((\langle [S] \Rightarrow S \rangle^l w) (\langle [S_1] \Rightarrow S_1 \rangle^{l'} w')) \longrightarrow_h^* \\ & \langle S_2\{x := \langle [S_1] \Rightarrow S_1 \rangle^{l'} w'\} \Rightarrow [S_2] \rangle^{l'} \\ & \langle [S_2] \Rightarrow S_2\{x := \langle [S_1] \Rightarrow S_1 \rangle^{l'} w'\} \rangle^l \\ & (w (\langle S_1 \Rightarrow [S_1] \rangle^l \langle [S_1] \Rightarrow S_1 \rangle^{l'} w')) \end{aligned}$$

By the IH on $c_1 \sim^{l',l} S_1 : T_1$ and $v' \sim w' : T_1$, we can find what we need for the domain: $\langle c_1 \rangle^{l',l} v' \sim \langle S_1 \Rightarrow [S_1] \rangle^l \langle [S_1] \Rightarrow S_1 \rangle^{l'} w' : T_1$. By assumption, the results of applying v and w to these values correspond. (And they *are* values, since function contracts/casts applied to values are values.)

We have $c_2\{x := \langle c_1 \rangle^{l',l} v'\} \sim^{l,l'} S_2\{x := \langle [S_1] \Rightarrow S_1 \rangle^{l'} w'\} : T_2$ by assumption, so the IH tells us that the codomain contract and bulletproofing correspond. We conclude by expansion (Lemma 6.1). \square

Having characterized how contracts and pairs of related casts relate, we show that terms correspond to their translation.

5.7 Theorem [Behavioral correspondence]: If $\vdash \Gamma$, then:

1. If $\phi(\Gamma \vdash t : T) = s$ then $\Gamma \vdash t \sim s : T$.
2. If $\phi(\Gamma \vdash^{l,l'} c : T) = S$ then $\Gamma \vdash c \sim^{l,l'} S : T$.

Proof

We simultaneously show both properties by induction on the depth of ϕ 's recursion. \square

We can now prove that ϕ preserves types, using Theorem 5.7 to show that ϕ preserves the implication judgment. As a preliminary, we use the behavioral correspondence to show that ϕ preserves the implication judgment.

5.8 Lemma: If $t_1 \longrightarrow_c^* \text{true}$ implies $t_2 \longrightarrow_c^* \text{true}$ then $\emptyset \vdash \phi(\emptyset \vdash t_1 : \text{Bool}) \supset \phi(\emptyset \vdash t_1 : \text{Bool})$.

Proof

By the logical relation. \square

The type preservation proof is very similar to the correspondence proof of Theorem 5.7.

5.9 Theorem [Type preservation]: If $\phi(\vdash \Gamma) = \Delta$, then:

1. $\vdash \Delta$.
2. If $\phi(\Gamma \vdash t : T) = s$ then $\Delta \vdash s : [T]$.
3. If $\phi(\Gamma \vdash^{l,l'} c : T) = S$ then $\Delta \vdash S$.

Proof

We prove all three properties simultaneously, by induction on the depth of ϕ 's recursion.

The proof is by cases on the λ_C context well-formedness/term typing/contract well-formedness derivations, which determine the branch of ϕ taken. \square

Term translation

$$\boxed{\psi : s \rightarrow t}$$

$$\begin{aligned} \psi(x) &= x & \psi(k) &= k \\ \psi(\lambda x:S. s) &= \lambda x:[S]. \psi(s) & \psi(s_1 s_2) &= \psi(s_1) \psi(s_2) \\ \psi(\langle S_1 \Rightarrow S_2 \rangle^l) &= \langle \psi^l(S_1, S_2) \rangle^{l,l} & \psi(\uparrow l) &= \uparrow l \\ \psi(\langle \{x:B \mid s_1\}, s_2, k \rangle^l) &= \langle \{x:B \mid \psi(s_1)\}, \psi(s_2), k \rangle^l \end{aligned}$$

Cast translation

$$\boxed{\psi : S \times S \times l \rightarrow T}$$

$$\begin{aligned} \psi^l(\{x:B \mid s_1\}, \{x:B \mid s_2\}) &= \{x:B \mid \psi(s_2)\} \\ \psi^l(x:S_{11} \rightarrow S_{12}, x:S_{21} \rightarrow S_{22}) &= x:\psi^l(S_{21}, S_{11}) \mapsto \psi^l(S_{12} \{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}, S_{22}) \end{aligned}$$

Fig. 13. ψ mapping dependent λ_H to dependent λ_C **5.2 Translating λ_H to lax λ_C : dependent ψ**

In this section, we formally define ψ for the dependent versions of lax λ_C and λ_H . We prove that ψ is type preserving and induces behavioral correspondence.

The full definition of ψ is in Figure 13. Most terms are translated homomorphically. In abstractions, the annotation is translated by erasing the refined λ_H type to its skeleton. As we mentioned in Section 3, the trickiest part is the translation of casts between function types: when generating the codomain contract from a cast between two function types, we perform the same asymmetric substitution as F_CDECOMP. Since ψ inserts new casts, we need to pick a blame label: $\psi(\langle S_1 \Rightarrow S_2 \rangle^l)$ passes l as an index to $\psi^l(S_1, S_2)$.

We reuse the term correspondence $t \sim s : T$ (Figure 10) and define a new contract/cast correspondence $c \sim S_1 \Rightarrow^l S_2 : T$ (Figure 14), relating contracts and pairs of λ_H types—effectively, casts. This correspondence uses the term correspondence in the base type case and follows the pattern of F_CDECOMP in the function case. Since it inserts a cast in the function case, we index the relation with a label, just like ψ . Note that the correspondence is blame-exact, relating λ_C and λ_H terms that either blame the same label or go to corresponding values. We define closing substitutions ignoring the contracts in the context; we lift the relation to open terms in the standard way.

We begin with some standard properties of the term correspondence relation.

5.10 Lemma [Expansion and contraction]: If $t \rightarrow_c^* t'$, and $s \rightarrow_h^* s'$ then $t \sim s : T$ iff $t' \sim s' : T$.

5.11 Lemma [Blame corresponds to blame]: For all T , $\uparrow l \sim \uparrow l : T$.

5.12 Lemma [Constants correspond to themselves]: For all k , $k \approx k : \text{ty}_c(k)$.

As a corollary of Lemma 5.11 and Lemma 5.10, if two terms evaluate to blame, then they correspond. This will be used extensively in the proofs below, as it allows us to eliminate many cases.

5.13 Lemma [Corresponding terms coevaluate]: If $t \sim s : T$ then $t \rightarrow_c^* v \wedge s \rightarrow_h^* w$ or $t \rightarrow_c^* \uparrow l \wedge s \rightarrow_h^* \uparrow l$; moreover, $t \rightarrow_c^* r$ and $s \rightarrow_h^* q$ such that $r \approx q : T$.

5.14 Lemma [Contract/cast correspondence]: If $c \sim S_1 \Rightarrow^l S_2 : T$ and $t \sim s : T$ then $\langle c \rangle^{l,l} t \sim \langle S_1 \Rightarrow S_2 \rangle^l s : T$.

Contract/type correspondence

$$\boxed{c \sim S_1 \Rightarrow^l S_2 : T}$$

$$\{x:B \mid t\} \sim \{x:B \mid s_1\} \Rightarrow^l \{x:B \mid s_2\} : B \iff \forall k \in \mathcal{K}_B. t\{x := k\} \sim s_2\{x := k\} : \text{Bool}$$

$$x:c_1 \mapsto c_2 \sim x:S_{11} \rightarrow S_{12} \Rightarrow^l S_{21} \rightarrow S_{22} : T_1 \rightarrow T_2 \iff c_1 \sim S_{21} \Rightarrow^l S_{11} : T_1 \wedge \forall t \sim s : T_1. c_2\{x := t\} \sim S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l s\} \Rightarrow^l S_{22}\{x := s\} : T_2$$

Closing substitutions

$$\Gamma \models \delta \iff \begin{cases} \forall x:T \in \Gamma, & \delta_1(x) \sim \delta_2(x) : T \\ \forall x:c^{l,l'} \in \Gamma, & \delta_1(x) \sim \delta_2(x) : [c] \end{cases}$$

Lifted to open terms

$$\begin{aligned} \Gamma \vdash t \sim s : T &\iff \forall \delta. (\Gamma \models \delta \text{ implies } \delta_1(t) \sim \delta_2(s) : T) \\ \Gamma \vdash c \sim S_1 \Rightarrow^l S_2 : T &\iff \forall \delta. (\Gamma \models \delta \text{ implies } \delta_1(c) \sim \delta_2(S_1) \Rightarrow^l \delta_2(S_2) : T) \end{aligned}$$

Fig. 14. Blame-exact correspondence for ψ into lax λ_C *Proof*

By induction on T . We reason via expansion (Lemma 5.10), showing that the initial terms reduce to corresponding terms.

$T = B$: So $c = \{x:B \mid t_1\}$, $S_1 = \{x:B \mid s_1\}$, and $S_2 = \{x:B \mid s_2\}$. Since $t \sim s : B$, we know that they either both reduce to $k \in \mathcal{K}_B$ or $\uparrow l'$. If the latter is the case, we are done. So suppose $t \xrightarrow{c}_* k$ along with $s \xrightarrow{h}_* k$.

We can step our terms into active checks as follows, then:

$$\begin{aligned} \langle \{x:B \mid t_1\} \rangle^{l,l} t &\xrightarrow{c}_* \langle \{x:B \mid t_1\}, t_1\{x := k\}, k \rangle^l \\ \{x:B \mid s_1\} \Rightarrow \{x:B \mid s_2\}^l s &\xrightarrow{h}_* \langle \{x:B \mid s_2\}, s_2\{x := k\}, k \rangle^l \end{aligned}$$

By inversion of the contract/cast correspondence, we know that $t_1\{x := k\} \sim s_2\{x := k\} : \text{Bool}$, so these terms go to blame or to a Bool together. If they go to $\uparrow l'$, we are done. If they go to false, then both the obligation and the cast will go to $\uparrow l$. Finally, if they both go to true, then both terms will evaluate to k .

$T = T_1 \rightarrow T_2$: $c = x:c_1 \mapsto c_2$, $S_1 = x:S_{11} \rightarrow S_{12}$, and $S_2 = x:S_{21} \rightarrow S_{22}$. We know by inversion of the contract/cast relation that $c_1 \sim S_{21} \Rightarrow^l S_{11} : T_1$ and that for all $t \sim s : T_1$, $c_2\{x := t\} \sim S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l s\} \Rightarrow^l S_{22}\{x := s\} : T_2$. We want to prove that $\langle c \rangle^{l,l} \sim \langle S_1 \Rightarrow S_2 \rangle^l s : T_1 \rightarrow T_2$. First, we can assume $t \xrightarrow{c}_* v$ and $s \xrightarrow{h}_* w$ where $v \sim w : T_1 \rightarrow T_2$ —if not, both cast and contracted terms go to blame and we are done.

We show that the decomposition of the contract and cast terms correspond for all inputs. Let $t' \sim s' : T_1$. Again, we can assume that they reduce to $v' \sim w' : T_1$, or else we are done by blame lifting. On the λ_C side, we have

$$(\langle c \rangle^{l,l} t) t' \xrightarrow{c}_* \langle c_2\{x := v'\} \rangle^{l,l} (v (\langle c_1 \rangle^{l,l} v'))$$

In λ_H , we find

$$(\langle S_1 \Rightarrow S_2 \rangle^l s) s' \xrightarrow{h}_* \langle S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w\} \Rightarrow S_{22}\{x := w'\} \rangle^l (w (\langle S_{21} \Rightarrow S_{11} \rangle^l w'))$$

By the IH, we know that $\langle c_1 \rangle^{l,l} v' \sim \langle S_{21} \Rightarrow S_{11} \rangle^l w' : T_1$. Since $v \sim w : T_1 \rightarrow T_2$, we have $v (\langle c_1 \rangle^{l,l} v') \sim w (\langle S_{21} \Rightarrow S_{11} \rangle^l w') : T_2$. Again by the IH, we can see that $\langle c_2 \{x := v'\} \rangle^{l,l} (v (\langle c_1 \rangle^{l,l} v')) \sim \langle S_{12} \{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w'\} \Rightarrow S_{22} \{x := w'\} \rangle^l w (\langle S_{21} \Rightarrow S_{11} \rangle^l w') : T_2$. \square

We prove three more technical lemmas necessary for the behavioral and type correspondence.

5.15 Lemma [Skeletal equality of subtypes]: If $\Delta \vdash S_1 <: S_2$, then $\lfloor S_1 \rfloor = \lfloor S_2 \rfloor$.

5.16 Lemma: If $\lfloor S_1 \rfloor = \lfloor S_2 \rfloor = T$, then $\lfloor \psi^l(S_1, S_2) \rfloor = T$.

5.17 Lemma: If $\Delta_1 \vdash S_1$ and $\Delta_1 \vdash S_2$, where $\lfloor S_1 \rfloor = \lfloor S_2 \rfloor$ then

1. If $\Delta_1, x:S_1, \Delta_2 \vdash s : S$ then $\Delta_1, x:S_2, \Delta_2 \{x := \langle S_2 \Rightarrow S_1 \rangle^l x\} \vdash s \{x := \langle S_2 \Rightarrow S_1 \rangle^l x\} : S \{x := \langle S_2 \Rightarrow S_1 \rangle^l x\}$, and
2. If $\Delta_1, x:S_1, \Delta_2 \vdash S$ then $\Delta_1, x:S_2, \Delta_2 \{x := \langle S_2 \Rightarrow S_1 \rangle^l x\} \vdash S \{x := \langle S_2 \Rightarrow S_1 \rangle^l x\}$.

We use the correspondence relations to show that s and its translation $\psi(s)$ correspond—i.e., that ψ faithfully translates the λ_H semantics. We must choose the subject of induction carefully, however, to ensure that we can apply the IH in the case for function casts. An induction on the height of the well-formedness derivation is tricky because of the “extra” substitution that ψ does. Instead, we do induction on the depth of ψ ’s recursion, (and also derivation height, for the S_SUB case).

5.18 Theorem [Behavioral correspondence]:

1. If $\Delta \vdash s : S$ then $\lfloor \Delta \rfloor \vdash \psi(s) \sim s : \lfloor S \rfloor$.
2. If $\Delta \vdash S_1$ and $\Delta \vdash S_2$, where $\lfloor S_1 \rfloor = \lfloor S_2 \rfloor = \lfloor S \rfloor$, then $\lfloor \Delta \rfloor \vdash \psi^l(S_1, S_2) \sim S_1 \Rightarrow^l S_2 : \lfloor S \rfloor$ (for all l).

Proof

By induction on the lexicographically ordered pairs (m, n) , where m is the depth of the recursion of the translation $\psi(s)$ (for part 1) or $\psi^l(S_1, S_2)$ (for part 2) and n is either $\lfloor \Delta \vdash s : S \rfloor$ (for part 1) or $\lfloor \Delta \vdash S_1 \rfloor + \lfloor \Delta \vdash S_2 \rfloor$ (for part 2). The first component decreases in all uses of the IH except for the S_SUB case, where only the second component decreases. Part (1) of the proof proceeds by case analysis on the final rule used in the typing derivation $\Delta \vdash s : S$. Which rule was used determines the shape of $\psi(s)$ in all cases but S_SUB.

We give only the most interesting cases for the first part: S_CAST, S_CHECKING, and S_SUB.

S_CAST: $\Delta \vdash \langle S_1 \Rightarrow S_2 \rangle^l : S_1 \rightarrow S_2$ and $\psi(\langle S_1 \Rightarrow S_2 \rangle^l) = \langle \psi^l(S_1, S_2) \rangle^{l,l}$. By inversion, $\Delta \vdash S_1$ and $\Delta \vdash S_2$, where $\lfloor S_1 \rfloor = \lfloor S_2 \rfloor$.

By the IH for proposition (2), $\lfloor \Delta \rfloor \vdash \psi^l(S_1, S_2) \sim S_1 \Rightarrow^l S_2 : \lfloor S_2 \rfloor$.

Let $\lfloor \Delta \rfloor \models \delta$; we must show $\delta_1(\langle \psi^l(S_1, S_2) \rangle^{l,l}) \sim \delta_2(\langle S_1 \Rightarrow S_2 \rangle^l) : \lfloor S_1 \rfloor \rightarrow \lfloor S_2 \rfloor$. Let $t \sim s : \lfloor S_1 \rfloor$. We have $\delta_1(\langle \psi^l(S_1, S_2) \rangle^{l,l}) t \sim \delta_2(\langle S_1 \Rightarrow S_2 \rangle^l) s : \lfloor S_2 \rfloor$ by Lemma 5.14.

S_CHECKING: We have $\Delta \vdash \langle \{x:B \mid s_1\}, s_2, k \rangle^l : \{x:B \mid s_1\}$; translating yields $\psi(\langle \{x:B \mid s_1\}, s_2, k \rangle^l) = \langle \{x:B \mid \psi(s_1)\}, \psi(s_2), k \rangle^l$. Recall that the terms of the active check are closed. By inversion we have $\emptyset \vdash s_2 : \{x:\text{Bool} \mid \text{true}\}$ and $\emptyset \vdash k : \{x:B \mid \text{true}\}$, so $k \in \mathcal{K}_B$.

By the IH, $\psi(s_2) \sim s_2 : \text{Bool}$. These two terms coevaluate to blame or a boolean constant. There are three cases, all of which result in the active checks evaluating to \approx -corresponding values:

- If they go to $\uparrow l'$, then the checks do too, and $\uparrow l' \approx \uparrow l' : B$.
- If they go to false, then the checks go to $\uparrow l$, and $\uparrow l \approx \uparrow l : B$.
- If they go to true, then the checks go to $k \in \mathcal{K}_B$, and $k \approx k : B$.

S.SUB: $\Delta \vdash s : S$; we do not know anything about the shape of $\psi(s)$. By inversion, $\Delta \vdash s : S'$ and $\Delta \vdash S' <: S$. By Lemma 5.15, $\lfloor S' \rfloor = \lfloor S \rfloor$.

Since the sub-derivation $\Delta \vdash s : S'$ is smaller, by the IH $\lfloor \Delta \rfloor \vdash \psi(s) \sim s : \lfloor S' \rfloor$. But $\lfloor S' \rfloor = \lfloor S \rfloor$, so we're done.

Part (2) of this proof proceeds by cases on $\psi^l(S_1, S_2) = c$.

$\psi^l(S_1, \{x:B \mid s_2\}) = \{x:B \mid \psi(s_2)\}$: Note that $S_2 = \{x:B \mid s_2\}$. By inversion of $\Delta \vdash \{x:B \mid s_2\}$, we have $\Delta, x:\{x:B \mid \text{true}\} \vdash s_2 : \{x:\text{Bool} \mid \text{true}\}$.

By the IH for proposition (1), $\lfloor \Delta \rfloor, x:B \vdash \psi(s_2) \sim s_2 : \text{Bool}$.

We must show $\lfloor \Delta \rfloor \vdash \{x:B \mid \psi(s_2)\} \sim S_1 \Rightarrow^l \{x:B \mid s_2\} : B$. Let $\lfloor \Delta \rfloor \models \delta$; we prove that $\delta_1(\{x:B \mid \psi(s_2)\}) \sim \delta_2(S_1) \Rightarrow^l \delta_2(\{x:B \mid s_2\}) : B$, i.e., for all $k \in \mathcal{K}_B$, that $\delta_1(\psi(s_2))\{x := k\} \sim \delta_2(s_2)\{x := k\} : \text{Bool}$. Since $k \sim k : B$, we can see this last by the IH.

$\psi^l(x:S_{11} \rightarrow S_{12}, x:S_{21} \rightarrow S_{22}) = x:\psi^l(S_{21}, S_{11}) \mapsto \psi^l(S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}, S_{22})$:

We can see $S_2 = x:S_{21} \rightarrow S_{22}$ and so $S_1 = x:S_{11} \rightarrow S_{12}$, where $\lfloor S_{21} \rfloor = \lfloor S_{11} \rfloor$ and $\lfloor S_{22} \rfloor = \lfloor S_{12} \rfloor$. By inversion, we have the following well-formedness derivations:

$$\begin{array}{cc} \Delta \vdash S_{21} & \Delta \vdash S_{11} \\ \Delta, x:S_{21} \vdash S_{22} & \Delta, x:S_{22} \vdash S_{12} \end{array}$$

We can apply the IH (contravariantly) to see

$$\lfloor \Delta \rfloor \vdash \psi^l(S_{21}, S_{11}) \sim S_{21} \Rightarrow^l S_{11} : \lfloor S_{11} \rfloor \quad (*)$$

By weakening (Lemma 4.22), we can see $\Delta, x:S_{21} \vdash S_{21}$ and $\Delta, x:S_{21} \vdash S_{11}$. We can reapply the IH to show $\lfloor \Delta \rfloor, x:\lfloor S_{21} \rfloor \vdash \psi^l(S_{21}, S_{11}) \sim S_{21} \Rightarrow^l S_{11} : \lfloor S_{11} \rfloor$. Now $\Delta, x:S_{21} \vdash \langle S_{21} \Rightarrow S_{11} \rangle^l : S_{21} \rightarrow S_{11}$ and $\Delta, x:S_{21} \vdash \langle S_{21} \Rightarrow S_{11} \rangle^l x : S_{11}$. By Lemma 5.17, we can substitute this last into $\Delta, x:S_{11} \vdash S_{12}$, finding $\Delta, x:S_{21} \vdash S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}$.

We apply the IH for proposition (2) on $\Delta, x:S_{21} \vdash S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}$ and $\Delta, x:S_{21} \vdash S_{22}$, showing

$$\begin{array}{c} \lfloor \Delta \rfloor, x:\lfloor S_{21} \rfloor \vdash \psi^l(S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}, S_{22}) \sim \\ S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\} \Rightarrow^l S_{22} : \lfloor S_{22} \rfloor \end{array} \quad (**)$$

We now combine (*) and (**) to show $\lfloor \Delta \rfloor \vdash \psi^l(x:S_{11} \rightarrow S_{12}, x:S_{21} \rightarrow S_{22}) \sim x:S_{11} \rightarrow S_{12} \Rightarrow^l x:S_{21} \rightarrow S_{22} : \lfloor S_2 \rfloor$. Let $\lfloor \Delta \rfloor \models \delta$. We can apply (*) to see $\delta_1(\psi^l(S_{21}, S_{11})) \sim \delta_2(S_{21}) \Rightarrow^l \delta_2(S_{11}) : \lfloor S_{11} \rfloor$. For the codomain we must show, for all $t \sim s : \lfloor S_{11} \rfloor$, that

$$\begin{array}{c} \delta_1(\psi^l(S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}, S_{22}))\{x := t\} \sim \\ \delta_2(S_{12})\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l s\} \Rightarrow^l \delta_2(S_{22})\{x := s\} : \lfloor S_{22} \rfloor \end{array}$$

Let $t \sim s : \lfloor S_{11} \rfloor$. Recalling that $\lfloor S_{11} \rfloor = \lfloor S_{21} \rfloor$, observe $\lfloor \Delta \rfloor, x:\lfloor S_{21} \rfloor \models \delta\{x := t, s\}$. Call this δ' . By (**) we see

$$\delta'_1(\psi^l(S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}, S_{22})) \sim \delta'_2(S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}) \Rightarrow^l \delta'_2(S_{22}) : \lfloor S_{22} \rfloor$$

which we can rewrite to

$$\begin{array}{c} \delta_1(\psi^l(S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}, S_{22}))\{x := t\} \sim \\ \delta_2(S_{12})\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l s\} \Rightarrow^l \delta_2(S_{22})\{x := s\} : \lfloor S_{22} \rfloor \end{array}$$

This is exactly what we needed to finish the proof of correspondence. \square

As a preliminary to type-preservation, we use behavioral correspondence to show that the implication judgment is preserved.

5.19 Lemma: If $\emptyset \vdash s_1 : \{x:\text{Bool} \mid \text{true}\}$ and $\emptyset \vdash s_2 : \{x:\text{Bool} \mid \text{true}\}$ and $\emptyset \vdash s_1 \supset s_2$, then $\psi(s_1) \longrightarrow_c^* \text{true}$ implies $\psi(s_2) \longrightarrow_c^* \text{true}$.

Proof

By the logical relation. \square

The type preservation proof is very similar to the correspondence proof of Theorem 5.18, though the function case of the type/contract correspondence is intricate.

5.20 Theorem [Type preservation for ψ]:

1. If $\Delta \vdash s : S$ then $\lfloor \Delta \rfloor \vdash \psi(s) : \lfloor S \rfloor$.
2. If $\Delta \vdash S_1, \Delta \vdash S_2$, where $\lfloor S_1 \rfloor = \lfloor S_2 \rfloor = T$, then $\lfloor \Delta \rfloor \vdash^{l,l'} \psi^l(S_1, S_2) : T$.

Proof

By induction on the lexicographically ordered pair containing (a) the depth of the recursion of the translation ψ or $\psi(s)$, and (b) $|\Delta \vdash s : S|$ or $|\Delta \vdash S_1| + |\Delta \vdash S_2|$.

Part (1) of the proof proceeds by case analysis on the final rule of $\Delta \vdash s : S$, which determines the shape of $\psi(s) = t$ in all cases but S_SUB. Part (2) of the proof proceeds by case analysis on $\psi^l(S_1, S_2) = c$.

$\psi^l(x:S_{11} \rightarrow S_{12}, x:S_{21} \rightarrow S_{22}) = x:\psi^l(S_{21}, S_{11}) \mapsto \psi^l(S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}, S_{22})$:

We must have $S_2 = x:S_{21} \rightarrow S_{22}$ and $S_1 = x:S_{11} \rightarrow S_{12}$, where $\lfloor S_{21} \rfloor = \lfloor S_{11} \rfloor$ and $\lfloor S_{22} \rfloor = \lfloor S_{12} \rfloor$. By inversion, we have the following well-formedness derivations:

$$\begin{array}{cc} \Delta \vdash S_{21} & \Delta \vdash S_{11} \\ \Delta, x:S_{21} \vdash S_{22} & \Delta, x:S_{22} \vdash S_{12} \end{array}$$

By the IH $\lfloor \Delta \rfloor \vdash^{l,l'} \psi^l(S_{21}, S_{11}) : \lfloor S_{11} \rfloor$. Note that $\lfloor \psi^l(S_{21}, S_{11}) \rfloor = \lfloor S_{21} \rfloor$.

By weakening, we can see $\Delta, x:S_{21} \vdash S_{21}$ and $\Delta, x:S_{21} \vdash S_{11}$. We can reapply the IH to show $\lfloor \Delta \rfloor, x:\lfloor S_{21} \rfloor \vdash^{l,l'} \psi^l(S_{21}, S_{11}) : \lfloor S_{11} \rfloor$. Now $\Delta, x:S_{21} \vdash \langle S_{21} \Rightarrow S_{11} \rangle^l : S_{21} \rightarrow S_{11}$. Next $\Delta, x:S_{21} \vdash \langle S_{21} \Rightarrow S_{11} \rangle^l x : S_{11}$. By Lemma 5.17, we can substitute this last into $\Delta, x:S_{11} \vdash S_{12}$, finding $\Delta, x:S_{21} \vdash S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}$.

By the IH for proposition (2) on $\Delta, x:S_{21} \vdash S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}$ and $\Delta, x:S_{21} \vdash S_{22}$,

$$\lfloor \Delta \rfloor, x:\lfloor S_{21} \rfloor \vdash^{l,l'} \psi^l(S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}, S_{22}) : \lfloor S_{22} \rfloor$$

By Lemma 5.16, $\lfloor \psi^l(S_{21}, S_{11}) \rfloor = \lfloor S_{21} \rfloor$, so we can rewrite the above derivation to

$$\lfloor \Delta \rfloor, x:\psi^l(S_{21}, S_{11})^{l,l'} \vdash^{l,l'} \psi^l(S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}, S_{22}) : \lfloor S_{22} \rfloor$$

Now by T_FUNC

$$\lfloor \Delta \rfloor \vdash^{l,l'} x:\psi^l(S_{21}, S_{11}) \mapsto \psi^l(S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l x\}, S_{22}) : \lfloor S_{21} \rfloor \rightarrow \lfloor S_{22} \rfloor \quad \square$$

$$\begin{array}{c}
\textbf{Value correspondence} \\
\boxed{v \approx_{\sim} w : T} \\
k \approx_{\sim} k : B \iff k \in \mathcal{K}_B \\
v \approx_{\sim} w : T_1 \rightarrow T_2 \iff \forall t \sim_{\sim} s : T_1. v t \sim_{\sim} w s : T_2 \\
\textbf{Term correspondence} \\
\boxed{t \sim_{\sim} s : T} \\
t \sim_{\sim} s : T \iff s \rightarrow_h^* \uparrow l \vee (t \rightarrow_c^* v \wedge s \rightarrow_h^* w \wedge v \approx_{\sim} w : T) \\
\textbf{Contract/type correspondence} \\
\boxed{c \sim_{\sim} S : T} \\
\{x:B \mid t\} \sim_{\sim} \{x:B \mid s\} : B \iff \forall k \in \mathcal{K}_B. t\{x := k\} \sim_{\sim} s\{x := k\} : \text{Bool} \\
x:c_1 \mapsto c_2 \sim_{\sim} x:S_1 \rightarrow S_2 : T_1 \rightarrow T_2 \iff c_1 \sim_{\sim} S_1 : T_1 \wedge \\
\forall t \sim_{\sim} s : T_1. c_2\{x := t\} \sim_{\sim} S_2\{x := s\} : T_2 \\
\textbf{Dual closing substitutions} \\
\Gamma \models_{\sim} \delta \iff \forall x \in \text{dom}(\Gamma). \delta_1(x) \sim_{\sim} \delta_2(x) : [\Gamma(x)] \\
\textbf{Lifted to open terms} \\
\Gamma \vdash t \sim_{\sim} s : T \iff \forall \delta. (\Gamma \models_{\sim} \delta \text{ implies } \delta_1(t) \sim_{\sim} \delta_2(s) : T) \\
\Gamma \vdash c \sim_{\sim} S : T \iff \forall \delta. (\Gamma \models_{\sim} \delta \text{ implies } \delta_1(c) \sim_{\sim} \delta_2(S) : T)
\end{array}$$

Fig. 15. Blame-inexact correspondence for ϕ from lax λ_C

6 Inexact translations

The same translations ϕ and ψ can be used to move right on the axis of blame (Figure 1). However, in this direction the images of these translations blame strictly more than their pre-images. We were able to use the same correspondence for both exact proofs in Section 5, but the following two proofs use custom correspondences: one where lax λ_C terms correspond to λ_H terms (with possibly more blame), and one where λ_H terms correspond to picky λ_C terms (with possibly more blame). In both cases, the λ_C terms will be on the left and the λ_H terms on the right.

6.1 Translating lax λ_C to λ_H

Translating with ϕ from terms in picky λ_C to exactly corresponding terms in λ_H was a relatively straightforward generalization of the nondependent case; things get more interesting when we consider the translation ϕ from lax λ_C to dependent λ_H . We can prove that it preserves types (for terms without active checks), but we can only show a weaker behavioral correspondence: sometimes lax λ_C terms terminate with values when their ϕ -images go to blame. This weaker property is a consequence of bulletproofing, the asymmetrically substituting F_CDECOMP rule, and the extra casts inserted for type preservation (i.e., for T_VARC derivations). This is not a weakness of our proof technique—we give a counterexample, a lax λ_C term $\emptyset \vdash t : T$ such that $t \rightarrow_c^* v$ and $\phi(\emptyset \vdash t : T) \rightarrow_h^* \uparrow l$.

We can show the behavioral correspondence using a blame-inexact logical relation, defined in Figure 15. The behavioral correspondence here, though weaker than before, is still pretty strong: if $t \sim_{\sim} s : B$ (read “ t blames no more than s at type B ”), then either $s \rightarrow_h^* \uparrow l$ or t and s both go to $k \in \mathcal{K}_B$. This correspondence differs slightly in construction from the earlier exact one—we define \sim_{\sim} as a relation on *values*, while \approx is a relation on *results*. Doing so simplifies our inexact treatment of blame—in particular, Lemma 6.2. We again use the term correspondence to relate contracts and λ_H types. We then lift the correspondences to open terms (Figure 15). Closing substitutions map variables to corresponding terms of appropriate type. Note that closing substitutions ignore the contract part of $x:c^{l,l'}$ bindings, treating them as if they were $x:[c]$.

6.1 Lemma [Expansion and contraction]: If $t \rightarrow_c^* t'$, and $s \rightarrow_h^* s'$ then $t \sim_{\sim} s : T$ iff $t' \sim_{\sim} s' : T$.

Note that there are corresponding terms at every type. We can prove a much stronger lemma than we did for \sim in Lemma 5.11, since the correspondence here is much weaker.

6.2 Lemma [Everything is corresponds to blame]: For all t and T , $t \sim_{\sim} \uparrow l' : T$.

6.3 Lemma [Constants correspond to themselves]: For all k , $k \approx_{\sim} k : \text{ty}_c(k)$.

Proof

By induction on $\text{ty}_c(k)$, recalling that constants are first order. \square

As a corollary of Lemma 5.11 and Lemma 6.1, if two terms evaluate to blame—or even just the λ_H side!—then they correspond. This will be used extensively in the proofs below, as it allows us to eliminate many cases.

We prove three lemmas about contracts and casts at base types. The first two characterize contracts and casts at base types. The third shows that correspondence is closed under adding extra casts to the λ_H term, due to the inexactness of our behavioral correspondence.

6.4 Lemma [Trivial casts]: If $t \sim_{\sim} s : B$, then $t \sim_{\sim} \langle S \Rightarrow [B] \rangle^l s : B$ for all S .

6.5 Lemma [Related base casts]: If $\{x:B \mid t\} \sim_{\sim} \{x:B \mid s\} : B$ and $t' \sim_{\sim} s' : B$, then $\langle \{x:B \mid t\} \rangle^{l,l'} t' \sim_{\sim} \langle S \Rightarrow \{x:B \mid s\} \rangle^l s' : B$ for all S .

Since λ_H terms can go to blame more often than corresponding lax λ_C terms, we can add “extra” casts to λ_H terms. We formalize this in the following lemma, which captures the asymmetric treatment of blame by the \sim_{\sim} relation. We use it to show that the cast substituted in the codomain by F_CDECOMP does not affect behavioral correspondence. Note that the statement of the lemma requires that the types of the cast correspond to *some* contracts at the same type T , but we never use the contracts in the proof—they witness the well-formedness of the λ_H types.

6.6 Lemma [Extra casts]: If $t \sim_{\sim} s : T$ and $c_1 \sim_{\sim} S_1 : T$ and $c_2 \sim_{\sim} S_2 : T$, then $t \sim_{\sim} \langle S_1 \Rightarrow S_2 \rangle^l s : T$.

Proof

The proof is by induction on T . Note that we do not use c_1 or c_2 at all in the proof, but instead they are witnesses to the well-formedness of S_1 and S_2 .

$\lfloor S_1 \rfloor = \lfloor S_2 \rfloor = T$. Either $s \rightarrow_h^* \uparrow l'$ or t and s both go to corresponding values at T . If $s \rightarrow_h^* \uparrow l'$, then $\langle S_1 \Rightarrow S_2 \rangle^l s \rightarrow_h^* \uparrow l'$ and $t \sim_{\sim} \uparrow l' : T$ since everything is related to blame (Lemma 6.2).

Therefore, suppose that $t \rightarrow_c^* v$ and $s \rightarrow_h^* w$ and $v \approx_{\sim} w : T$ in each of the following cases of the induction.

$T = B$: So $S_2 = \{x:B \mid s_2\}$, and $c_2 = \{x:B \mid t_2\}$.

So $t \rightarrow_c^* k$ and $s \rightarrow_h^* k$ for $k \in \mathcal{K}_B$. If t and s both go to k , then $\langle S_1 \Rightarrow S_2 \rangle^l s \rightarrow_h^* \langle \{x:B \mid s_2\}, s_2\{x := k\}, k \rangle^l$. By $c_2 \sim_{\sim} S_2 : B$ we see (in particular) $t_2\{x := k\} \sim_{\sim} s_2\{x := k\} : \text{Bool}$. So $s_2\{x := k\}$ either goes to $\uparrow l'$ or $s_2\{x := k\}$ (and, irrelevantly, $t_2\{x := k\}$) go to some $k' \in \mathcal{K}_{\text{Bool}}$. In the former case, $\langle \{x:B \mid s_2\}, s_2\{x := k\}, k \rangle^l \rightarrow_h^* \uparrow l'$ and we are done (by Lemma 6.2). In the latter case, the λ_H term either goes to $\uparrow l'$ (and everything is related to blame) or goes to k —but so does t , and $k \approx_{\sim} k : B$.

$T = T_1 \rightarrow T_2$: We have:

$$\begin{array}{ll} S_1 &= x:S_{11} \rightarrow S_{12} & S_2 &= x:S_{21} \rightarrow S_{22} \\ c_1 &= x:c_{11} \mapsto c_{12} & c_2 &= x:c_{21} \mapsto c_{22} \end{array}$$

We have $t \rightarrow_c^* v$ and $s \rightarrow_h^* w$, where $v \approx_{\sim} w : T_1 \rightarrow T_2$.

Let $t' \sim_{\sim} s' : T_1$; we wish to see that $v \ t' \sim_{\sim} (\langle S_1 \Rightarrow S_2 \rangle^l w) \ s' : T_2$. Either $s' \rightarrow_h^* \uparrow l'$ or both go to values. In the former case the whole cast goes to $\uparrow l'$ we are done by Lemma 6.2, so let $t' \rightarrow_c^* v'$ and $s' \rightarrow_h^* w'$.

Decomposing the cast in λ_H ,

$$\begin{aligned} &(\langle S_1 \Rightarrow S_2 \rangle^l w) \ s' \rightarrow_h^* \\ &\langle S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w'\} \Rightarrow S_{22}\{x := w'\} \rangle^l (w (\langle S_{21} \Rightarrow S_{11} \rangle^l w')) \end{aligned}$$

We have $c_{21} \sim_{\sim} S_{21} : T_1$ and $c_{11} \sim_{\sim} S_{11} : T_1$, so $v' \sim_{\sim} \langle S_{21} \Rightarrow S_{11} \rangle^l w' : T_1$ by the IH. Since $v \approx_{\sim} w : T_1 \rightarrow T_2$, we can see that $v \ v' \sim_{\sim} w (\langle S_{21} \Rightarrow S_{11} \rangle^l w') : T_2$.

Furthermore, we know that for all $t'' \sim_{\sim} s'' : T_1$ that

- $c_{12}\{x := t''\} \sim_{\sim} S_{12}\{x := s''\} : T_2$ and
- $c_{22}\{x := t''\} \sim_{\sim} S_{22}\{x := s''\} : T_2$.

We know that $v' \sim_{\sim} w' : T_1$ and $v' \sim_{\sim} \langle S_{21} \Rightarrow S_{11} \rangle^l w' : T_1$, so we can see

- $c_{12}\{x := v'\} \sim_{\sim} S_{12}\{x := w'\} : T_2$ and
- $c_{22}\{x := v'\} \sim_{\sim} S_{22}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w'\} : T_2$.

So by the IH,

$$v \ v' \sim_{\sim} \langle S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w'\} \Rightarrow S_{22}\{x := w'\} \rangle^l (w (\langle S_{21} \Rightarrow S_{11} \rangle^l w')) : T_2$$

and we are done by expansion (Lemma 6.1). \square

To apply the extra cast lemma, we'll need these “witness” contracts for raw types; to that end we define trivial contracts. These contracts are *lifted* from types, and are the λ_C correlate to λ_H 's raw types.

$$\begin{aligned} B\uparrow &= \{x:B \mid \text{true}\} \\ (T_1 \rightarrow T_2)\uparrow &= (T_1\uparrow) \mapsto (T_2\uparrow) \end{aligned}$$

6.7 Lemma [Lifted types logically relate to raw types]: For all T , $T \uparrow \sim_{\sim} [T] : T$.

The “bulletproofing” lemma is the key to the behavioral correspondence proof. We show that a contract application corresponds to bulletproofing with related types. Note that we allow for different types in the two casts. This is necessary due to an asymmetric substitution that occurs when $T = B \rightarrow T_2$.

6.8 Lemma [Bulletproofing]: If $t \sim_{\sim} s : T$ and $c \sim_{\sim} S : T$ and $c \sim_{\sim} S' : T$, then $\langle c \rangle^{l,l'} t \sim_{\sim} \langle S' \Rightarrow [S'] \rangle^{l'} \langle [S] \Rightarrow S \rangle^l s : T$.

Proof

By induction on T . First, observe that either $s \rightarrow_h^* \uparrow^{l''}$ or both t and s go to values related at T . In the former case, $\langle S' \Rightarrow [S'] \rangle^{l'} \langle [S] \Rightarrow S \rangle^l s \rightarrow_h^* \uparrow^{l''}$, and everything is related to blame (Lemma 6.2). So $t \rightarrow_c^* v$, $s \rightarrow_h^* w$, and $v \approx_{\sim} w : T$.

$T = B$: So $c = \{x:B \mid t_1\}$ and $S = \{x:B \mid s_1\}$ and $S' = \{x:B \mid s_2\}$. By Lemma 6.5 we have $\langle c \rangle^{l,l'} t \sim_{\sim} \langle [S] \Rightarrow S \rangle^l s : B$. By Lemma 6.4 we can add the extra, trivial cast.

$T = T_1 \rightarrow T_2$: We know that $c = x:c_1 \mapsto c_2$, $S = x:S_1 \rightarrow S_2$ and $S' = x:S'_1 \rightarrow S'_2$. Let $t' \sim_{\sim} s' : T_1$. By Lemma 6.2, we only need to consider the case where $t' \rightarrow_c^* v'$ and $s' \rightarrow_h^* w'$ —if $s' \rightarrow_h^* \uparrow^{l''}$ we’re done.

On the λ_C side, $(\langle c \rangle^{l,l'} t) t' \rightarrow_c^* \langle c_2 \{x := v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l',l} v'))$. In λ_H , we can see

$$\begin{aligned} & (\langle S' \Rightarrow [S'] \rangle^{l'} \langle [S] \Rightarrow S \rangle^l s) s' \rightarrow_h^* \\ & \langle S'_2 \{x := \langle [S'_1] \Rightarrow S'_1 \rangle^{l'} w'\} \Rightarrow [S'_2] \rangle^{l'} ((\langle [S] \Rightarrow S \rangle^l w) (\langle [S'_1] \Rightarrow S'_1 \rangle^{l'} w')) \end{aligned}$$

We cannot determine where the redex is until we know the shape of T_1 —does the negative argument cast step to an active check, or do we decompose the positive cast?

— $T_1 = B$.

By Lemma 6.5 and $c_1 \sim_{\sim} S'_1 : B$, we know that $\langle c_1 \rangle^{l',l} v' \sim_{\sim} \langle [S'_1] \Rightarrow S'_1 \rangle^{l'} w' : B$. The λ_H term goes to blame or both terms go to the same value, $v' = w' = k \in \mathcal{K}_B$. In the former case, the entire λ_H term goes to blame and we are done by Lemma 6.2. So suppose $\langle c_1 \rangle^{l',l} k \rightarrow_c^* k$ and $\langle [S'_1] \Rightarrow S'_1 \rangle^{l'} w' \rightarrow_h^* k$. Now the terms evaluate like so:

$$\langle c_2 \{x := v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l',l} v')) \rightarrow_c^* \langle c_2 \{x := k\} \rangle^{l,l'} (v k)$$

$$\begin{aligned} & \langle S'_2 \{x := \langle [S'_1] \Rightarrow S'_1 \rangle^{l'} w'\} \Rightarrow [S'_2] \rangle^{l'} \\ & ((\langle [S] \Rightarrow S \rangle^l w) (\langle [S'_1] \Rightarrow S'_1 \rangle^{l'} w')) \rightarrow_h^* \\ & \langle S'_2 \{x := \langle [S'_1] \Rightarrow S'_1 \rangle^{l'} w'\} \Rightarrow [S'_2] \rangle^{l'} \\ & \langle [S_2] \Rightarrow S_2 \{x := k\} \rangle^l (w (\langle S_1 \Rightarrow [S_1] \rangle^l k)) \end{aligned}$$

By Lemma 6.4, $k \sim_{\sim} \langle S_1 \Rightarrow [S_1] \rangle^l k : B$, so $v k \sim_{\sim} w (\langle S_1 \Rightarrow [S_1] \rangle^l k) : T_2$.

Noting that $k \sim_{\sim} k : B$ and $k \sim_{\sim} \langle [S_1] \Rightarrow S_1 \rangle^l k : B$, we can see that $c_2 \{x := k\} \sim_{\sim} S_2 \{x := k\} : T_2$ and $c_2 \{x := k\} \sim_{\sim} S'_2 \{x := \langle [S_1] \Rightarrow S_1 \rangle^l k\} : T_2$. Now the IH shows that $\langle c_2 \{x := k\} \rangle^{l,l'} (v k) \sim_{\sim} \langle S'_2 \{x := \langle [S'_1] \Rightarrow S'_1 \rangle^{l'} w'\} \Rightarrow [S'_2] \rangle^{l'} \langle [S_2] \Rightarrow S_2 \{x := k\} \rangle^l (w (\langle S_1 \Rightarrow [S_1] \rangle^l k)) : T_2$, and we conclude the case with expansion (Lemma 6.1).

- $T_1 = T_{11} \rightarrow T_{12}$. We can continue with an application of $F_CDECOMP$ in λ_H and find:

$$\begin{aligned} & \langle S'_2 \{x := \langle \lceil S'_1 \rceil \Rightarrow S'_1 \rangle^{l'} w'\} \Rightarrow \lceil S'_2 \rceil \rangle^{l'} \\ & \quad (((\langle \lceil S \rceil \Rightarrow S \rangle^l w) (\langle \lceil S'_1 \rceil \Rightarrow S'_1 \rangle^{l'} w')) \longrightarrow_h^* \\ & \langle S'_2 \{x := \langle \lceil S'_1 \rceil \Rightarrow S'_1 \rangle^{l'} w'\} \Rightarrow \lceil S'_2 \rceil \rangle^{l'} \\ & \quad \langle \lceil S_2 \rceil \Rightarrow S_2 \{x := \langle \lceil S'_1 \rceil \Rightarrow S'_1 \rangle^{l'} w'\} \rangle^l \\ & \quad (w (\langle S_1 \Rightarrow \lceil S_1 \rceil \rangle^l \langle \lceil S'_1 \rceil \Rightarrow S'_1 \rangle^{l'} w')) \end{aligned}$$

By the IH, $\langle c_1 \rangle^{l',l} v' \sim_{\sim} \langle S_1 \Rightarrow \lceil S_1 \rceil \rangle^l \langle \lceil S'_1 \rceil \Rightarrow S'_1 \rangle^{l'} w' : T_1$. By assumption, the results of applying v and w to these values correspond. (And they *are* values, since function contracts/casts applied to values are values.)

We know $c_1 \sim_{\sim} S'_1 : T_1$, and by Lemma 6.7 $T_1 \uparrow \sim_{\sim} \lceil S'_1 \rceil : T_1$. Since $v' \sim_{\sim} w' : T_1$, Lemma 6.6 shows $v' \sim_{\sim} \langle \lceil S'_1 \rceil \Rightarrow S'_1 \rangle^{l'} w' : T_1$. This lets us see that $c_2 \{x := v'\} \sim_{\sim} S'_2 \{x := \langle \lceil S'_1 \rceil \Rightarrow S'_1 \rangle^{l'} w'\} : T_2$ and $c_2 \{x := v'\} \sim_{\sim} S_2 \{x := \langle \lceil S'_1 \rceil \Rightarrow S'_1 \rangle^{l'} w'\} : T_2$. Now the IH and expansion (Lemma 6.1) complete the proof. \square

Having characterized how contracts and pairs of related casts relate, we show that translated terms correspond to their sources.

6.9 Theorem [Behavioral correspondence]: If $\vdash \Gamma$, then:

1. If $\phi(\Gamma \vdash t : T) = s$ then $\Gamma \vdash t \sim_{\sim} s : T$.
2. If $\phi(\Gamma \vdash^{l,l'} c : T) = S$ then $\Gamma \vdash c \sim_{\sim} S : T$.

Proof

We simultaneously show both properties by induction on the depth of ϕ 's recursion. To show $\Gamma \vdash t \sim_{\sim} s : T$, let $\Gamma \models \delta$ —we will show $\delta_1(t) \sim_{\sim} \delta_2(s) : T$.

The proof proceeds by case analysis on the final rule of the translated typing and well-formedness derivations. \square

We find a weak corollary: $\phi(\Gamma \vdash t : B) \longrightarrow_h^* k$ implies $t \longrightarrow_c^* k$: if the λ_H term does *not* go to blame, then the original λ_C term must go to the same constant.

We can also show type preservation for terms not containing active checks. (We don't know that translated active checks are well typed, because Theorem 6.9 isn't strong enough to preserve the implication judgment. We only expect these checks to occur at runtime, so this is good enough: ϕ preserves the types of source programs.)

6.10 Theorem [Type preservation]: For programs without active checks, if $\phi(\vdash \Gamma) = \Delta$, then:

1. $\vdash \Delta$.
2. $\Delta \vdash \phi(\Gamma \vdash t : T) : \lceil T \rceil$.
3. $\Delta \vdash \phi(\Gamma \vdash^{l,l'} c : T)$.

Proof

We prove all three properties simultaneously, by induction on the depth of ϕ 's recursion.

The proof is by cases on the λ_C context well-formedness/term typing/contract well-formedness derivations, which determine the branch of ϕ taken. \square

To see that the ϕ in Figure 11 does not give us exact blame, let us look at two counterexamples; in both cases, a lax λ_C term goes to a value while its translation goes to blame. In the first example, blame is raised in λ_H due to bulletproofing. In the second, blame is raised due to the extra cast from the translation of T_VARC . For the first counterexample, let

$$\begin{aligned} c &= f:(x:\{x:\text{Int} \mid \text{true}\} \mapsto \{y:\text{Int} \mid \text{nonzero } y\}) \mapsto \{z:\text{Int} \mid f \ 0 = 0\} \\ S_1 &= x:\{x:\text{Int} \mid \text{true}\} \rightarrow \{y:\text{Int} \mid \text{nonzero } y\} \\ S &= \phi(\emptyset \vdash^{l,l} c : (\text{Int} \rightarrow \text{Int}) \rightarrow \text{Int}) \\ &= f:S_1 \rightarrow \{z:\text{Int} \mid (\langle S_1 \Rightarrow \lceil S_1 \rceil \rangle^l f) \ 0 = 0\}. \end{aligned}$$

We find $\langle c \rangle^{l,l} (\lambda f.0) (\lambda x.0) \longrightarrow_c^* 0$ but

$$(\lambda x:\lceil c \rceil. \langle S \Rightarrow \lceil S \rceil \rangle^l (\langle \lceil S \rceil \Rightarrow S \rangle^l x)) (\lambda f.0) (\lambda x.0) \longrightarrow_h^* \uparrow l.$$

For the second counterexample, let

$$\begin{aligned} c' &= f:(x:\{x:\text{Int} \mid \text{nonzero } x\} \mapsto \{y:\text{Int} \mid \text{true}\}) \mapsto \{z:\text{Int} \mid f \ 0 = 0\} \\ S'_1 &= x:\{x:\text{Int} \mid \text{nonzero } x\} \rightarrow \{y:\text{Int} \mid \text{true}\} \\ S' &= \phi(\emptyset \vdash^{l,l} c' : (\text{Int} \rightarrow \text{Int}) \rightarrow \text{Int}) \\ &= f:S'_1 \rightarrow \{z:\text{Int} \mid (\langle S'_1 \Rightarrow \lceil S'_1 \rceil \rangle^l f) \ 0 = 0\}. \end{aligned}$$

We find $\langle c' \rangle^{l,l} (\lambda f.0) (\lambda x.0) \longrightarrow_c^* 0$ but

$$(\lambda x:\lceil c' \rceil. \langle S' \Rightarrow \lceil S' \rceil \rangle^l (\langle \lceil S' \rceil \Rightarrow S' \rangle^l x)) (\lambda f.0) (\lambda x.0) \longrightarrow_h^* \uparrow l.$$

The extra casts that ϕ inserts are all necessary—none can be removed. So while variations on this ϕ are possible, they can only add more casts, which won't resolve the problem that λ_H blames *more*.

6.2 Translating λ_H to picky λ_C

Terms in λ_H and their ψ -images in lax λ_C correspond exactly, as shown Section 5.2. When we change the operational semantics of λ_C to be *picky*, however, $\psi(s)$ blames (strictly) more often than s . Nevertheless, we can show an inexact correspondence, as we did for ϕ and lax λ_C in Section 6.1. We use a logical relation \sim_{\prec} (similar to our earlier inexact relation) for ϕ into lax λ_C (Figure 15). Here we've reversed the asymmetry: picky λ_C may blame more than λ_H . The proof follows the same general pattern: we first show that it is safe to add extra contract checks, then we show that contracts and casts correspond (inexactly), then the correspondence for well-typed terms. We can also show type preservation for source programs (excluding active checks).

6.11 Lemma [Expansion and contraction]: If $t \longrightarrow_c^* t'$ and $s \longrightarrow_h^* s'$ then $t \sim_{\prec} s : T$ iff $t' \sim_{\prec} s' : T$.

6.12 Lemma [Blame corresponds to everything]: For all T , $\uparrow l \sim_{\prec} s : T$.

6.13 Lemma [Constants correspond to themselves]: For all k , $k \approx_{\prec} k : \text{ty}_c(k)$.

Proof

By induction on $\text{ty}_c(k)$, recalling that constants are first order. \square

$$\begin{array}{c}
\textbf{Value correspondence} \\
\boxed{v \approx_{\prec} w : T} \\
k \approx_{\prec} k : B \iff k \in \mathcal{K}_B \\
v \approx_{\prec} w : T_1 \rightarrow T_2 \iff \forall t \sim_{\prec} s : T_1. v t \sim_{\prec} w s : T_2 \\
\textbf{Term correspondence} \\
\boxed{t \sim_{\prec} s : T} \\
t \sim_{\prec} s : T \iff t \rightarrow_c^* \uparrow l \vee t \rightarrow_c^* v \wedge s \rightarrow_h^* w \wedge v \approx_{\prec} w : T \\
\textbf{Contract/type correspondence} \\
\boxed{c \sim_{\prec} S_1 \Rightarrow S_2 : T} \\
\{x:B \mid t\} \sim_{\prec} \{x:B \mid s_1\} \Rightarrow \{x:B \mid s_2\} : B \iff \\
\forall k \in \mathcal{K}_B. t\{x := k\} \sim_{\prec} s_2\{x := k\} : \text{Bool} \\
x:c_1 \mapsto c_2 \sim_{\prec} x:S_{11} \rightarrow S_{12} \Rightarrow x:S_{21} \rightarrow S_{22} : T_1 \rightarrow T_2 \iff c_1 \sim_{\prec} S_{21} \Rightarrow S_{11} : T_1 \wedge \\
\forall l. \forall t \sim_{\prec} s : T_1. c_2\{x := t\} \sim_{\prec} S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l s\} \Rightarrow S_{22}\{x := s\} : T_2 \\
\textbf{Dual closing substitutions} \\
\Gamma \models \delta \iff \forall x \in \text{dom}(\Gamma). \delta_1(x) \sim_{\prec} \delta_2(x) : [\Gamma(x)] \\
\textbf{Lifted to open terms} \\
\Gamma \vdash t \sim_{\prec} s : T \iff \forall \delta. (\Gamma \models \delta \text{ implies } \delta_1(t) \sim_{\prec} \delta_2(s) : T) \\
\Gamma \vdash c \sim_{\prec} S_1 \Rightarrow S_2 : T \iff \forall \delta. (\Gamma \models \delta \text{ implies } \delta_1(c) \sim_{\prec} \delta_2(S_1) \Rightarrow \delta_2(S) : T)
\end{array}$$

Fig. 16. Blame-inexact correspondence for ψ into picky λ_C

As a corollary of Lemma 6.12 and Lemma 5.10, if a picky λ_C term evaluates to blame, then it corresponds to any λ_H term. This will be used extensively in the proofs below, as it allows us to eliminate many cases.

6.14 Lemma [Extra contracts]: If $t \sim_{\prec} s : T$ and $c \sim_{\prec} S_1 \Rightarrow S_2 : T$ then $\langle c \rangle^{l,l'} t \sim_{\prec} s : T$.

Proof

By induction on T . If $t \rightarrow_c^* \uparrow l''$ we are done, so let $t \rightarrow_c^* v$ and $s \rightarrow_h^* w$ such that $v \approx_{\prec} w : T$.

$T = B$: So $c = \{x:B \mid t_2\}$ and $S_2 = \{x:B \mid s_2\}$. Moreover, $v = w = k \in \mathcal{K}_B$, since those are the only corresponding values at B .

We can step and see $\langle c \rangle^{l,l'} t \rightarrow_c^* \langle c, t_2\{x := k\}, k \rangle^l$. We know that $t_2\{x := k\} \sim_{\prec} s_2\{x := k\} : \text{Bool}$. There are two possibilities: either $t_2\{x := k\} \rightarrow_c^* \uparrow l''$ or both terms go to corresponding Booleans. In the former case, the whole λ_C term goes to blame and we are done by Lemma 6.12. If both go to false, then the outer λ_C term evaluates to $\uparrow l$ and we're done by Lemma 6.12 again. If both go to true, then both outer terms go to k , and $k \approx_{\prec} k : B$.

$T = T_1 \rightarrow T_2$: So $c = x:c_1 \mapsto c_2$ and $S_1 = x:S_{11} \rightarrow S_{12}$ and $S_2 = x:S_{21} \rightarrow S_{22}$. Let $t' \sim_{\prec} s' : T_1$. If $t' \rightarrow_c^* \uparrow l''$ we're done by Lemma 6.12, so let $t' \rightarrow_c^* v'$ and $s' \rightarrow_h^* w'$, where $v' \approx_{\prec} w' : T_1$. We want to prove $\langle c \rangle^{l,l'} t' \sim_{\prec} s' s' : T_2$, which is true iff:

$$\langle c_2\{x := \langle c_1 \rangle^{l,l'} v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l,l'} v')) \sim_{\prec} w w' : T_2$$

By the IH on $v \sim_{\prec} w' : T_1$ and $c_1 \sim_{\prec} S_{21} \Rightarrow S_{11} : T_1$, we have $\langle c_1 \rangle^{l',l} v' \sim_{\prec} w' : T_1$. By definition, applying v and w yields related terms at T_2 . Since $\langle c_1 \rangle^{l',l} v' \sim_{\prec} w' : T_1$, we have $c_2\{x := \langle c_1 \rangle^{l',l} v'\} \sim_{\prec} S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^{l''} w'\} \Rightarrow S_{22}\{x := w'\} : T_2$. We can now apply the IH and see:

$$\langle c_2\{x := \langle c_1 \rangle^{l',l} v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l',l} v')) \sim_{\prec} w w' : T_2 \quad \square$$

6.15 Lemma [Contract/cast correspondence]: If $c \sim_{\prec} S_1 \Rightarrow S_2 : T$ and $t \sim_{\prec} s : T$ then $\langle c \rangle^{l,l'} t \sim_{\prec} \langle S_1 \Rightarrow S_2 \rangle^{l''} s : T$.

Proof

By induction on T . We reason via expansion (Lemma 6.11), showing that the initial terms reduce to corresponding terms.

$T = B$: So $c = \{x:B \mid t_1\}$, $S_1 = \{x:B \mid s_1\}$, and $S_2 = \{x:B \mid s_2\}$. Since $t \sim_{\prec} s : B$, we know that they either both reduce to $k \in \mathcal{K}_B$ or $t \rightarrow_c^* k$ and $s \rightarrow_h^* k$. If the latter is the case, we are done. So suppose $t \rightarrow_c^* k$ along with $s \rightarrow_h^* k$.

We can step our terms into active checks as follows, then:

$$\begin{aligned} \langle \{x:B \mid t_1\} \rangle^{l,l} t &\rightarrow_c^* \langle \{x:B \mid t_1\}, t_1\{x := k\}, k \rangle^l \\ \langle \{x:B \mid s_1\} \Rightarrow \{x:B \mid s_2\} \rangle^l s &\rightarrow_h^* \langle \{x:B \mid s_2\}, s_2\{x := k\}, k \rangle^l \end{aligned}$$

By the contract/cast correspondence, we know that $t_1\{x := k\} \sim_{\prec} s_2\{x := k\} : \text{Bool}$, so either $t_1\{x := k\}$ goes to blame or both terms go to a Bool together. In the former case, the outer λ_C term goes to blame and we are done by Lemma 6.12.. If they go to false, then both the active check goes to $\uparrow l$ and we are done, again by Lemma 6.12. Finally, if they both go to true, then both terms will evaluate to $k \in \mathcal{K}_B$, and $k \approx_{\prec} k : B$.

$T = T_1 \rightarrow T_2$: $c = x:c_1 \mapsto c_2$, $S_1 = x:S_{11} \rightarrow S_{12}$, and $S_2 = x:S_{21} \rightarrow S_{22}$. We know by inversion of the contract/cast relation that $c_1 \sim_{\prec} S_{21} \Rightarrow S_{11} : T_1$ and that for all l' and $t \sim_{\prec} s : T_1$, $c_2\{x := t\} \sim_{\prec} S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^{l''} s\} \Rightarrow S_{22}\{x := s\} : T_2$. We want to prove that $\langle c \rangle^{l,l'} \sim_{\prec} \langle S_1 \Rightarrow S_2 \rangle^l s : T_1 \rightarrow T_2$. First, we can assume $t \rightarrow_c^* v$ and $s \rightarrow_h^* w$ where $v \sim_{\prec} w : T_1 \rightarrow T_2$ —if not, both the contracted term goes to blame and we are done by Lemma 6.12.

We show that the decomposition of the contract and cast terms correspond for all inputs. Let $t' \sim_{\prec} s' : T_1$. Again, we can assume that they reduce to $v' \sim_{\prec} w' : T_1$, or else we are done by blame lifting in λ_C . On the λ_C side, we have

$$(\langle c \rangle^{l,l'} t) t' \rightarrow_c^* \langle c_2\{x := \langle c_1 \rangle^{l',l} v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l',l} v'))$$

In λ_H , we find

$$\begin{aligned} (\langle S_1 \Rightarrow S_2 \rangle^{l''} s) s' &\rightarrow_h^* \\ \langle S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^{l''} w\} \Rightarrow S_{22}\{x := w'\} \rangle^{l''} (w (\langle S_{21} \Rightarrow S_{11} \rangle^{l''} w')) \end{aligned}$$

By the IH, we know that $\langle c_1 \rangle^{l',l} v' \sim_{\prec} \langle S_{21} \Rightarrow S_{11} \rangle^{l''} w' : T_1$. Since $v \sim_{\prec} w : T_1 \rightarrow T_2$, we have $v (\langle c_1 \rangle^{l',l} v') \sim_{\prec} w (\langle S_{21} \Rightarrow S_{11} \rangle^{l''} w') : T_2$. By Lemma 6.14, $\langle c_1 \rangle^{l',l} v' \sim_{\prec} w' : T_1$. We can then see that $c_2\{x := \langle c_1 \rangle^{l',l} v'\} \sim_{\prec} S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^{l''} w'\} \Rightarrow S_{22}\{x := w'\} : T_2$.

$S_{22}\{x := w'\} : T_2$. By the IH, we therefore have

$$\begin{aligned} & \langle c_2\{x := \langle c_1 \rangle^{l',l} v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l',l} v')) \sim_{\prec} \\ & \langle S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^{l''} w'\} \Rightarrow S_{22}\{x := w'\} \rangle^{l''} w (\langle S_{21} \Rightarrow S_{11} \rangle^{l''} w') : T_2 \quad \square \end{aligned}$$

6.16 Theorem [Behavioral correspondence]:

1. If $\Delta \vdash s : S$ then $\lfloor \Delta \rfloor \vdash \psi(s) \sim_{\prec} s : \lfloor S \rfloor$.
2. If $\Delta \vdash S_1$ and $\Delta \vdash S_2$, where $\lfloor S_1 \rfloor = \lfloor S_2 \rfloor = \lfloor S \rfloor$, then $\lfloor \Delta \rfloor \vdash \psi^l(S_1, S_2) \sim_{\prec} S_1 \Rightarrow S_2 : \lfloor S \rfloor$.

Proof

By an induction similar to the proof Theorem 5.18. \square

6.17 Theorem [Type preservation for ψ]: For programs without active checks, if $\vdash \Delta$, then:

1. If $\Delta \vdash s : S$ then $\lfloor \Delta \rfloor \vdash \psi(s) : \lfloor S \rfloor$.
2. If $\Delta \vdash S_1$, $\Delta \vdash S_2$, where $\lfloor S_1 \rfloor = \lfloor S_2 \rfloor = T$, then $\lfloor \Delta \rfloor \vdash^{l,l'} \psi^l(S_1, S_2) : T$.

Proof

By an induction similar to the proof Theorem 5.20. \square

Here is an example where a λ_H term reduces to a value while its ψ -image in picky λ_C term reduces to blame:

$$\begin{aligned} S_1 &= f : S_{11} \rightarrow S_{12} \\ &= f : (x : \lfloor \text{Int} \rfloor \rightarrow \{y : \text{Int} \mid \text{nonzero } y\}) \rightarrow \lfloor \text{Int} \rfloor \\ S_2 &= f : S_{21} \rightarrow S_{22} \\ &= f : (x : \lfloor \text{Int} \rfloor \rightarrow \lfloor \text{Int} \rfloor) \rightarrow \{z : \text{Int} \mid f \ 0 = 0\} \\ c &= \psi^l(S_1, S_2) \\ &= f : \psi^l(S_{21}, S_{11}) \mapsto \psi^l(S_{12}\{f := \langle S_{21} \Rightarrow S_{11} \rangle^l f\}, S_{22}) \\ &= f : (x : \{x : \text{Int} \mid \text{true}\} \mapsto \{y : \text{Int} \mid \text{nonzero } y\}) \mapsto \{z : \text{Int} \mid f \ 0 = 0\} \end{aligned}$$

Let $w = (\lambda f : (x : \{x : \text{Int} \mid \text{true}\} \rightarrow \{y : \text{Int} \mid \text{nonzero } y\}). 0)$ and $w' = (\lambda x : \{x : \text{Int} \mid \text{true}\}. 0)$. The term is well typed: we can show $\emptyset \vdash w : S_1$ and $\emptyset \vdash w' : S_{21}$. Therefore $\emptyset \vdash (\langle S_1 \Rightarrow S_2 \rangle^l w) w' : S_{22}\{f := w'\}$. Translating, we find

$$\psi((\langle S_1 \Rightarrow S_2 \rangle^l w) w') = (\langle \psi^l(S_1, S_2) \rangle^{l,l} \psi(w)) \psi(w') = (\langle c \rangle^{l,l} \lambda f : \text{Int}. 0) \lambda x : \text{Int}. 0.$$

On the one hand $(\langle S_1 \Rightarrow S_2 \rangle^l w) w' \rightarrow_h^* 0$, while $(\langle c \rangle^{l,l} \lambda f : \text{Int}. 0) \lambda x : \text{Int}. 0 \rightarrow_c^* \uparrow l$. This means we cannot hope to use ψ as an exact correspondence between λ_H and picky λ_C . (Removing the extra cast ψ inserts into S_{12} doesn't affect our example, since ψ ignores S_{12} here.) For example,

$$\psi^l(\{z : \text{Int} \mid \text{true}\}\{f := \langle S_{21} \Rightarrow S_{11} \rangle^l f\}, \{z : \text{Int} \mid f \ 0 = 0\}) = \{x : B \mid \psi(f \ 0 = 0)\}.$$

6.3 Restricted calculi

While ϕ from lax λ_C and ψ to picky λ_C don't induce exact behavioral correspondences in the dependent case, some reasonable restrictions of the languages *are* equivalent.

Contract/type correspondence

$$\{x:B \mid t\} \sim^{l,l'} \{x:B \mid s\} : B \iff \forall k \in \mathcal{K}_B, t\{x := k\} \sim s\{x := k\} : \text{Bool}$$

$$\begin{aligned} x:c_1 \mapsto c_2 \sim^{l,l'} x:S_1 \rightarrow S_2 : T_1 \rightarrow T_2 &\iff \\ c_1 \sim^{l,l'} S_1 : T_1 \wedge \forall t \sim s : T_1, c_2\{x := t\} \sim^{l,l'} S_2\{x := s\} : T_2 & \end{aligned}$$

Dual closing substitutions

$$\Gamma \models \delta \iff \begin{cases} \forall x:T \in \Gamma, \delta_1(x) \sim \delta_2(x) : T \\ \forall x:c^{l,l'} \in \Gamma, \delta_1(x) \sim \delta_2(x) : \lfloor c \rfloor \end{cases}$$

Lifted to open terms

$$\begin{aligned} \Gamma \vdash t \sim s : T &\iff \forall \delta. (\Gamma \models \delta \text{ implies } \delta_1(t) \sim \delta_2(s) : T) \\ \Gamma \vdash c \sim^{l,l'} S : T &\iff \forall \delta. (\Gamma \models \delta \text{ implies } \delta_1(c) \sim^{l,l'} \delta_2(S) : T) \end{aligned}$$

Fig. 17. Blame-exact correspondence for ϕ from first-order picky λ_C

Gronski & Flanagan (2007) have already shown that ϕ induces an exact correspondence on the nondependent restriction. Since the lax/picky distinction requires dependency, exact equivalence in the nondependent case is a restriction of the results of Section 5.

Moreover, the first-order dependent restrictions of λ_C and λ_H also correspond exactly. The intuition here is that rechecking a first-order contract in a new context can't change the result of checking—first-order contracts can't be abusive. We can show this for ϕ using our existing parallel reduction for λ_H . We can show it for ψ , as well, using a similar parallel reduction for λ_C . For this second proof we assume (but do not prove) that evaluation and reduction in λ_C commute as they do in λ_H .

6.18 Lemma [Bulletproofing]: If $t \sim s : T$ and $c \sim^{l,l'} S : T$ then $\langle c \rangle^{l,l'} t \sim \langle S \Rightarrow \lceil S \rceil \rangle^{l'} \langle \lceil S \rceil \Rightarrow S \rangle^{l'} s : T$.

Proof

By induction on T . First, observe that either both t and s go to $\uparrow^{l''}$ or both t and s go to values related at T . In the former case, the outer terms also go to blame. So we only consider the case where $t \rightarrow_c^* v$, $s \rightarrow_h^* w$, and $v \approx w : T$.

$T = B$: So $c = \{x:B \mid t_1\}$ and $S = \{x:B \mid s_1\}$ and $S' = \{x:B \mid s_2\}$. By Lemma 5.5 we have $\langle c \rangle^{l,l'} t \sim \langle \lceil S \rceil \Rightarrow S \rangle^{l'} s : B$. By Lemma 5.4 we can add the extra, trivial cast $\langle S \Rightarrow \lceil S \rceil \rangle^{l'}$.

$T = B \rightarrow T_2$: We know that $c = x:\{x:B \mid t_1\} \mapsto c_2$ and $S = x:\{x:B \mid s_1\} \rightarrow S_2$. Let $t' \sim s' : T_1$. We only need to consider the case where $t' \rightarrow_c^* v'$ and $s' \rightarrow_h^* w'$ —if $t' \rightarrow_c^* \uparrow^{l''}$ and $s' \rightarrow_h^* \uparrow^{l''}$ the outer terms correspond because both blame l'' .

On the λ_C side, $(\langle c \rangle^{l,l'} t) t' \rightarrow_c^* \langle c_2\{x := v'\} \rangle^{l,l'} (v (\langle c_1 \rangle^{l,l'} v'))$. In λ_H , we can see

$$\begin{aligned} (\langle S \Rightarrow \lceil S \rceil \rangle^{l'} \langle \lceil S \rceil \Rightarrow S \rangle^{l'} s) s' &\rightarrow_h^* \\ (S_2\{x := \langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} w'\} \Rightarrow \lceil S_2 \rceil)^{l'} ((\langle \lceil S \rceil \Rightarrow S \rangle^{l'} w) (\langle \lceil S_1 \rceil \Rightarrow S_1 \rangle^{l'} w')) & \end{aligned}$$

Since $v' \approx w' : B$, we must have $v' = w' = k \in \mathcal{K}_B$. By Lemma 5.5 and $\{x:B \mid t_1\} \sim^{l',l} \{x:B \mid s_1\} : B$, we know that $\langle \{x:B \mid t_1\} \rangle^{l',l} v' \sim \langle \lceil B \rceil \Rightarrow \{x:B \mid s_1\} \rangle^{l'} w' : B$. Both terms go to blame or to the same value—which must be k , from inspection of the contract and cast evaluation rules. The former case is immediate, since the outer terms then

go to blame. So suppose $\langle \{x:B \mid t_1\} \rangle^{l',l} k \longrightarrow_c^* k$ and $\langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k \longrightarrow_h^* k$. Now the terms evaluate like so:

$$\begin{aligned} \langle c_2\{x := v'\} \rangle^{l,l'} (v (\langle \{x:B \mid t_1\} \rangle^{l',l} v')) &\longrightarrow_c^* \langle c_2\{x := k\} \rangle^{l,l'} (v k) \\ \langle S_2\{x := \langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k \Rightarrow [S_2]\rangle^{l'} & \\ ((\langle [S] \Rightarrow S \rangle^l w) (\langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k)) &\longrightarrow_h^* \\ \langle S_2\{x := \langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k \Rightarrow [S_2]\rangle^{l'} & \\ \langle [S_2] \Rightarrow S_2\{x := k\} \rangle^l (w (\langle \{x:B \mid s_1\} \Rightarrow [B] \rangle^l k)) & \end{aligned}$$

By Lemma 5.4, $k \sim \langle \{x:B \mid s_1\} \Rightarrow [B] \rangle^l k : B$, so $v k \sim w (\langle \{x:B \mid s_1\} \Rightarrow [B] \rangle^l k) : T_2$. We have by definition (and $k \sim k : B$) that $c_2\{x := k\} \sim^{l,l'} S_2\{x := \langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k \rangle^{l'} : T_2$. Recall that $\langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k \longrightarrow_h^* k$. This implies $\langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k \Rightarrow^* k$ (Lemma A4 in the Coq). We can then see that $S_2\{x := \langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k \rangle^{l'} \Rightarrow^* S_2\{x := k\}$ by Lemma A1 in the Coq. By extension with the congruence rules:

$$\begin{aligned} \langle S_2\{x := \langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k \Rightarrow [S_2]\rangle^{l'} & \\ \langle [S_2] \Rightarrow S_2\{x := \langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k \rangle^{l'} \rangle^l (w (\langle \{x:B \mid s_1\} \Rightarrow [B] \rangle^l k)) &\Rightarrow \\ \langle S_2\{x := \langle [B] \Rightarrow \{x:B \mid s_1\} \rangle^{l'} k \Rightarrow [S_2]\rangle^{l'} & \\ \langle [S_2] \Rightarrow S_2\{x := k\} \rangle^l (w (\langle \{x:B \mid s_1\} \Rightarrow [B] \rangle^l k)) & \end{aligned}$$

By the IH $\langle c_2\{x := k\} \rangle^{l,l'} (v k)$ corresponds to the former, which means it is related to the latter by Lemma 5.3. We conclude by expansion (Lemma 6.1). \square

6.19 Theorem [Behavioral correspondence]: If $\vdash \Gamma$, then:

1. If $\phi(\Gamma \vdash t : T) = s$ then $\Gamma \vdash t \sim s : T$.
2. If $\phi(\Gamma \vdash^{l,l'} c : T) = S$ then $\Gamma \vdash c \sim^{l,l'} S : T$.

Proof

We simultaneously show both properties by induction on the depth of ϕ 's recursion. To show $\Gamma \vdash t \sim s : T$, let $\Gamma \models \delta$ —we will show $\delta_1(t) \sim \delta_2(s) : T$.

The proof proceeds by case analysis on the final rule of the translated typing and well-formedness derivations. \square

Let's pretend we've developed a theory of parallel reduction for picky λ_C and that the correspondence is closed under parallel reduction on the left.

6.20 Lemma [Contract/cast correspondence]: If $c \sim S_1 \Rightarrow^l S_2 : T$ and $t \sim s : T$ then $\langle c \rangle^{l,l} t \sim \langle S_1 \Rightarrow S_2 \rangle^l s : T$.

Proof

By induction on T . We reason via expansion (Lemma 5.10), showing that the initial terms reduce to corresponding terms.

$T = B$: So $c = \{x:B \mid t_1\}$, $S_1 = \{x:B \mid s_1\}$, and $S_2 = \{x:B \mid s_2\}$. Since $t \sim s : B$, we know that they either both reduce to $k \in \mathcal{K}_B$ or $\uparrow^{l'}$. If the latter is the case, we are done. So suppose $t \longrightarrow_c^* k$ along with $s \longrightarrow_h^* k$.

Contract/type correspondence

$$\boxed{c \sim S_1 \Rightarrow^l S_2 : T}$$

$$\{x:B \mid t\} \sim \{x:B \mid s_1\} \Rightarrow^l \{x:B \mid s_2\} : B \iff \forall k \in \mathcal{K}_B. t\{x := k\} \sim s_2\{x := k\} : \text{Bool}$$

$$x:c_1 \mapsto c_2 \sim x:S_{11} \rightarrow S_{12} \Rightarrow^l S_{21} \rightarrow S_{22} : T_1 \rightarrow T_2 \iff c_1 \sim S_{21} \Rightarrow^l S_{11} : T_1 \wedge \forall t \sim s : T_1. c_2\{x := t\} \sim S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l s\} \Rightarrow^l S_{22}\{x := s\} : T_2$$

Closing substitutions

$$\Gamma \models \delta \iff \forall x \in \text{dom}(\Gamma). \delta_1(x) \sim_{\prec} \delta_2(x) : [\Gamma(x)]$$

Lifted to open terms

$$\begin{aligned} \Gamma \vdash t \sim s : T &\iff \forall \delta. (\Gamma \models \delta \text{ implies } \delta_1(t) \sim \delta_2(s) : T) \\ \Gamma \vdash c \sim S_1 \Rightarrow^l S_2 : T &\iff \forall \delta. (\Gamma \models \delta \text{ implies } \delta_1(c) \sim \delta_2(S_1) \Rightarrow^l \delta_2(S_2) : T) \end{aligned}$$

Fig. 18. Blame-exact correspondence for ψ into first-order lax λ_C

We can step our terms into active checks as follows, then:

$$\begin{aligned} \langle \{x:B \mid t_1\} \rangle^{l,l} t &\longrightarrow_c^* \langle \{x:B \mid t_1\}, t_1\{x := k\}, k \rangle^l \\ \langle \{x:B \mid s_1\} \Rightarrow \{x:B \mid s_2\} \rangle^l s &\longrightarrow_h^* \langle \{x:B \mid s_2\}, s_2\{x := k\}, k \rangle^l \end{aligned}$$

By inversion of the contract/cast correspondence, we know that $t_1\{x := k\} \sim s_2\{x := k\} : \text{Bool}$, so these terms go to blame or to a Bool together. If they go to $\uparrow l'$, we are done. If they go to false, then both the obligation and the cast will go to $\uparrow l$. Finally, if they both go to true, then both terms will evaluate to k .

$T = B \rightarrow T_2$: $c = x:\{x:B \mid t_1\} \mapsto c_2$, $S_1 = x:\{x:B \mid s_{11}\} \rightarrow S_{12}$, and $S_2 = x:\{x:B \mid s_{21}\} \rightarrow S_{22}$. We know by inversion of the contract/cast relation that $\{x:B \mid t_1\} \sim \{x:B \mid s_{21}\} \Rightarrow^l \{x:B \mid s_{11}\} : B$ and that for all $t' \sim s' : B$, $c_2\{x := t'\} \sim S_{12}\{x := s'\} \Rightarrow^l S_{22}\{x := s'\} : T_2$. We want to prove that $\langle c \rangle^{l,l} \sim \langle S_1 \Rightarrow S_2 \rangle^l s : B \rightarrow T_2$. First, we can assume $t \longrightarrow_c^* v$ and $s \longrightarrow_h^* w$ where $v \sim w : B \rightarrow T_2$ —if not, both cast and contracted terms go to blame and we are done.

We show that the decomposition of the contract and cast terms correspond for all inputs. Let $t' \sim s' : B$. Again, we can assume that they reduce to $v' \sim w' : B$, or else we are done by blame lifting. Moreover, we must have $v' = w' = k \in \mathcal{K}_B$. On the λ_C side, we have

$$(\langle c \rangle^{l,l} t) t' \longrightarrow_c^* \langle c_2\{x := \langle \{x:B \mid t_1\} \rangle^{l,l} k \} \rangle^{l,l} (v (\langle \{x:B \mid t_1\} \rangle^{l,l} k))$$

In λ_H , we find

$$\begin{aligned} &(\langle S_1 \Rightarrow S_2 \rangle^l s) s' \longrightarrow_h^* \\ &\langle S_{12}\{x := \langle \{x:B \mid s_{21}\} \Rightarrow \{x:B \mid s_{11}\} \rangle^l k \} \Rightarrow S_{22}\{x := k\} \rangle^l \\ &(w (\langle \{x:B \mid s_{21}\} \Rightarrow \{x:B \mid s_{11}\} \rangle^l k)). \end{aligned}$$

By the IH, we know that $\langle \{x:B \mid t_1\} \rangle^{l,l} k \sim \langle \{x:B \mid s_{21}\} \Rightarrow \{x:B \mid s_{11}\} \rangle^l k : B$. If they both go to blame, we're done, so let them both go to k .

Since $v \sim w : T_1 \rightarrow T_2$, we have $v (\langle \{x:B \mid t_1\} \rangle^{l,l} k) \sim w (\langle \{x:B \mid s_{21}\} \Rightarrow \{x:B \mid s_{11}\} \rangle^l k) : T_2$. Then, since $k \sim k : T_1$, we have $\langle c_2\{x := k\} \rangle^{l,l} (v (\langle \{x:B \mid t_1\} \rangle^{l,l} k)) \sim \langle S_{12}\{x := \langle \{x:B \mid s_{21}\} \Rightarrow \{x:B \mid s_{11}\} \rangle^l k \} \Rightarrow S_{22}\{x := k\} \rangle^l w (\langle \{x:B \mid s_{21}\} \Rightarrow \{x:B \mid s_{11}\} \rangle^l k)$.

$s_{11}\rangle^l k) : T_2$ by assumption. Since $\langle \{x:B \mid t_1\} \rangle^{l,l} k \longrightarrow_c^* k$, we can parallel reduce $\langle c_2\{x := \langle \{x:B \mid t_1\} \rangle^{l,l} k \rangle^{l,l} (v (\langle \{x:B \mid t_1\} \rangle^{l,l} k)) \rangle^{l,l} (v (\langle \{x:B \mid t_1\} \rangle^{l,l} k))$, showing

$$\begin{aligned} & \langle c_2\{x := k\} \rangle^{l,l} (v (\langle \{x:B \mid t_1\} \rangle^{l,l} k)) \sim \\ & \langle S_{12}\{x := \langle \{x:B \mid s_{21}\} \Rightarrow \{x:B \mid s_{11}\} \rangle^l k \rangle \Rightarrow S_{22}\{x := k\} \rangle^l \\ & w (\langle \{x:B \mid s_{21}\} \Rightarrow \{x:B \mid s_{11}\} \rangle^l k) : T_2. \quad \square \end{aligned}$$

6.21 Theorem [Behavioral correspondence]:

1. If $\Delta \vdash s : S$ then $[\Delta] \vdash \psi(s) \sim s : [S]$.
2. If $\Delta \vdash S_1$ and $\Delta \vdash S_2$, where $[S_1] = [S_2] = [S]$, then $[\Delta] \vdash \psi^l(S_1, S_2) \sim S_1 \Rightarrow^l S_2 : [S]$ (for all l).

Proof

By an induction similar to the proof of Theorem 5.18. \square

6.4 Alternative calculi

There are three alternative calculi we have not considered here: indy λ_C (Dimoulas *et al.*, 2011), superpicky λ_H , and nonterminating calculi. We describe them in detail below, but we leave them as future work.

Dimoulas *et al.* (2011) add a third blame label to λ_C , representing the contract itself; we write it here as a subscript. They accordingly change the picky E_CDECOMP rule:

$$(\langle x:c_1 \mapsto c_2 \rangle_{l''}^{l,l'} v_1) v_2 \longrightarrow_c \langle c_2\{x := \langle c_1 \rangle_{l''}^{l'',l} v_2 \rangle \rangle_{l''}^{l,l'} (v_1 (\langle c_1 \rangle_{l''}^{l'',l} v_2))$$

In the substitution in the codomain, note that the blame label on the domain contract uses the contract's blame label l'' . The intuition here is that any problem arising in c_2 is in the contract's context (label l''), not the original negative context (label l'). We conjecture (but have not proven) that indy λ_C is in the same position on the axis of blame as picky λ_C . We should only need to change the labels on the contracts ϕ inserts to have an exact correspondence; however, ψ will remain inexact.

Superpicky λ_H reworks the F_CDECOMP rule in an attempt to harmonize λ_H and picky λ_C semantics:¹⁰

$$\begin{aligned} & (\langle x:S_{11} \rightarrow S_{12} \Rightarrow x:S_{21} \rightarrow S_{22} \rangle^l w_1) w_2 \rightsquigarrow_h \\ & \langle S_{12}\{x := \langle S_{21} \Rightarrow S_{11} \rangle^l w_2\} \Rightarrow S_{22}\{x := \langle S_{11} \Rightarrow S_{21} \rangle^l (\langle S_{21} \Rightarrow S_{11} \rangle^l w_2)\} \rangle^l \\ & (w_1 (\langle S_{21} \Rightarrow S_{11} \rangle^l w_2)) \end{aligned}$$

This seems to resolve the problem with ψ into picky λ_C , but it poses problems in the proof semantic type soundness for λ_H : how do $S_{22}\{x := w_2\}$ and $S_{22}\{x := \langle S_{11} \Rightarrow S_{21} \rangle^l (\langle S_{21} \Rightarrow S_{11} \rangle^l w_2)\}$ relate?

Finally, we have been careful to ensure that all of our calculi are strongly normalizing. We don't believe this to be essential, though we would have to change our logical relations— λ_H 's type semantics and the correspondences—to account for nontermination. We conjecture that step-indexing (Ahmed, 2006) will suffice.

¹⁰ This idea is due to Jeremy Siek (personal communication, January 2010).

Table 1. *Comparison between contract systems*
Latent systems

	FF02 (1)	HJL06 (2)	GF07 λ_C (3)	BM06 (4)	DFFF11 (5)	our λ_C
dep (6)	✓ lax	✓ picky	×	(7)	✓ indy	✓ either
eval order	CBV	lazy	CBV	CBV	CBV	CBV
blame (8)	$\uparrow l$	$\uparrow l$	$\uparrow l$	$\uparrow l$ or \perp	$\uparrow l$	$\uparrow l$
checking (9)	if	if	\bigcirc	active	active	active
typing (10)	✓	✓	✓	n/a	✓	✓
any con (11)	✓	✓	✓	✓	✓	✓

	GF07 λ_H (3)	F06 (12)	KF10 (13)	WF09 (14)	OTMW04 (15)	BGIP11 (16)	our λ_H
dep (6)	×	✓	✓	×	✓	✓	✓
eval order	CBV	CBN(17)	full β	CBV	CBV	CBV	CBV
blame (8)	$\uparrow l$	stuck	stuck	$\uparrow l$	\uparrow	$\uparrow l$	$\uparrow l$
checking (9)	\bigcirc	\bigcirc	active	active	if	active	active
typing (10)	×	×	✓	✓	✓	✓	✓
any con (11)	✓	✓	✓	✓	×	✓	✓

(1) Findler & Felleisen (2002). (2) Hinze *et al.* (2006). (3) Gronski & Flanagan (2007). (4) Blume & McAllester (2006). (5) Dimoulas *et al.* (2011). (12) Flanagan (2006). (13) Knowles & Flanagan (2010). (14) Wadler & Findler (2009). (15) Ou *et al.* (2004). (6) Does the system include dependent contracts or function types (✓) or not (×) and, for latent systems, is the semantics lax or picky? (See below for more on “indy” checking.) (7) An “unusual” form of dependency, where negative blame in the codomain results in nontermination. (17) A nondeterministic variant of CBN. (8) Do failed contracts raise labeled blame ($\uparrow l$), raise blame without a label (\uparrow), get stuck, or sometimes raise blame and sometimes diverge (\perp)? (9) Is contract or cast checking performed using an “active check” syntactic form (active), an “if” construct with a refined typing rule (if), or “inlined” by making the operational semantics refer to its own reflexive and transitive closure (\bigcirc)? (10) Is the typing relation well defined (i.e., for dependently typed systems, is it based on a type semantics or, as in WF09, a “tagging” scheme), or is the definition circular? (11) Are arbitrary user-defined boolean functions allowed as contracts or refinements (✓), or only built-in ones (×)?

7 Related work

Conferences in recent years have seen a profusion of papers on higher-order contracts and related features. This is all to the good, but for newcomers to the area it can be a bit overwhelming, especially given the great variety of technical approaches. To help reduce the level of confusion, in Table 1 we summarize the important points of comparison between a number of systems that are closely related to ours. This table is an updated version compared to that in Greenberg *et al.* (2010).

The largest difference is between latent and manifest treatments of contracts—i.e., whether contract checking (under whatever name) is a completely dynamic matter or whether it leaves a “trace” that the type system can track.

Another major distinction (labeled “dep” in the figure) is the presence of dependent contracts or, in manifest systems, dependent function types. Latent systems with dependent contracts also vary in whether their semantics is lax or picky.

Next, most contract calculi use a standard call-by-value order of evaluation (“eval order” in the figure). Notable exceptions include those of Hinze *et al.* (2006), which is embedded in Haskell, Flanagan (2006), which uses a variant of call-by-name, and Knowles & Flanagan (2010), which uses full β -reduction (more on this below).

Another point of variation (“blame” in the figure) is how contract violations or cast failures are reported—by raising an exception or by getting stuck. We also return to this below.

The next two rows in the table (“checking” and “typing”) concern more technical points in the papers most closely related to ours. In both Gronski & Flanagan (2007) and Flanagan (2006), the operational semantics checks casts “all in one go”:

$$\frac{s_2\{x := k\} \longrightarrow_h^* \text{true}}{\langle \{x:B \mid s_1\} \Rightarrow \{x:B \mid s_2\} \rangle^l k \longrightarrow_h k}$$

Such rules are formally awkward, and in any case they violate the spirit of a small-step semantics. Also, the formal definitions of λ_H in both Gronski & Flanagan (2007) and Flanagan (2006) involve a circularity between the typing, subtyping, and implication relations. Knowles & Flanagan (2010) improve the technical presentation of λ_H in both respects. In particular, they avoid circularity (as we do) by introducing a denotational interpretation of types and maintain small-step evaluation by using a new syntactic form of “partially evaluated casts” (like most of the other systems).

The main contributions of the present paper are (1) the dependent translations ϕ and ψ and their properties, and (2) the formulation and metatheory of dependent λ_H . (Dependent λ_C is not a contribution on its own: many similar systems have been studied, and in any case its properties are simple.) The nondependent part of our ϕ translation essentially coincides with the one studied by Gronski & Flanagan (2007), and our behavioral correspondence theorem is essentially the same as theirs. Our ψ translation completes their story for the nondependent case, establishing a tight connection between the systems. The full dependent forms of ϕ and ψ studied here are novel, as is the observation that the correspondence between the latent and manifest worlds is more problematic in this setting.

Our formulation of λ_H is most comparable to that of Knowles & Flanagan (2010), but there are some significant differences. First, our cast-checking constructs are equipped with labels, and failed casts go to explicit blame—i.e., they raise labeled exceptions. In the λ_H of Knowles and Flanagan (though not the earlier one of Gronski and Flanagan), failed casts are simply stuck terms—their progress theorem says “If a well-typed term cannot step, then either it is a value or it contains a stuck cast.” Second, their operational semantics uses full, non-deterministic β -reduction, rather than specifying a particular order of reduction, as we have done. This significantly simplifies parts of the metatheory by allowing them to avoid introducing parallel reduction. We prefer standard call-by-value reduction because we consider blame as an exception—a computational effect—and we want to be able to reason about *which* blame will be raised by expressions involving many casts.

The system studied by Ou *et al.* (2004) is also close in spirit to our λ_H . The main difference is that, because their system includes general recursion, they restrict the terms

that can appear in contracts to just applications involving predefined constants: only “pure” terms can be substituted into types, and these do not include lambda-abstractions. Our system (like all of the others in Table 1—see the row labeled “any con”) allows arbitrary user-defined boolean functions to be used as contracts.

Our description of λ_C is ultimately based on λ_{CON} (Findler & Felleisen, 2002), though our presentation is slightly different in its use of checks. Hinze *et al.* (2006) adapted Findler & Felleisen-style contracts to a location-passing implementation in Haskell, using picky dependent function contracts.

Our λ_H type semantics in Section 4 is effectively a semantics of contracts. Blume & McAllester (2006) offers a semantics of contracts that is slightly different—our semantics includes blame at every type, while theirs explicitly excludes it. Xu *et al.* (2009) is also similar, though their “contracts” have no dynamic semantics at all: they are simply specifications.

Dimoulas *et al.* (2011) introduce a new dialect of picky λ_C , where contract checks in the codomain are given a distinct negative label. If labels represent “contexts” for values, then this treats the contract as an independent context. “Indy” λ_C and picky λ_C will raise exactly the same *amount* of blame, but they will blame different labels.

Belo *et al.* (2011) at once simplify and extend the CBV λ_H given here. The type system is redesigned to avoid subtyping and closing substitutions, so type soundness is proved with easy syntactic methods (Wright & Felleisen, 1992). What’s more, the language allows general refinements—refinements of any type, not just base types—and extends the type system to polymorphism. This can be seen as completing some of the future work of Greenberg *et al.* (2010).

We have discussed only a small sample of the many papers on contracts and related ideas. We refer the reader to Knowles & Flanagan (2010) for a more comprehensive survey. Another useful resource is Wadler & Findler (2007) (technically superseded by Wadler & Findler (2009), but with a longer related work section), which surveys work combining contracts with type Dynamic and related features.

There are also *many* other systems that employ various kinds of precise types, but in a completely static manner. One notable example is the work of Xu *et al.* (2009), which uses user-defined boolean predicates to classify values (justifying their use of the term “contracts”) but checks statically that these predicates hold.

Sage (Knowles *et al.*, 2006) and Knowles & Flanagan (2010) both support mixed static and dynamic checking of contracts, using, e.g., a theorem prover. We have not addressed this aspect of their work, since we have chosen to work directly with the core calculus λ_H , which for them was the target of an elaboration function.

8 Conclusion

We can faithfully encode dependent λ_H into λ_C —the behavioral correspondence is tight. λ_H ’s F_CDECOMP rule forces us to accept a weaker behavioral correspondence when encoding λ_C into λ_H , so we conclude that the manifest and latent approaches are *not* equivalent in the dependent case. We do find, however, that the two approaches are entirely inter-encodable in the nondependent restriction.

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