

Higher-ranked Exception Types

Ruud Koot

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Contents

1	Higher-ranked Exception Types	7
1.1	Motivation	7
1.1.1	Overview	10
1.1.2	Contributions	10
1.2	The λ^U -calculus	11
1.2.1	Typing relation	11
1.2.2	Semantics	11
1.2.3	Normalization	13
1.2.4	Pattern unification	14
1.3	Source language	15
1.3.1	Underlying type system	15
1.3.2	Operational semantics	15
1.4	Exception types	16
1.4.1	Exception types	16
1.4.2	Conservativeness	18
1.4.3	Exception type completion	19
1.4.4	Declarative exception type system	19
1.4.5	Type elaboration system	20
1.5	Type inference algorithm	26
1.5.1	Subtyping	26
1.6	Interesting observations	26
1.7	Metatheory	26
1.7.1	Declarative type system	26
1.7.2	Syntax-directed type elaboration	31
1.7.3	Type inference algorithm	31

1.8	Further research	31
1.9	Related work	33
1.9.1	Higher-ranked polymorphism in type-and-effect systems	33
1.9.2	λ^U -calculus	35
1.9.3	Exception analyses	35
1.10	Conclusion	36
A	Old stuff	41
A.1	λ^U -calculus	41
A.1.1	Reduction relation (wrong!)	41
A.1.2	Semantics	43
A.1.3	Normalization (with widening)	44
A.2	Completion	45
A.3	TODO	46
A.3.1	Introduction	46
A.3.2	The λ^U -calculus	47

List of Figures

1.1	λ^U -calculus: syntax	11
1.2	λ^U -calculus: type system	12
1.3	λ^U -calculus: denotational semantics	13
1.4	Normalization algorithm for λ^U -terms.	15
1.5	Source language: syntax	16
1.6	Underlying type system ($\Gamma \vdash t : \tau$)	21
1.7	Operational semantics ($t_1 \longrightarrow t_2$)	22
1.8	Exception types: syntax	23
1.9	Exception types: well-formedness	23
1.10	Type completion ($\Gamma \vdash \tau : \hat{\tau} \ \& \ \xi \triangleright \Gamma'$)	24
1.11	Declarative type system ($\Gamma; \Delta \vdash t : \hat{\tau} \ \& \ \xi$)	24
1.12	Syntax-directed type elaboration system ($\Gamma; \Delta \vdash t \hookrightarrow t' : \hat{\tau} \ \& \ \xi$)	25
1.13	Type inference algorithm	37
1.14	Subtyping	38
1.15	Annotation substitution	38
1.16	Term substitution	39
A.1	Normalization algorithm for λ^U -terms.	45

Chapter 1

Higher-ranked Exception Types

1.1 Motivation

An often heard selling point of non-strict functional languages is that they provide strong and expressive type systems that make side-effects explicit. This supposedly makes software more reliable by lessening the mental burden of programmers. Many object-oriented programmers are quite surprised, then, that when they make the transition to a functional language, that they lose a feature their type system formerly did provide: tracking of uncaught exceptions.

There is a good excuse why this feature is missing from the type systems of contemporary non-strict functional languages: in a strict first-order language it is sufficient to annotate each function with a single set of uncaught exceptions the function may throw, in a non-strict higher-order language the situation becomes significantly more complicated. Let us first consider the two aspects “higher-order” and “non-strict” in isolation:

Higher-order functions The set of exceptions that may be raised by a higher-order function are not given by a fixed set of exceptions, but depends on the set of exceptions that may be raised by the function that is passed as its functional argument. Higher-order functions will thus end up being

exception polymorphic.

Non-strict evaluation In non-strictly evaluated languages, exceptions are not a form of control flow, but a kind of value. Typically the set of values of each type are extended with an *exceptional value* \bot (more commonly denoted \perp , but we shall not do so for reasons of ambiguity), or family of exceptional values \bot^ℓ . This means we do not only need to give all functions an exception-annotated function type, but every expression an exception-annotated type.

Now let us consider these two aspects in combination. Take as an example the *map* function:

$$\begin{aligned} \text{map} &: \forall \alpha \beta. (\alpha \rightarrow \beta) \rightarrow [\alpha] \rightarrow [\beta] \\ \text{map} &= \lambda f. \lambda xs. \text{case } xs \text{ of} \\ &\quad [] \quad \mapsto [] \\ &\quad (y :: ys) \mapsto f y :: \text{map } f \text{ } ys \end{aligned}$$

For each type τ , we denote its exception-annotated type by $\tau\langle\tilde{\xi}\rangle$. For function types we will write $\tau_1\langle\tilde{\xi}_1\rangle \xrightarrow{\tilde{\xi}} \tau_2\langle\tilde{\xi}_2\rangle$ instead of $(\tau_1\langle\tilde{\xi}_1\rangle \rightarrow \tau_2\langle\tilde{\xi}_2\rangle)\langle\tilde{\xi}\rangle$. If $\tilde{\xi}$ is the empty exception set, then we will omit it completely.

The fully exception-polymorphic and exception-annotated type, or *exception type*, of *map* is **To do**. cramped formatting

$$\begin{aligned} &\forall \alpha \beta e_2 e_3. (\forall e_1. \alpha\langle e_1\rangle \xrightarrow{e_3} \beta\langle e_2 e_1\rangle) \\ &\rightarrow (\forall e_4 e_5. [\alpha\langle e_4\rangle]\langle e_5\rangle \rightarrow [\beta\langle e_2 e_4 \cup e_3\rangle]\langle e_5\rangle) \end{aligned}$$

The exception type of the first argument $\forall e_1. \alpha\langle e_1\rangle \xrightarrow{e_3} \beta\langle e_2 e_1\rangle$ states that it can be instantiated with a function that accepts any exceptional value as its argument (as the exception set e_1 is universally quantified) and returns a possibly exceptional value. In case the return value is exceptional, then it will be one from the exception set $e_2 e_1$. Here e_2 is an *exception operator*—a function that takes a number of exception sets and exception operators, and transforms it into another exception set, for example by adding a number of new elements to it, or discarding it and returning the empty set. Furthermore, the function itself may be an exceptional value from the exception set e_3 .

The exception type of the second argument $[\alpha\langle e_4\rangle]\langle e_5\rangle$ states it should be a list. Any of the elements in the lists may be exception values from the

exception set e_4 . Any of the constructors that form the spine of the list must be exceptional values from the exception set e_5 .

The result of *map* will be a list with the exception type $[\beta\langle e_2\ e_4 \cup e_3 \rangle]\langle e_5 \rangle$. Any constructors in the spine of this list may be exceptional values from the exception set e_5 , the same exception set as where exceptional values in the spine of the input list could come from. By looking at the definition of *map* we can see why this is the case: *map* will only produce non-exceptional constructors, but the pattern-match on the input list will propagate any exceptional values encountered there. The elements of the list are produced by the function application $f\ y$. Recall that f has the exception type $\forall e_1.\alpha\langle e_1 \rangle \xrightarrow{e_3} \beta\langle e_2\ e_1 \rangle$. Now one of two things can happen:

1. If f is an exceptional function value, then it must be one from the exception set e_3 . Applying an argument to an exceptional value will cause the exceptional value to be propagated.
2. Otherwise f is a non-exceptional value. The argument y has exception type $\alpha\langle e_4 \rangle$ —it is an element from the input list—and so can only be applied to f if we instantiate e_1 to e_4 first. If $f\ y$ will produce an exceptional value it will thus be on from the exception set $e_2\ e_4$.

To account for both cases we need to take the union of the two exception sets, giving us a value with the exception type $\beta\langle e_2\ e_4 \cup e_3 \rangle$.

The get a better feeling of how these exception type and exception operators behave let us see what happens when we apply two different functions to *map*: the identity function *id* and the constant exceptional values *const* \perp^E . These two functions can be given the exception types:

$$\begin{aligned} id & : \forall e_1.\alpha\langle e_1 \rangle \xrightarrow{\emptyset} \alpha\langle e_1 \rangle \\ const\ \perp^E & : \forall e_1.\alpha\langle e_1 \rangle \xrightarrow{\emptyset} \beta\langle \{E\} \rangle \end{aligned}$$

The term *id* simply propagates its input, so it will also propagate any exceptional values. The term *const* \perp^E discards its input and will always return the exceptional value \perp^E . This behavior is also reflected in their exception types.

If we apply *map* to *id* we need to unify the exception type of the formal parameter $\forall e_1.\alpha\langle e_1 \rangle \xrightarrow{e_3} \beta\langle e_2\ e_1 \rangle$ with the exception type of the actual parameter $\forall e_1.\alpha\langle e_1 \rangle \xrightarrow{\emptyset} \alpha\langle e_1 \rangle$. This can be accomplished by instantiating e_3 to \emptyset and e_2 to $\lambda x.x$, as $(\lambda x.x)\ e_1 \rightsquigarrow e_1$. This gives us the resulting exception type

$$\text{map id} : \forall \alpha \, e_4 \, e_5. [\alpha \langle e_4 \rangle] \langle e_5 \rangle \xrightarrow{\emptyset} [\alpha \langle e_4 \rangle] \langle e_5 \rangle$$

I.e., mapping the identity function over a list will propagate all existing exceptional values in the list and add no new ones.

If we apply *map* to $\text{const } \perp^E$ we need to unify the exception type of the formal parameter with $\forall e_1. \alpha \langle e_1 \rangle \xrightarrow{\emptyset} \beta \langle \{E\} \rangle$, which can be accomplished by instantiating e_3 to \emptyset and e_2 to $\lambda x. \{E\}$, as $(\lambda x. \{E\}) \, e_1 \rightsquigarrow \{E\}$. This gives us the resulting exception type

$$\text{map} (\text{const } \perp^E) : \forall \alpha \, \beta \, e_4 \, e_5. [\alpha \langle e_4 \rangle] \langle e_5 \rangle \xrightarrow{\emptyset} [\beta \langle \{E\} \rangle] \langle e_5 \rangle$$

I.e., mapping the constant exceptional value over a list will discard all existing exceptional values from the list and only output non-exceptional values or the exceptional value \perp^E as elements of the lists.

1.1.1 Overview

In Section 1.2 we will introduce the λ^U -calculus, a simply typed λ -calculus embellished with an associative, commutative, idempotent and unit (ACI1) structure. The λ^U -calculus will form language of effects in our type-and-effect system. Section 1.3 describes the language to which our analysis applies. In Section 1.4 we introduce the language of exception types and two type-and-effect systems for deriving exception types for a given term in the source language that is well-typed in the underlying type system: a declarative type system and a syntax-directed elaboration system that also produces an explicitly typed term. A type inference algorithm for this type-and-effect system is presented in Section 1.5.

1.1.2 Contributions

- A *type-and-effect system* that precisely tracks the uncaught exceptions using higher-ranked types.
- An *inference algorithm* that automatically infers such higher-ranked exception types.

1.2 The λ^U -calculus

The λ^U -calculus is simply typed λ -calculus extended with a set-union operator and singleton-set and empty-set constants at the term level.

Types

$$\begin{array}{ll} \tau \in \mathbf{Ty} ::= \mathcal{C} & \text{(base type)} \\ \quad \mid \tau_1 \rightarrow \tau_2 & \text{(function type)} \end{array}$$

Terms

$$\begin{array}{ll} t \in \mathbf{Tm} ::= x, y, \dots & \text{(variable)} \\ \quad \mid \lambda x : \tau. t & \text{(abstraction)} \\ \quad \mid t_1 t_2 & \text{(application)} \\ \quad \mid \emptyset & \text{(empty)} \\ \quad \mid \{c\} & \text{(singleton)} \\ \quad \mid t_1 \cup t_2 & \text{(union)} \end{array}$$

Environments

$$\Gamma \in \mathbf{Env} ::= \cdot \quad \mid \quad \Gamma, x : \tau$$

Figure 1.1: λ^U -calculus: syntax

1.2.1 Typing relation

The typing relation of the λ^U -calculus is an extension of the simply types λ -calculus' typing relation.

The empty-set and singleton-set constants are of base type and we can only take the set-union of two terms if they have the same type.

1.2.2 Semantics

In the λ^U -calculus terms are interpreted as sets and types as powersets.

$$\begin{array}{c}
\frac{}{\Gamma, x : \tau \vdash x : \tau} [\text{T-VAR}] \quad \frac{\Gamma, x : \tau_1 \vdash t : \tau_2}{\Gamma \vdash \lambda x : \tau_1. t : \tau_1 \rightarrow \tau_2} [\text{T-ABS}] \\
\\
\frac{\Gamma \vdash t_1 : \tau_1 \rightarrow \tau_2 \quad \Gamma \vdash t_2 : \tau_1}{\Gamma \vdash t_1 t_2 : \tau_2} [\text{T-APP}] \\
\\
\frac{}{\Gamma \vdash \emptyset : \mathcal{C}} [\text{T-EMPTY}] \quad \frac{}{\Gamma \vdash \{c\} : \mathcal{C}} [\text{T-CON}] \\
\\
\frac{\Gamma \vdash t_1 : \tau \quad \Gamma \vdash t_2 : \tau}{\Gamma \vdash t_1 \cup t_2 : \tau} [\text{T-UNION}]
\end{array}$$

Figure 1.2: λ^\cup -calculus: type system

Lemma 1. *The terms $(\lambda x : \tau. t_1) \cup (\lambda x : \tau. t_2)$ and $\lambda x : \tau. t_1 \cup t_2$ are extensionally equal.*

Proof. We prove that

$$\llbracket ((\lambda x : \tau. t_1) \cup (\lambda x : \tau. t_2)) t_3 \rrbracket_\rho = \llbracket (\lambda x : \tau. t_1 \cup t_2) t_3 \rrbracket_\rho$$

for all ρ and t_3 .

$$\begin{aligned}
& \llbracket ((\lambda x : \tau. t_1) \cup (\lambda x : \tau. t_2)) t_3 \rrbracket_\rho \\
&= \bigcup \{ \varphi(\llbracket t_3 \rrbracket_\rho) \mid \varphi \in \llbracket (\lambda x : \tau. t_1) \cup (\lambda x : \tau. t_2) \rrbracket_\rho \} \\
&= \bigcup \{ \varphi(\llbracket t_3 \rrbracket_\rho) \mid \varphi \in \llbracket \lambda x : \tau. t_1 \rrbracket_\rho \cup \llbracket \lambda x : \tau. t_2 \rrbracket_\rho \} \\
&= \bigcup \left\{ \varphi(\llbracket t_3 \rrbracket_\rho) \mid \varphi \in \left\{ \lambda v \in V_\tau. \llbracket t_i \rrbracket_{\rho[x \mapsto v]} \mid i \in \{1, 2\} \right\} \right\} \\
&= \bigcup \left\{ \llbracket t_1 \rrbracket_{\rho[x \mapsto \llbracket t_3 \rrbracket_\rho]}, \llbracket t_2 \rrbracket_{\rho[x \mapsto \llbracket t_3 \rrbracket_\rho]} \right\} \\
&= \llbracket t_1 \rrbracket_{\rho[x \mapsto \llbracket t_3 \rrbracket_\rho]} \cup \llbracket t_2 \rrbracket_{\rho[x \mapsto \llbracket t_3 \rrbracket_\rho]} \\
&= \bigcup \left\{ \llbracket t_1 \rrbracket_{\rho[x \mapsto \llbracket t_3 \rrbracket_\rho]} \cup \llbracket t_2 \rrbracket_{\rho[x \mapsto \llbracket t_3 \rrbracket_\rho]} \right\} \\
&= \bigcup \left\{ \llbracket t_1 \cup t_2 \rrbracket_{\rho[x \mapsto \llbracket t_3 \rrbracket_\rho]} \right\} \\
&= \bigcup \left\{ \varphi(\llbracket t_3 \rrbracket_\rho) \mid \varphi \in \left\{ \lambda v \in V_\tau. \llbracket t_1 \cup t_2 \rrbracket_{\rho[x \mapsto v]} \right\} \right\}
\end{aligned}$$

Types and values

$$V_{\mathcal{C}} = \mathcal{P}(\mathbf{Con})$$

$$V_{\tau_1 \rightarrow \tau_2} = \mathcal{P}(V_{\tau_1} \rightarrow V_{\tau_2})$$

Environments

$$\rho : \mathbf{Var} \rightarrow \bigcup \{V_{\tau} \mid \tau \text{ type}\}$$

Terms

$$\begin{aligned} \llbracket x \rrbracket_{\rho} &= \rho(x) \\ \llbracket \lambda x : \tau. t \rrbracket_{\rho} &= \left\{ \lambda v \in V_{\tau}. \llbracket t \rrbracket_{\rho[x \mapsto v]} \right\} \\ \llbracket t_1 \ t_2 \rrbracket_{\rho} &= \bigcup \{ \varphi(\llbracket t_2 \rrbracket_{\rho}) \mid \varphi \in \llbracket t_1 \rrbracket_{\rho} \} \\ \llbracket \emptyset \rrbracket_{\rho} &= \emptyset \\ \llbracket \{c\} \rrbracket_{\rho} &= \{c\} \\ \llbracket t_1 \cup t_2 \rrbracket_{\rho} &= \llbracket t_1 \rrbracket_{\rho} \cup \llbracket t_2 \rrbracket_{\rho} \end{aligned}$$

Figure 1.3: λ^{\cup} -calculus: denotational semantics

$$\begin{aligned} &= \bigcup \{ \varphi(\llbracket t_3 \rrbracket_{\rho}) \mid \varphi \in \llbracket \lambda x : \tau. t_1 \cup t_2 \rrbracket_{\rho} \} \\ &= \llbracket (\lambda x : \tau. t_1 \cup t_2) \ t_3 \rrbracket_{\rho} \end{aligned}$$

□

1.2.3 Normalization

- **To do:** We can make union only work on base types (as we not longer *need* to distribute unions over applications)? Then the denotation of the function space would be simpler and might generalize to other structures..

To reduce λ^{\cup} -terms to a normal form we combine the β -reduction rule of the simply typed λ -calculus with rewrite rules that deal with the associativity, commutativity, idempotence and identity (ACI1) properties of set-union

operator.

If a term t is η -long it can be written in the form

$$t = \lambda x_1 \cdots x_n. \{f_1(t_{11}, \dots, t_{1q_1}), \dots, f_p(t_{p1}, \dots, t_{pq_p})\}$$

where f_i can be a free or bound variable, a singleton-set constant, or another η -long term; and q_i is equal to the arity of f_i (for all $1 \leq i \leq p$). The notation $\{f_1(t_{11}, \dots, t_{1q_1}), \dots, f_p(t_{p1}, \dots, t_{pq_p})\}$ is a shorthand for $f_1(t_{11}, \dots, t_{1q_1}) \cup \dots \cup f_p(t_{p1}, \dots, t_{pq_p})$, where we forget the associativity of the set-union operator and any empty-set constants. Note that despite the suggestive notation, this is not a true set, as there may still be duplicate elements $f_i(t_{i1}, \dots, t_{iq_i})$.

A normal form v of a term t can be written as

$$v = \lambda x_1 \cdots x_n. \{k_1(v_{11}, \dots, v_{1q_1}), \dots, k_p(v_{p1}, \dots, v_{pq_p})\}$$

where k_i can be a free or bound variable, or a singleton-set constant, but not a term as this would form a β -redex.¹ For each k_i, k_j with $i < j$ we must also have that $k_i < k_j$ for some total order on $\mathbf{Var} \cup \mathbf{Con}$. Not only does this imply that each term $k_i(v_{i1}, \dots, v_{iq_i})$ occurs only once in $k_1(v_{11}, \dots, v_{1q_1}), \dots, k_p(v_{p1}, \dots, v_{pq_p})$, but also the stronger condition that $k_i \neq k_j$ for all $i \neq j$.

1.2.4 Pattern unification

Definition 1. A λ^\cup -term π is called a *pattern* if it is of the form $f(e_1, \dots, e_n)$ where f is a free variable and e_1, \dots, e_n are distinct bound variables.

Note that this definition is a special case of what is normally called a *pattern* in higher-order unification theory Miller (1991); Dowek (2001).

If $f(e_1, \dots, e_n)$ is a λ^\cup -pattern and t a λ^\cup -term, then the equation

$$f : \tau_1 \Rightarrow \dots \Rightarrow \tau_n \Rightarrow \tau \vdash \forall e_1 : \tau_1, \dots, e_n : \tau_n. f(e_1, \dots, e_n) = t$$

can be uniquely solved by the unifier

$$\theta = [f \mapsto \lambda e_1 : \tau_1, \dots, e_n : \tau_n. t].$$

¹Technically, terms that bind at least one variable would form a β -redex. Terms that do not bind any variables do not occur either as they merely form a subsequence of $k_1(v_{11}, \dots, v_{1q_1}), \dots, k_p(v_{p1}, \dots, v_{pq_p})$ in this notation.

```

-- normalization of terms
 $\llbracket \cdot \rrbracket :: \mathbf{Tm} \rightarrow \mathbf{Nf}$ 
 $\llbracket \lambda x_1 \cdots x_n. T \rrbracket =$ 
 $\lambda x_1 \cdots x_n. \{ \llbracket f_i(\llbracket t_{i1} \rrbracket, \dots, \llbracket t_{iq_i} \rrbracket) \rrbracket \mid f_i(t_{i1}, \dots, t_{iq_i}) \in T \}$ 

--  $\beta$ -reduction
 $\llbracket k(v_1, \dots, v_q) \rrbracket$ 
 $= k(v_1, \dots, v_q)$ 
 $\llbracket (\lambda y_1 \cdots y_q. T) (v_1, \dots, v_q) \rrbracket$ 
 $= SUBST\ x\ y\ z$ 

-- set-rewriting
 $\{ \{ \dots, k.i(\dots), \dots, k.j(\dots), \dots \} \}$ 
 $\mid k.j < k.i = \{ \{ \dots, k.i(\dots), \dots, k.j(\dots), \dots \} \}$ 
 $\{ \{ \dots, k(\dots), k(\dots), \dots \} \}$ 
 $= \{ \{ \dots, k(\dots), \dots \} \}$ 
 $\{ T \}$ 
 $= T$ 

```

Figure 1.4: Normalization algorithm for λ^U -terms.

1.3 Source language

To do.no catching

Our analysis is applicable to a simple non-strict functional language that supports Boolean, integer and list data types. In section we'll give its syntax and semantics.

- **To do:** Remove type annotation from c_τ , make the operator a boolean one, and drop support for integers?

1.3.1 Underlying type system

1.3.2 Operational semantics

- The reduction relation is non-deterministic.
- We do not have a Haskell-style imprecise exception semantics (e.g. E-If).

Terms

$t \in \mathbf{Tm} ::= x$	(term variable)
c_τ	(term constant)
$\lambda x : \tau. t$	(term abstraction)
$t_1 t_2$	(term application)
$t_1 \oplus t_2$	(operator)
if t_1 then t_2 else t_3	(conditional)
$\not\downarrow_\tau^\ell$	(exception constant)
t_1 seq t_2	(forcing)
fix $x : \tau. t$	(fixpoint)
$[]_\tau$	(nil constructor)
$t_1 :: t_2$	(cons constructor)
case t_1 of $\{ [] \mapsto t_2; x_1 :: x_2 \mapsto t_3 \}$	(list eliminator)

Figure 1.5: Source language: syntax

- We either need to omit the type annotations on $\not\downarrow_\tau^\ell$, or add them to **if then else** and **case of** $\{ [] \mapsto; :: \mapsto \}$.
- We do not have a rule E-ANNAPPEXN. Check that the canonical forms lemma gives us that terms of universally quantified type cannot be exceptional values.

1.4 Exception types**1.4.1 Exception types**

- **To do.** $e \in \mathbf{ExnVar}$
- **To do.** Well-formedness of exception types: embed conservativity / full-flexibility?

- **To do.** Can we roll UNIV and ARR into a single construct: $\forall e : \kappa. \widehat{\tau}_1 \langle e \rangle \rightarrow \widehat{\tau}_2 \langle \zeta(e) \rangle$? Still need to deal with the well-formedness of $\widehat{\tau}_1 \dots$. Also may need to quantify over more than one variable simultaneously...

The syntax of well-formed exception types are given in Figure 1.8 and Figure 1.9. An exception type $\widehat{\tau}$ is formed out of base types (booleans), compound types (lists), function types and quantification over exception variables.²

For a list with exception type $[\widehat{\tau} \langle \zeta \rangle]$ and effect ζ , the type $\widehat{\tau}$ of the elements in the list is *annotated* with an exception set expression ζ of kind EXN. This expression gives a set of exceptions which may be raised when an element of the list is forced. The effect gives a set of exceptions may be raised when a constructor forming the spine of the list is forced.

For a function with exception type $\widehat{\tau}_1 \langle \zeta_1 \rangle \rightarrow \widehat{\tau}_2 \langle \zeta_2 \rangle$ and effect ζ , the argument of type $\widehat{\tau}_1$ is annotated with an exception set expression ζ_1 that gives set of exceptions that may be raised if the argument is forced by the function. The result of type $\widehat{\tau}_2$ is annotated with an exception set expression ζ_2 that gives the set of exceptions that may be raised when the result of the function is forced. The effect ζ gives the set of exceptions that may be raised if the function closure is forced.

$$\begin{aligned} id &: \forall e. \widehat{\mathbf{bool}} \langle e \rangle \rightarrow \widehat{\mathbf{bool}} \langle e \rangle \ \& \ \emptyset \\ id &= \lambda x. x \end{aligned}$$

$$\perp^E : \forall e. \widehat{\mathbf{bool}} \langle e \rangle \rightarrow \widehat{\mathbf{bool}} \langle \emptyset \rangle \ \& \ \{E\}$$

The exception set expressions ζ and their kinds κ are an instance of the λ^U -calculus, where exception set expressions are terms and kinds are the types. Two exception set expressions are considered equivalent if they are convertible as λ^U -terms, which is to say that they reduce to the same normal form.

- **To do.** Some more complicated examples

²To avoid complicating the presentation would do *not* allow quantification over type variables, i.e. polymorphism in the underlying type system.

1.4.2 Conservativeness

To do: Atomicity: $e_1 \cup e_2 \rightarrow \text{ExnTyList } e_1 \ e_2$ is not useful, because no introspection

Any program that is typeable in the underlying type system should also have an exception type: the exception type system is a conservative extension of the underlying type system. Like type systems for strictness or control flow analysis, and unlike type systems for information flow security or dimensional analysis, we do not want to reject any program that is well-typed in the underlying type system, but merely provide more insight in its behavior.

If we furthermore want our type system to be modular—allowing type checking and inference to work on individual modules instead of whole programs—we cannot make any assumptions about the exception types of the arguments that are applied to any function, as the function may be called from outside the module with an argument that also comes from outside the module and which we cannot know anything about.³

For base and compound types that stand in an argument position their effect and any nested annotations must thus be instantiatable to any arbitrary exception set expression. They must thus be exception set variables that have been universally quantified.

- To do: check all example types against prototype
- To do: properly typeset example types
- To do: Skolemization and explicit existential quantification over unification variables?

Example 1.

$$\text{tail} : \forall e_1 \ e_2. [\widehat{\text{bool}}\langle e_1 \rangle]\langle e_2 \rangle \rightarrow [\widehat{\text{bool}}\langle e_1 \rangle]\langle e_2 \cup \{E\} \rangle \ \& \ \emptyset$$

$$\wedge : \forall e_1. \widehat{\text{bool}}\langle e_1 \rangle \rightarrow (\forall e_2. \widehat{\text{bool}}\langle e_2 \rangle \rightarrow \widehat{\text{bool}}\langle e_1 \cup e_2 \rangle)\langle \emptyset \rangle \ \& \ \emptyset$$

For function types that stand in an argument position (the functional parameters of a higher-order function) the situation is slightly more complicated. For the argument of this function we can inductively assume that this will be

³Holdermans and Hage (2010) call such types *fully flexible*.

a universally quantified exception set variable. The result of this function, however, is some exception set expression that depends on the exception set variables that were quantified over in the argument. We cannot simply introduce a new exception set variable here, but must introduce a Skolem function that depends on each of the universally quantified exception set variables.

Example 2. Consider the higher-order function *apply* that applies its first argument to the second.

$$\begin{aligned}
 \text{apply} & : \forall e_2 :: \text{EXN}. \forall e_3 :: \text{EXN} \Rightarrow \text{EXN}. \\
 & \quad (\forall e_1 :: \text{EXN}. \widehat{\text{bool}}\langle e_1 \rangle \rightarrow \widehat{\text{bool}}\langle e_3 \ e_1 \rangle) \langle e_2 \rangle \rightarrow \\
 & \quad (\forall e_4 :: \text{EXN}. \widehat{\text{bool}}\langle e_4 \rangle \rightarrow \widehat{\text{bool}}\langle e_2 \cup e_3 \ e_4 \rangle) \langle \emptyset \rangle \\
 & \quad \& \ \emptyset \\
 \text{apply} & = \lambda f. \lambda x. f \ x
 \end{aligned}$$

The first, functional, argument of *apply* has exception type $\forall e_1 : \text{EXN}. \widehat{\text{bool}}\langle e_1 \rangle \rightarrow \widehat{\text{bool}}\langle e_3 \ e_1 \rangle$ and effect e_2 . It can be instantiated with any function that accepts an argument annotated with any exception set effect, and produces a result annotated with some exception set effect depending on the exception set effect of the argument; the function closure itself may raise any exception. All functions of underlying type $\text{bool} \rightarrow \text{bool}$ satisfy these constraints, so they do not really constrain us in any way.

As e_1 has been quantified over only the exception set operator e_3 and the effect e_2 are left free. We quantify over them outside the outer function space constructor, allowing them to appear in the annotation $e_2 \cup e_3 \ e_4$ in the result. The exception set operator e_3 is now applied to e_4 , as the application $f \ x$ will instantiate the quantified exception set variable e_1 to e_4 .

1.4.3 Exception type completion

During exception type inference will want to compute the least constraint exception type that erases to a given underlying type.

1.4.4 Declarative exception type system

- In T-Abs and T-AnnAbs, should the term-level term-abstraction also have an explicit effect annotation?
- In T-AnnAbs, might need a side condition stating that e is not free in Δ .

- In T-App, note the double occurrence of ξ when typing t_1 . Is subeffecting sufficient here? Also note that we do *not* expect an exception variable in the left-hand side annotation of the function space constructor.
- In T-AnnApp, note the substitution. We will need a substitution lemma for annotations.
- In T-Fix, the might be some universal quantifiers in our way. Do annotation applications in t take care of this, already? Perhaps we do need to change **fix** t into a binding construct to resolve this? Also, there is some implicit subeffecting going on between the annotations and effect.
- In T-Case, note the use of explicit subeffecting. Can this be done using implicit subeffecting?
- For T-Sub, should we introduce a term-level coercion, as in Dussart–Henglein–Mossin? We now do shape-conformant subtyping, is subeffecting sufficient?
- Do we need additional kinding judgements in some of the rules? Can we merge the kinding judgement with the subtyping and/or -effecting judgement? Kind-preserving substitutions.

1.4.5 Type elaboration system

- In T-App and T-Fix, note that there are substitutions in the premises of the rules. Are these inductive? (Probably, as these premises are not “recursive” ones.)
- For T-Fix: how would a binding fixpoint construct work?

$$\begin{array}{c}
\frac{}{\Gamma, x : \tau \vdash x : \tau} \text{[T-VAR]} \quad \frac{}{\Gamma \vdash c_\tau : \tau} \text{[T-CON]} \quad \frac{}{\Gamma \vdash \not\downarrow_\tau^\ell : \tau} \text{[T-CRASH]} \quad \frac{\Gamma, x : \tau_1 \vdash t : \tau_2}{\Gamma \vdash \lambda x : \tau_1. t : \tau_1 \rightarrow \tau_2} \text{[T-ABS]} \\
\\
\frac{\Gamma \vdash t_1 : \tau_2 \rightarrow \tau \quad \Gamma \vdash t_2 : \tau_2}{\Gamma \vdash t_1 t_2 : \tau} \text{[T-APP]} \quad \frac{\Gamma \vdash t : \tau \rightarrow \tau}{\Gamma \vdash \mathbf{fix} \, t : \tau} \text{[T-FIX]} \quad \frac{\Gamma \vdash t_1 : \mathbf{int} \quad \Gamma \vdash t_2 : \mathbf{int}}{\Gamma \vdash t_1 \oplus t_2 : \mathbf{bool}} \text{[T-OP]} \\
\\
\frac{\Gamma \vdash t_1 : \tau_1 \quad \Gamma \vdash t_2 : \tau_2}{\Gamma \vdash t_1 \mathbf{seq} \, t_2 : \tau_2} \text{[T-SEQ]} \quad \frac{\Gamma \vdash t_1 : \mathbf{bool} \quad \Gamma \vdash t_2 : \tau \quad \Gamma \vdash t_3 : \tau}{\Gamma \vdash \mathbf{if} \, t_1 \mathbf{then} \, t_2 \mathbf{else} \, t_3 : \tau} \text{[T-IF]} \\
\\
\frac{}{\Gamma \vdash []_\tau : [\tau]} \text{[T-NIL]} \quad \frac{\Gamma \vdash t_1 : \tau \quad \Gamma \vdash t_2 : [\tau]}{\Gamma \vdash t_1 :: t_2 : [\tau]} \text{[T-CONS]} \quad \frac{\Gamma \vdash t_1 : [\tau_1] \quad \Gamma \vdash t_2 : \tau \quad \Gamma, x_1 : \tau_1, x_2 : [\tau_1] \vdash t_3 : \tau}{\Gamma \vdash \mathbf{case} \, t_1 \mathbf{of} \, \{ [] \mapsto t_2; x_1 :: x_2 \mapsto t_3 \}} \text{[T-CASE]}
\end{array}$$

Figure 1.6: Underlying type system ($\Gamma \vdash t : \tau$)

$$\begin{array}{c}
\frac{t_1 \longrightarrow t'_1}{t_1 \ t_2 \longrightarrow t'_1 \ t_2} \text{[E-APP]} \quad \frac{}{(\lambda x : \widehat{\tau} \ \& \ \xi.t) \ t_2 \longrightarrow t_1[t_2/x]} \text{[E-APPABS]} \quad \frac{t \longrightarrow t'}{t \ \langle \xi \rangle \longrightarrow t' \ \langle \xi \rangle} \text{[E-ANN]} \\
\\
\frac{}{(\Lambda e : \kappa.t) \ \langle \xi \rangle \longrightarrow t[\xi/e]} \text{[E-ANNAbsAbs]} \quad \frac{t \longrightarrow t'}{\mathbf{fix} \ t \longrightarrow \mathbf{fix} \ t'} \text{[E-FIX]} \quad \frac{}{\mathbf{fix} \ (\lambda x : \widehat{\tau} \ \& \ \xi.t) \longrightarrow t[\mathbf{fix}/x]} \text{[E-FIXABS]} \\
\\
\frac{}{\downarrow^\ell t_2 \longrightarrow \downarrow^\ell t_2} \text{[E-APPEXN]} \quad \frac{}{\mathbf{fix} \ \downarrow^\ell \longrightarrow \downarrow^\ell} \text{[E-FIXEXN]} \quad \frac{t_1 \longrightarrow t'_1}{t_1 \oplus t_2 \longrightarrow t'_1 \oplus t_2} \text{[E-OP}_1\text{]} \quad \frac{t_2 \longrightarrow t'_2}{t_1 \oplus t_2 \longrightarrow t_1 \oplus t'_2} \text{[E-OP}_2\text{]} \\
\\
\frac{}{v_1 \oplus v_2 \longrightarrow \llbracket v_1 \oplus v_2 \rrbracket} \text{[E-OP]} \quad \frac{}{\downarrow^\ell \oplus t_2 \longrightarrow \downarrow^\ell} \text{[E-OPEXN}_1\text{]} \quad \frac{}{t_1 \oplus \downarrow^\ell \longrightarrow \downarrow^\ell} \text{[E-OPEXN}_2\text{]} \\
\\
\frac{t_1 \longrightarrow t'_1}{t_1 \ \mathbf{seq} \ t_2 \longrightarrow t'_1 \ \mathbf{seq} \ t_2} \text{[E-SEQ}_1\text{]} \quad \frac{}{v_1 \ \mathbf{seq} \ t_2 \longrightarrow t_2} \text{[E-SEQ}_2\text{]} \quad \frac{}{\downarrow^\ell \ \mathbf{seq} \ t_2 \longrightarrow \downarrow^\ell} \text{[E-SEQEXN]} \\
\\
\frac{t_1 \longrightarrow t'_1}{\mathbf{if} \ t_1 \ \mathbf{then} \ t_2 \ \mathbf{else} \ t_3 \longrightarrow \mathbf{if} \ t'_1 \ \mathbf{then} \ t_2 \ \mathbf{else} \ t_3} \text{[E-IF]} \quad \frac{}{\mathbf{if} \ \mathbf{true} \ \mathbf{then} \ t_2 \ \mathbf{else} \ t_3 \longrightarrow t_2} \text{[E-IFTRUE]} \\
\\
\frac{}{\mathbf{if} \ \mathbf{false} \ \mathbf{then} \ t_2 \ \mathbf{else} \ t_3 \longrightarrow t_3} \text{[E-IFFALSE]} \quad \frac{}{\mathbf{if} \ \downarrow^\ell \ \mathbf{then} \ t_2 \ \mathbf{else} \ t_3 \longrightarrow \downarrow^\ell} \text{[E-IFEXN]} \\
\\
\frac{t_1 \longrightarrow t'_1}{\mathbf{case} \ t_1 \ \mathbf{of} \ \{\square \mapsto t_2; x_1 :: x_2 \mapsto t_3\} \longrightarrow \mathbf{case} \ t'_1 \ \mathbf{of} \ \{\square \mapsto t_2; x_1 :: x_2 \mapsto t_3\}} \text{[E-CASE]} \\
\\
\frac{}{\mathbf{case} \ \square \ \mathbf{of} \ \{\square \mapsto t_2; x_1 :: x_2 \mapsto t_3\} \longrightarrow t_2} \text{[E-CASENIL]} \quad \frac{}{\mathbf{case} \ t_1 :: t'_1 \ \mathbf{of} \ \{\square \mapsto t_2; x_1 :: x_2 \mapsto t_3\} \longrightarrow t_2} \text{[E-CASECONS]} \\
\\
\frac{}{\mathbf{case} \ \downarrow^\ell \ \mathbf{of} \ \{\square \mapsto t_2; x_1 :: x_2 \mapsto t_3\} \longrightarrow \downarrow^\ell} \text{[E-CASEEXN]}
\end{array}$$

Figure 1.7: Operational semantics ($t_1 \longrightarrow t_2$)

$\kappa \in \mathbf{Kind} ::= \mathbf{EXN}$	(exception set)
$\kappa_1 \Rightarrow \kappa_2$	(exception operator)
$\tilde{\zeta}, \zeta \in \mathbf{Exn} ::= e$	(exception variables)
$\lambda e : \kappa. \tilde{\zeta}$	(exception abstraction)
$\tilde{\zeta}_1 \tilde{\zeta}_2$	(exception application)
\emptyset	(empty exception set)
$\{\ell\}$	(singleton exception)
$\tilde{\zeta}_1 \cup \tilde{\zeta}_2$	(exception set union)
$\hat{\tau} \in \mathbf{ExnTy} ::= \forall e :: \kappa. \hat{\tau}$	(exception quantification)
$\widehat{\mathbf{bool}}$	(boolean type)
$[\hat{\tau}(\tilde{\zeta})]$	(list type)
$\hat{\tau}_1 \langle \tilde{\zeta}_1 \rangle \rightarrow \hat{\tau}_2 \langle \tilde{\zeta}_2 \rangle$	(function type)

Figure 1.8: Exception types: syntax

$$\begin{array}{c}
\frac{\Delta, e :: \kappa \vdash \hat{\tau} \text{ wff}}{\Delta \vdash \forall e : \kappa. \hat{\tau} \text{ wff}} [\mathbf{W-UNIV}] \quad \frac{}{\Delta \vdash \widehat{\mathbf{bool}} \text{ wff}} [\mathbf{W-BOOL}] \\
\\
\frac{\Delta \vdash \hat{\tau} \text{ wff} \quad \Delta \vdash \tilde{\zeta} : \mathbf{EXN}}{\Delta \vdash [\hat{\tau}(\tilde{\zeta})] \text{ wff}} [\mathbf{W-LIST}] \\
\\
\frac{\Delta \vdash \hat{\tau}_1 \text{ wff} \quad \Delta \vdash \tilde{\zeta}_1 : \mathbf{EXN} \quad \Delta \vdash \hat{\tau}_2 \text{ wff} \quad \Delta \vdash \tilde{\zeta}_2 : \mathbf{EXN}}{\Delta \vdash \hat{\tau}_1 \langle \tilde{\zeta}_1 \rangle \rightarrow \hat{\tau}_2 \langle \tilde{\zeta}_2 \rangle \text{ wff}} [\mathbf{W-ARR}]
\end{array}$$

Figure 1.9: Exception types: well-formedness

$$\begin{array}{c}
\frac{}{\overline{e_i :: \kappa_i \vdash \mathbf{bool} : \widehat{\mathbf{bool}} \ \& \ e \ \overline{e_i} \triangleright e :: \kappa_i \Rightarrow \text{EXN}}} \text{[C-BOOL]} \quad \frac{\overline{e_i :: \kappa_i \vdash \tau : \widehat{\tau} \ \& \ \xi \triangleright \overline{e_j :: \kappa_j}}}{\overline{e_i :: \kappa_i \vdash [\tau] : [\widehat{\tau}(\xi)] \ \& \ e \ \overline{e_i} \triangleright e :: \kappa_i \Rightarrow \text{EXN}}} \\
\frac{\vdash \tau_1 : \widehat{\tau}_1 \ \& \ \xi_1 \triangleright \overline{e_j :: \kappa_j} \quad \overline{e_i :: \kappa_i}, \overline{e_j :: \kappa_j} \vdash \tau_2 : \widehat{\tau}_2 \ \& \ \xi_2 \triangleright \overline{e_j :: \kappa_j}}{\overline{e_i :: \kappa_i \vdash \tau_1 \rightarrow \tau_2 : \forall \overline{e_j :: \kappa_j}. (\widehat{\tau}_1(\xi_1) \rightarrow \widehat{\tau}_2(\xi_2)) \ \& \ e \ \overline{e_i} \triangleright e :: \kappa_j \Rightarrow \text{EXN}, \overline{e_k :: \kappa_k}}} \text{[C-ARR]}
\end{array}$$

Figure 1.10: Type completion ($\Gamma \vdash \tau : \widehat{\tau} \ \& \ \xi \triangleright \Gamma'$)

$$\begin{array}{c}
\frac{}{\overline{\Gamma, x : \widehat{\tau} \ \& \ \xi; \Delta \vdash x : \widehat{\tau} \ \& \ \xi}} \text{[T-VAR]} \quad \frac{}{\overline{\Gamma; \Delta \vdash c_\tau : \perp_\tau \ \& \ \emptyset}} \text{[T-CON]} \quad \frac{}{\overline{\Gamma; \Delta \vdash \not\leq_\tau^\ell : \perp_\tau \ \& \ \{\ell\}}} \text{[T-CR]} \\
\frac{\Gamma, x : \widehat{\tau}_1 \ \& \ \xi_1; \Delta \vdash t : \widehat{\tau}_2 \ \& \ \xi_2}{\overline{\Gamma; \Delta \vdash \lambda x : \widehat{\tau}_1 \ \& \ \xi_1. t : \widehat{\tau}_1(\xi_1) \rightarrow \widehat{\tau}_2(\xi_2) \ \& \ \emptyset}} \text{[T-ABS]} \quad \frac{\Gamma; \Delta, e : \kappa \vdash t : \widehat{\tau} \ \& \ \xi \quad e \notin \text{fv}(\xi)}{\overline{\Gamma; \Delta \vdash \Lambda e : \kappa. t : \forall e : \kappa. \widehat{\tau} \ \& \ \xi}} \text{[T-ANN]} \\
\frac{\Gamma; \Delta \vdash t_1 : \widehat{\tau}_1(\xi_1) \rightarrow \widehat{\tau}(\xi) \ \& \ \xi \quad \Gamma; \Delta \vdash t_2 : \widehat{\tau}_2 \ \& \ \xi_2}{\overline{\Gamma; \Delta \vdash t_1 \ t_2 : \widehat{\tau} \ \& \ \xi}} \text{[T-APP]} \quad \frac{\Gamma; \Delta \vdash t_1 : \forall e : \kappa. \widehat{\tau} \ \& \ \xi \quad \Delta \vdash \xi_2 : t}{\overline{\Gamma; \Delta \vdash t_1 \ \langle \xi_2 \rangle : \widehat{\tau}[\xi_2/e] \ \& \ \xi}} \\
\frac{\Delta \vdash \xi' \leq \xi \quad \Delta \vdash \xi'' \leq \xi \quad \Gamma; \Delta \vdash t : \widehat{\tau}(\xi') \rightarrow \widehat{\tau}(\xi'') \ \& \ \xi''}{\overline{\Gamma; \Delta \vdash \mathbf{fix} \ t : \widehat{\tau} \ \& \ \xi}} \text{[T-FIX]} \\
\frac{\Gamma; \Delta \vdash t_1 : \widehat{\mathbf{int}} \ \& \ \xi \quad \Gamma; \Delta \vdash t_2 : \widehat{\mathbf{int}} \ \& \ \xi}{\overline{\Gamma; \Delta \vdash t_1 \oplus t_2 : \widehat{\mathbf{bool}} \ \& \ \xi}} \text{[T-OP]} \quad \frac{\Gamma; \Delta \vdash t_1 : \widehat{\tau}_1 \ \& \ \xi \quad \Gamma; \Delta \vdash t_2 : \widehat{\tau}_2 \ \& \ \xi}{\overline{\Gamma; \Delta \vdash t_1 \ \mathbf{seq} \ t_2 : \widehat{\tau}_2 \ \& \ \xi}} \text{[T-SEQ]} \\
\frac{\Gamma; \Delta \vdash t_1 : \widehat{\mathbf{bool}} \ \& \ \xi \quad \Gamma; \Delta \vdash t_2 : \widehat{\tau} \ \& \ \xi \quad \Gamma; \Delta \vdash t_3 : \widehat{\tau} \ \& \ \xi}{\overline{\Gamma; \Delta \vdash \mathbf{if} \ t_1 \ \mathbf{then} \ t_2 \ \mathbf{else} \ t_3 : \widehat{\tau} \ \& \ \xi}} \text{[T-IF]} \\
\frac{}{\overline{\Gamma; \Delta \vdash []_\tau : [\perp_\tau(\emptyset)] \ \& \ \emptyset}} \text{[T-NIL]} \quad \frac{\Gamma; \Delta \vdash t_1 : \widehat{\tau} \ \& \ \xi_1 \quad \Gamma; \Delta \vdash t_2 : [\widehat{\tau}(\xi_1)] \ \& \ \xi_2}{\overline{\Gamma; \Delta \vdash t_1 :: t_2 : [\widehat{\tau}(\xi_1)] \ \& \ \xi_2}} \text{[T-CONS]} \\
\frac{\Gamma; \Delta \vdash t_1 : [\widehat{\tau}_1(\xi_1)] \ \& \ \xi' \quad \Gamma; \Delta \vdash t_2 : \widehat{\tau} \ \& \ \xi \quad \Gamma, x_1 : \widehat{\tau}_1 \ \& \ \xi_1, x_2 : [\widehat{\tau}_1(\xi_1)] \ \& \ \xi'; \Delta \vdash t_3 : \widehat{\tau} \ \& \ \xi}{\overline{\Gamma; \Delta \vdash \mathbf{case} \ t_1 \ \mathbf{of} \ \{ [] \mapsto t_2; x_1 :: x_2 \mapsto t_3 \} : \widehat{\tau} \ \& \ \xi}} \\
\frac{\Gamma; \Delta \vdash t : \widehat{\tau}' \ \& \ \xi' \quad \Delta \vdash \widehat{\tau}' \leq \widehat{\tau} \quad \Delta \vdash \xi' \leq \xi}{\overline{\Gamma; \Delta \vdash t : \widehat{\tau} \ \& \ \xi}} \text{[T-SUB]}
\end{array}$$

Figure 1.11: Declarative type system ($\Gamma; \Delta \vdash t : \widehat{\tau} \ \& \ \xi$)

$$\begin{array}{c}
\frac{}{\Gamma; x : \widehat{\tau} \& \xi; \Delta \vdash x \hookrightarrow x : \widehat{\tau} \& \xi} \text{[T-VAR]} \quad \frac{}{\Gamma; \Delta \vdash c_\tau \hookrightarrow c_\tau : \tau \& \emptyset} \text{[T-CON]} \quad \frac{}{\Gamma; \Delta \vdash \not\downarrow_\tau^\ell \hookrightarrow \not\downarrow_\tau^\ell : \perp_\tau \& \emptyset} \\
\\
\frac{\Delta, \overline{e_i} : \overline{\kappa_i} \vdash \widehat{\tau}_1 \triangleright \tau_1 \quad \Delta, \overline{e_i} : \overline{\kappa_i} \vdash \xi_1 : \text{EXN} \quad \Gamma, x : \widehat{\tau}_1 \& \xi_1; \Delta, \overline{e_i} : \overline{\kappa_i} \vdash t \hookrightarrow t' : \widehat{\tau}_2 \& \xi_2}{\Gamma; \Delta \vdash \lambda x : \tau_1. t \hookrightarrow \Lambda \overline{e_i} : \overline{\kappa_i}. \lambda x : \widehat{\tau}_1 \& \xi_1. t' : \forall \overline{e_i} : \overline{\kappa_i}. \widehat{\tau}_1 \langle \xi_1 \rangle \rightarrow \widehat{\tau}_2 \langle \xi_2 \rangle \& \emptyset} \text{[T-ABS]} \\
\\
\frac{\Delta \vdash \widehat{\tau}_2 \leq \widehat{\tau}[\overline{\xi_i}/\overline{e_i}] \quad \Delta \vdash \xi_2 \leq \xi[\overline{\xi_i}/\overline{e_i}] \quad \overline{\Delta} \vdash \xi_i : \kappa_i}{\Gamma; \Delta \vdash t_1 \hookrightarrow t'_1 : \forall \overline{e_i} : \overline{\kappa_i}. \widehat{\tau}_1 \langle \xi_1 \rangle \rightarrow \widehat{\tau} \langle \xi \rangle \& \xi' \quad \Gamma; \Delta \vdash t_2 \hookrightarrow t'_2 : \widehat{\tau}_2 \& \xi_2} \text{[T-APP]} \\
\Gamma; \Delta \vdash t_1 t_2 \hookrightarrow t'_1 \langle \overline{\xi_i} \rangle t'_2 : \widehat{\tau}[\overline{\xi_i}/\overline{e_i}] \& \xi[\overline{\xi_i}/\overline{e_i}] \cup \xi' \\
\\
\frac{\Gamma; \Delta \vdash t \hookrightarrow t' : \forall \overline{e_i} : \overline{\kappa_i}. \widehat{\tau} \langle \xi \rangle \rightarrow \widehat{\tau}' \langle \xi' \rangle \& \xi'' \quad \Delta \vdash \widehat{\tau}'[\overline{\xi_i}/\overline{e_i}] \leq \widehat{\tau}[\overline{\xi_i}/\overline{e_i}] \quad \Delta \vdash \xi'[\overline{\xi_i}/\overline{e_i}] \leq \xi[\overline{\xi_i}/\overline{e_i}] \quad \overline{\Delta} \vdash \xi_i : \kappa_i}{\Gamma; \Delta \vdash \mathbf{fix} \, t \hookrightarrow \mathbf{fix} \, t' : \widehat{\tau}[\overline{\xi_i}/\overline{e_i}] \& \xi[\overline{\xi_i}/\overline{e_i}] \cup \xi''} \\
\\
\frac{\Gamma; \Delta \vdash t_1 \hookrightarrow t'_1 : \mathbf{int} \& \xi_1 \quad \Gamma; \Delta \vdash t_2 \hookrightarrow t'_2 : \mathbf{int} \& \xi_2}{\Gamma; \Delta \vdash t_1 \oplus t_2 \hookrightarrow t'_1 \oplus t'_2 : \mathbf{bool} \& \xi_1 \cup \xi_2} \text{[T-OP]} \quad \frac{\Gamma; \Delta \vdash t_1 \hookrightarrow t'_1 : \widehat{\tau}_1 \& \xi_1 \quad \Gamma; \Delta \vdash t_2 \hookrightarrow t'_2 : \widehat{\tau}_2 \& \xi_2}{\Gamma; \Delta \vdash t_1 \mathbf{seq} \, t_2 \hookrightarrow t'_1 \mathbf{seq} \, t'_2 : \widehat{\tau}_1 \& \xi_1 \cup \widehat{\tau}_2 \& \xi_2} \\
\\
\frac{\Gamma; \Delta \vdash t_1 \hookrightarrow t'_1 : \mathbf{bool} \& \xi_1 \quad \Gamma; \Delta \vdash t_2 \hookrightarrow t'_2 : \widehat{\tau}_2 \& \xi_2 \quad \Gamma; \Delta \vdash t_3 \hookrightarrow t'_3 : \widehat{\tau}_3 \& \xi_3}{\Gamma; \Delta \vdash \mathbf{if} \, t_1 \mathbf{then} \, t_2 \mathbf{else} \, t_3 \hookrightarrow \mathbf{if} \, t'_1 \mathbf{then} \, t'_2 \mathbf{else} \, t'_3 : \widehat{\tau}_2 \sqcup \widehat{\tau}_3 \& \xi_1 \cup \xi_2 \cup \xi_3} \text{[T-IF]} \\
\\
\frac{}{\Gamma; \Delta \vdash []_\tau \hookrightarrow []_\tau : \perp_\tau \& \emptyset} \text{[T-NIL]} \quad \frac{\Gamma; \Delta \vdash t_1 \hookrightarrow t'_1 : \widehat{\tau}_1 \& \xi_1 \quad \Gamma; \Delta \vdash t_2 \hookrightarrow t'_2 : [\widehat{\tau}'_1 \langle \xi'_1 \rangle] \& \xi_2}{\Gamma; \Delta \vdash t_1 :: t_2 \hookrightarrow t'_1 :: t'_2 : [\widehat{\tau}_1 \sqcup \widehat{\tau}'_1 \langle \xi_1 \cup \xi'_1 \rangle] \& \xi_2} \text{[T-CON]} \\
\\
\frac{\Gamma; \Delta \vdash t_1 \hookrightarrow t'_1 : [\tau_1 \langle \xi_1 \rangle] \& \xi'_1 \quad \Gamma; \Delta \vdash t_2 \hookrightarrow t'_2 : \widehat{\tau}_2 \& \xi_2 \quad \Gamma, x_1 : \widehat{\tau}_1 \& \xi_1, x_2 : [\tau_1 \langle \xi_1 \rangle] \& \xi'_1; \Delta \vdash t_3 \hookrightarrow t'_3 : \widehat{\tau}_2 \sqcup \widehat{\tau}_3 \& \xi'_1 \cup \xi_2}{\Gamma; \Delta \vdash \mathbf{case} \, t_1 \mathbf{of} \{ [] \mapsto t_2; x_1 :: x_2 \mapsto t_3 \} \hookrightarrow \mathbf{case} \, t'_1 \mathbf{of} \{ [] \mapsto t'_2; x_1 :: x_2 \mapsto t'_3 \} : \widehat{\tau}_2 \sqcup \widehat{\tau}_3 \& \xi'_1 \cup \xi_2}
\end{array}$$

Figure 1.12: Syntax-directed type elaboration system ($\Gamma; \Delta \vdash t \hookrightarrow t' : \widehat{\tau} \& \xi$)

1.5 Type inference algorithm

- In R-App and R-Fix: check that the fresh variables generated by \mathcal{I} are substituted away by the substitution θ created by \mathcal{M} . Also, we don't need those variables in the algorithm if we don't generate the elaborated term.
- In R-Fix we could get rid of the auxiliary underlying type function if the fixpoint construct was replaced with a binding variant with an explicit type annotation.
- For R-Fix, make sure the way we handle fixpoints of exceptional value in a manner that is sound w.r.t. to the operational semantics we are going to give to this.
- Note that we do not construct the elaborated term, as it is not useful other than for metatheoretic purposes.
- Lemma: The algorithm maintains the invariant that exception types and exceptions are in normal form.

1.5.1 Subtyping

- Is S-REFL an admissible/derivable rule, or should we drop S-BOOL and S-INT?
- Possibly useful lemma: $\hat{\tau}_1 = \hat{\tau}_2 \iff \hat{\tau}_1 \leq \hat{\tau}_2 \wedge \hat{\tau}_2 \leq \hat{\tau}_1$.

1.6 Interesting observations

- Exception types are not invariant under η -reduction.

1.7 Metatheory

1.7.1 Declarative type system

Lemma 2 (Canonical forms).

1. If \widehat{v} is a possibly exceptional value of type $\widehat{\mathbf{bool}}$, then \widehat{v} is either **true**, **false**, or \downarrow^ℓ .
2. If \widehat{v} is a possibly exceptional value of type $\widehat{\mathbf{int}}$, then \widehat{v} is either some integer n , or an exceptional value \downarrow^ℓ .
3. If \widehat{v} is a possibly exceptional value of type $[\widehat{\tau}\langle\zeta\rangle]$, then \widehat{v} is either $[], t :: t'$, or \downarrow^ℓ .
4. If \widehat{v} is a possibly exceptional value of type $\widehat{\tau}_1\langle\zeta_1\rangle \rightarrow \widehat{\tau}_2\langle\zeta_2\rangle$, then \widehat{v} is either $\lambda x : \widehat{\tau}_1 \ \& \ \zeta_1.t'$ or \downarrow^ℓ .
5. If \widehat{v} is a possibly exceptional value of type $\forall e : \kappa.\widehat{\tau}$, then \widehat{v} is $\Lambda e : \kappa.t$

Proof. For each part, inspect all forms of \widehat{v} and discard the unwanted cases by inversion of the typing relation. Note that \perp_τ cannot give us a type of the form $\forall e : \kappa.\widehat{\tau}$. \square

To do.: Say something about T-SUB?

Theorem 1 (Progress). *If $\Gamma; \Delta \vdash t : \widehat{\tau} \ \& \ \zeta$ with t a closed term, then t is either a possibly exceptional value \widehat{v} or there is a closed term t' such that $t \longrightarrow t'$.*

Proof. By induction on the typing derivation $\Gamma; \Delta \vdash t : \widehat{\tau} \ \& \ \zeta$.

The case T-VAR can be discarded, as a variable is not a closed term. The cases T-CON, T-CRASH, T-ABS, T-ANNAbs, T-NIL and T-CONS are immediate as they are values.

Case T-APP: We can immediately apply the induction hypothesis to $\Gamma; \Delta \vdash t_1 : \widehat{\tau}_1\langle\zeta_2\rangle \rightarrow \widehat{\tau}\langle\zeta\rangle \ \& \ \zeta$, giving us either a t'_1 such that $t_1 \longrightarrow t'_1$ or that $t_1 = \widehat{v}$. In the former case we can make progress using E-APP. In the latter case the canonical forms lemma tells us that either $t_1 = \lambda x : \widehat{\tau}_1 \ \& \ \zeta_2.t'_1$ or $t_1 = \downarrow^\ell$, in which case we can make progress using E-APPABS or E-APPEXN, respectively.

The remaining cases follow by analogous reasoning. \square

Lemma 3 (Annotation substitution).

1. If $\Delta, e : \kappa' \vdash \zeta : \kappa$ and $\Delta \vdash \zeta' : \kappa'$ then $\Delta \vdash \zeta[\zeta'/e] : \kappa$.
2. If $\Delta, e : \kappa' \vdash \zeta_1 \leq \zeta_2$ and $\Delta \vdash \zeta' : \kappa'$ then $\Delta \vdash \zeta_1[\zeta'/e] \leq \zeta_2[\zeta'/e]$.
3. If $\Delta, e : \kappa' \vdash \widehat{\tau}_1 \leq \widehat{\tau}_2$ and $\Delta \vdash \zeta' : \kappa'$ then $\Delta \vdash \widehat{\tau}_1[\zeta'/e] \leq \widehat{\tau}_2[\zeta'/e]$.

4. If $\Gamma; \Delta, e : \kappa' \vdash t : \hat{\tau} \& \zeta$ and $\Delta \vdash \zeta' : \kappa'$ then $\Gamma; \Delta \vdash t[\zeta'/e] : \hat{\tau}[\zeta'/e] \& \zeta$.

To do.: In part 4, either we need the assumption $e \notin \text{fv}(\zeta)$ (which seems to be satisfied everywhere we want to apply this lemma), or we also need to apply the substitution to ζ (is this expected or not in a type-and-effect system)? T-Fix seems to be the only rule where an exception variable can flow from $\hat{\tau}$ to ζ

Proof. 1. By induction on the derivation of $\Delta, e : \kappa' \vdash \zeta : \kappa$. The cases A-VAR, A-ABS and A-APP are analogous to the respective cases in the proof of term substitution below. In the case A-CON one can strengthen the assumption $\Delta, e : \kappa' \vdash \{\ell\} : \text{EXN}$ to $\Delta \vdash \{\ell\} : \text{EXN}$ as $e \notin \text{fv}(\{\ell\})$, the result is then immediate; similarly for A-EMPTY. The case A-UNION goes analogous to A-APP.

2. **To do.**

3. **To do.**

4. By induction on the derivation of $\Gamma; \Delta, e : \kappa' \vdash t : \hat{\tau} \& \zeta$. Most cases can be discarded by a straightforward application of the induction hypothesis; we show only the interesting case.

Case T-ANNAAPP: **To do.**

To do.

□

Lemma 4 (Term substitution). *If $\Gamma, x : \hat{\tau}' \& \zeta'; \Delta \vdash t : \hat{\tau} \& \zeta$ and $\Gamma; \Delta \vdash t' : \hat{\tau}' \& \zeta'$ then $\Gamma; \Delta \vdash t[t'/x] : \hat{\tau} \& \zeta$.*

Proof. By induction on the derivation of $\Gamma, x : \hat{\tau}' \& \zeta'; \Delta \vdash t : \hat{\tau} \& \zeta$.

Case T-VAR: We either have $t = x$ or $t = x'$ with $x \neq x'$. In the first case we need to show that $\Gamma; \Delta \vdash x[t'/x] : \hat{\tau} \& \zeta$, which by definition of substitution is equal to $\Gamma; \Delta \vdash x : \hat{\tau} \& \zeta$, but this is one of our assumptions. In the second case we need to show that $\Gamma, x' : \hat{\tau}' \& \zeta'; \Delta \vdash x'[t'/x] : \hat{\tau} \& \zeta$, which by definition of substitution is equal to $\Gamma, x' : \hat{\tau}' \& \zeta'; \Delta \vdash x' : \hat{\tau} \& \zeta$. This follows immediately from T-VAR.

Case T-ABS: Our assumptions are

$$\Gamma, x : \hat{\tau}' \& \zeta', y : \hat{\tau}_1 \& \zeta_1; \Delta \vdash t : \hat{\tau}_2 \& \zeta_2 \quad (1.1)$$

$$\Gamma; \Delta \vdash t' : \hat{\tau}' \& \zeta'. \quad (1.2)$$

By the Barendregt convention we may assume that $y \neq x$ and $y \notin \text{fv}(t')$. We need to show that $\Gamma; \Delta \vdash (\lambda y : \hat{\tau}_1 \& \zeta_1. t)[t'/x] : \hat{\tau}_1 \langle \zeta_1 \rangle \rightarrow \hat{\tau}_2 \langle \zeta_2 \rangle \& \emptyset$, which by

definition of substitution is equal to

$$\Gamma; \Delta \vdash \lambda y : \hat{\tau}_1 \ \& \ \xi_1.t[t'/x] : \hat{\tau}_1 \langle \xi_1 \rangle \rightarrow \hat{\tau}_2 \langle \xi_2 \rangle \ \& \ \emptyset. \quad (1.3)$$

We weaken (1.2) to $\Gamma, y : \hat{\tau}_1 \ \& \ \xi_1; \Delta \vdash t' : \hat{\tau}' \ \& \ \xi'$ and apply the induction hypothesis on this and (1.1) to obtain

$$\Gamma, y : \hat{\tau}_1 \ \& \ \xi_1; \Delta \vdash t[t'/x] : \hat{\tau}_2 \ \& \ \xi_2. \quad (1.4)$$

The desired result (1.3) can be constructed from (1.4) using T-ABS.

Case T-ANNAbs: Our assumptions are $\Gamma, x : \hat{\tau}' \ \& \ \xi'; \Delta, e : \kappa \vdash t : \hat{\tau} \ \& \ \xi$ and $\Gamma; \Delta \vdash t' : \hat{\tau}' \ \& \ \xi'$. By the Barendregt convention we may assume that $e \notin \text{fv}(t')$. We need to show that $\Gamma; \Delta \vdash (\Lambda e : \kappa.t)[t'/x] : \hat{\tau} \ \& \ \xi$, which is equal to $\Gamma; \Delta \vdash \Lambda e : \kappa.t[t'/\kappa] : \hat{\tau} \ \& \ \xi$ by definition of substitution. By applying the induction hypothesis we obtain $\Gamma; \Delta, e : \kappa \vdash t[t'/x] : \hat{\tau} \ \& \ \xi$. The desired result can be constructed using T-ANNAbs.

Case T-APP: Our assumptions are

$$\Gamma, x : \hat{\tau}' \ \& \ \xi'; \Delta \vdash t_1 : \hat{\tau}_2 \langle \xi_2 \rangle \rightarrow \hat{\tau} \langle \xi \rangle \ \& \ \xi \quad (1.5)$$

$$\Gamma, x : \hat{\tau}' \ \& \ \xi'; \Delta \vdash t_2 : \hat{\tau}_2 \ \& \ \xi_2. \quad (1.6)$$

We need to show that $\Gamma; \Delta \vdash (t_1 \ t_2)[t'/x] : \hat{\tau} \ \& \ \xi$, which by definition of substitution is equal to

$$\Gamma; \Delta \vdash (t_1[t'/x]) \ (t_2[t'/x]) : \hat{\tau} \ \& \ \xi. \quad (1.7)$$

By applying the induction hypothesis to (1.5) respectively (1.6) we obtain

$$\Gamma; \Delta \vdash t_1[t'/x] : \hat{\tau}_2 \langle \xi_2 \rangle \rightarrow \hat{\tau} \langle \xi \rangle \ \& \ \xi \quad (1.8)$$

$$\Gamma; \Delta \vdash t_2[t'/x] : \hat{\tau}_2 \ \& \ \xi_2. \quad (1.9)$$

The desired result (1.7) can be constructed by applying T-APP to (1.8) and (1.9).

All other cases are either immediate or analogous to the case of T-APP. \square

Lemma 5 (Inversion).

1. If $\Gamma; \Delta \vdash \lambda x : \hat{\tau} \ \& \ \xi.t : \hat{\tau}_1 \langle \xi_1 \rangle \rightarrow \hat{\tau}_2 \langle \xi_2 \rangle \ \& \ \xi_3$, then

$$- \Gamma, x : \hat{\tau} \ \& \ \xi; \Delta \vdash t : \hat{\tau}' \ \& \ \xi',$$

- $\Delta \vdash \hat{\tau}_1 \leq \hat{\tau}$ and $\Delta \vdash \xi_1 \leq \xi$,
 - $\Delta \vdash \hat{\tau}' \leq \hat{\tau}_2$ and $\Delta \vdash \xi' \leq \xi_2$.
2. If $\Gamma; \Delta \vdash \Lambda e : \kappa.t : \forall e : \kappa. \hat{\tau} \ \& \ \xi$, then
- $\Gamma; \Delta, e : \kappa \vdash t : \hat{\tau}' \ \& \ \xi'$,
 - $\Delta, e : \kappa \vdash \hat{\tau}' \leq \hat{\tau}$,
 - $\Delta \vdash \xi' \leq \xi$.
 - **To DO.** $e \notin \text{fv}(\xi)$ and/or $e \notin \text{fv}(\xi')$.

Proof. 1. By induction on the typing derivation.

Case T-ABS: We have $\hat{\tau} = \hat{\tau}_1$, $\xi = \xi_1$ and take $\hat{\tau}' = \hat{\tau}_2$, $\xi' = \xi_2$, the result then follows immediately from the assumption $\Gamma, x : \hat{\tau} \ \& \ \xi; \Delta \vdash t : \hat{\tau}_2 \ \& \ \xi_2$ and reflexivity of the subtyping and subeffecting relations.

Case T-SUB: We are given the additional assumptions

$$\Gamma; \Delta \vdash \lambda x : \hat{\tau} \ \& \ \xi. t : \hat{\tau}'_1 \langle \xi'_1 \rangle \rightarrow \hat{\tau}'_2 \langle \xi'_2 \rangle \ \& \ \xi'_3, \quad (1.10)$$

$$\Delta \vdash \hat{\tau}'_1 \langle \xi'_1 \rangle \rightarrow \hat{\tau}'_2 \langle \xi'_2 \rangle \leq \hat{\tau}_1 \langle \xi_1 \rangle \rightarrow \hat{\tau}_2 \langle \xi_2 \rangle, \quad (1.11)$$

$$\Delta \vdash \xi'_3 \leq \xi_3. \quad (1.12)$$

Applying the induction hypothesis to (1.10) gives us

$$\Gamma, x : \hat{\tau} \ \& \ \xi; \Delta \vdash t : \hat{\tau}''_2 \ \& \ \xi''_2, \quad (1.13)$$

$$\Delta \vdash \hat{\tau}'_1 \leq \hat{\tau}, \quad \Delta \vdash \xi'_1 \leq \xi, \quad (1.14)$$

$$\Delta \vdash \hat{\tau}''_2 \leq \hat{\tau}'_2, \quad \Delta \vdash \xi''_2 \leq \xi'_2. \quad (1.15)$$

Inversion of the subtyping relation on (1.11) gives us

$$\Delta \vdash \hat{\tau}'_1 \leq \hat{\tau}, \quad \Delta \vdash \xi'_1 \leq \xi, \quad (1.16)$$

$$\Delta \vdash \hat{\tau}''_2 \leq \hat{\tau}'_2, \quad \Delta \vdash \xi''_2 \leq \xi'_2. \quad (1.17)$$

The result follows from (1.13) and combining (1.16) with (1.14) and (1.15) with (1.17) using the transitivity of the subtyping and subeffecting relations.

2. By induction on the typing derivation.

Case T-ANNAbs: We need to show that $\Gamma; \Delta, e : \kappa \vdash t : \hat{\tau} \ \& \ \xi$, which is one of our assumptions, and that $\Delta, e : \kappa \vdash \hat{\tau} \leq \hat{\tau}$ and $\Delta \vdash \xi \leq \xi$; this follows from the reflexivity of the subtyping, respectively subeffecting, relation (noting that $e \notin \text{fv}(\xi)$).

Case T-SUB: Similar to the case T-SUB in part 1. □

Theorem 2 (Preservation). *If $\Gamma; \Delta \vdash t : \hat{\tau} \ \& \ \check{\zeta}$ and $t \longrightarrow t'$, then $\Gamma; \Delta \vdash t' : \hat{\tau} \ \& \ \check{\zeta}$.*

Proof. By induction on the typing derivation $\Gamma; \Delta \vdash t : \hat{\tau} \ \& \ \check{\zeta}$.

The cases for T-VAR, T-CON, T-CRASH, T-ABS, T-ANNABS, T-NIL, and T-CONS can be discarded immediately, as they have no applicable evaluation rules.

To do. □

1.7.2 Syntax-directed type elaboration

1.7.3 Type inference algorithm

Theorem 3 (Syntactic soundness). *If $\mathcal{R} \ \Gamma \ \Delta \ t = \langle \hat{\tau}; \check{\zeta} \rangle$, then $\Gamma; \Delta \vdash t : \hat{\tau} \ \& \ \check{\zeta}$.*

Proof. By induction on the term t .

To do. □

Theorem 4 (Termination). *$\mathcal{R} \ \Gamma \ \Delta \ t$ terminates.*

Proof. By induction on the term t .

To do. □

1.8 Further research

At least two opportunities for further research present itself:

1. Can we adapt our type inference algorithm to automatically infer types for Kennedy's higher-ranked $\Lambda_{\delta\omega}$? One problem that immediately presents itself is that this type system is not a conservative extension of the underlying type system: programs can be rejected because, while type correct in the underlying type system, the program may still be dimensionally inconsistent.

Unlike in this system, the annotations on function arguments will no longer be of the simple form (variables only) required for the straight-forward matching step in our type inference algorithm. Instead we will likely have to solve a higher-order (equational) unification problem, which is only semi-decidable. Snyder (1990) and Nipkow and Qian

(1991) do gives us semi-algorithms for solving such problems (although, at least in the latter approach, the equational theory is assumed to be regular, which the theory of Abelian groups is not).

2. Can we further improve the precision of exception types? In previous work Koot and Hage (2015) we argued that an accurate exception typing system for non-strict languages should also take the data flow of the program into account, as many exceptions that can be raised in non-strict languages are caused by incomplete pattern matches. The canonical example is the *risers* function—which splits a list into monotonically increasing subsegments; e.g., *risers* $[1, 3, 5, 1, 2] \rightarrow [[1, 3, 5], [1, 2]]$ —by Mitchell and Runciman (2008):

```

risers : [int] → [[int]]
risers [] = []
risers [x] = [[x]]
risers (x1 :: x2 :: xs) =
  if x1 ≤ x2 then (x1 :: y) :: ys else [x1] :: (y :: ys)
  where (y :: ys) = risers (x2 :: xs)

```

Our type inference algorithm assigns *risers* the exception type

$$\forall e_1 :: \text{EXN}. \forall e_2 :: \text{EXN}. \\ [\mathbf{int}\langle e_2 \rangle] \langle e_1 \rangle \rightarrow [[\mathbf{int}\langle e_2 \rangle] \langle \emptyset \rangle] \langle e_1 \cup e_2 \cup \{E\} \rangle \& \emptyset$$

where **E** is the label of the exception raised when the pattern match in the **where**-clause fails.⁴ However, we pattern match on the result of the recursive call *risers* $(x_2 :: xs)$. When *risers* is given a non-empty list (such as $x_2 :: xs$) as an argument, it will always return a non-empty list as its result. The pattern match can thus never fail, and the exception labelled **E** will thus never be raised.

In our previous work we demonstrated how this exception can be elided, by having the exception flow depend on the data flow. The λ^\cup -calculus terms that form our effect annotations cannot express this dependence, however. In this earlier work we used a slightly ad hoc form of conditional constraints to model this dependence. We now believe extending a

⁴This exception is left implicit in the above program, but becomes explicit when the code is desugared into our core language.

λ -calculus with an equational theory of Boolean rings may form the basis of more principled approach. Booleans rings have already been successfully used to design type system for strictness analysis Wright (1991), records Kennedy (1996b) and exception tracking Benton and Buchlovsky (2007).

1.9 Related work

1.9.1 Higher-ranked polymorphism in type-and-effect systems

Effect polymorphism For plain type systems Hindley–Milner’s **let**-bound polymorphism generally provides a good compromise between expressiveness of the type system and complexity of the inference algorithm Hindley (1969); Milner (1978); Damas and Milner (1982). These systems were extended with effects—including **let**-bound effect polymorphism—by Lucassen and Gifford (1988); Jouvelot and Gifford (1991); and Talpin and Jouvelot (1992, 1994). In type-and-effect systems it has long been recognized that **fix**-bound polymorphism (polymorphic recursion) *in the effects* is often beneficial or even necessary for achieving precise analysis results. For example, in type-and-effect systems for regions Tofte and Talpin (1994), dimensions Kennedy (1994); Rittri (1994, 1995), binding-times Dussart et al. (1995), and exceptions Glynn et al. (2002); Koot and Hage (2015). Inferring principal types in a type system with polymorphic recursion is equivalent to solving an undecidable semi-unification problem Mycroft (1984); Kfoury et al. (1990, 1993); Henglein (1993).

When restricted to polymorphic recursion in the effects, the problem often becomes decidable again. In Tofte and Talpin (1994) this is a conjecture based on empirical observation. Rittri (1995) gives a semi-unification procedure based on the general semi-algorithm by Baaz (1993) and proves it terminates in the special case of semi-unification in Abelian groups. Dussart et al. (1995) use a constraint-based algorithm and show that all variables that do not occur free in the context or type can be eliminated from the constraint set by a constraint reduction step during each Kleene–Mycroft iteration. As at most n^2 subeffecting constraints can be formed over n variables, the whole procedure must terminate. By not restarting the Kleene–Mycroft iteration from bottom their algorithm will run in polynomial time, even in the presence of nested fixpoints.

Somehow, the obvious extension to **lambda**-bound effect polymorphism (higher-ranked polymorphism) remained mostly unstudied, possibly because it is of limited use without the simultaneous introduction of effect operators, in contrast to higher-ranked polymorphism in plain type systems.

Effect operators Kennedy (1996a) presents a type system that ensures the dimensional consistency of a ML-like language extended with units of measure (ML_δ). This language has predicative, prenex dimension polymorphism. Kennedy gives an Algorithm \mathcal{W} -like type inference procedure that uses equational unification to deal with the Abelian group (AG) structure of dimension expressions. Also described are two explicitly typed variants of the language: a System F-like language with arbitrary rank dimension polymorphism (Λ_δ), and a System F_ω -like language that extends Λ_δ with dimension operators ($\Lambda_{\delta\omega}$). This language can type strictly more programs than the language without dimension operators:

$$\begin{aligned}
\text{twice} & : \forall F :: \text{DIM} \Rightarrow \text{DIM}. \\
& \quad (\forall d :: \text{DIM}. \mathbf{real}\langle d \rangle \rightarrow \mathbf{real}\langle F\ d \rangle) \rightarrow \\
& \quad (\forall d :: \text{DIM}. \mathbf{real}\langle d \rangle \rightarrow \mathbf{real}\langle F\ (F\ d) \rangle) \\
\text{twice} & = \Lambda F :: \text{DIM} \Rightarrow \text{DIM}. \\
& \quad \lambda f : (\forall d :: \text{DIM}. \mathbf{real}\langle d \rangle \rightarrow \mathbf{real}\langle F\ d \rangle). \\
& \quad \Lambda d :: \text{DIM}. \lambda x : \mathbf{real}\langle d \rangle. f\ \langle F\ d \rangle\ (f\ \langle d \rangle\ x) \\
\text{square} & : \forall d :: \text{DIM}. \mathbf{real}\langle d \rangle \rightarrow \mathbf{real}\langle d^2 \rangle \\
\text{square} & = \Lambda d :: \text{DIM}. \lambda x : \mathbf{real}\langle d \rangle. x^2 \\
\text{fourth} & : \forall d :: \text{DIM}. \mathbf{real}\langle d \rangle \rightarrow \mathbf{real}\langle d^4 \rangle \\
\text{fourth} & = \text{twice}\ \langle \Lambda d :: \text{DIM}. d^2 \rangle\ \text{square} \\
\text{sixteenth} & : \forall d :: \text{DIM}. \mathbf{real}\langle d \rangle \rightarrow \mathbf{real}\langle d^{16} \rangle \\
\text{sixteenth} & = \text{twice}\ \langle \Lambda d :: \text{DIM}. d^4 \rangle\ \text{fourth}
\end{aligned}$$

Without dimension operators we would have to specialize the type of the higher-order function *twice* to one that is either only application in *fourth*, or one that is only applicable in *sixteenth*.

The language $\Lambda_{\delta\omega}$ bears a striking resemblance to our language in Figure 1.8: the empty and singleton exception sets constants and the exception set union operator have been replaced with a unit dimension and dimension product and inverse operators, as dimensions have an AG structure, while

exception sets have an ACI1 structure; in the dimension type system the annotation is placed only on the real number base type instead of on the compound types, and there is no effect. No type inference algorithm is presented for this language, however.

To do: Stefan

To do: RW in Stefan (Faxen? Smith & Wang?)

1.9.2 λ^U -calculus

Tannen (1988) and Tannen and Gallier (1991) prove that if a simply typed λ -calculus is extended with a many-sorted algebraic rewrite system R (by introducing the symbols of the algebraic theory as higher-order constants in the λ -calculus), then the combined rewrite system $\beta\eta R$ is confluent and strongly normalizing if R is confluent and strongly normalizing.

Révész (1992) introduced an untyped λ -calculus with applicative lists. A model is given by Durfee (1997). This calculus satisfies the equations

$$\langle t_1, \dots, t_n \rangle t' = \langle t_1 t', \dots, t_n t' \rangle \quad (\gamma_1)$$

$$\lambda x. \langle t_1, \dots, t_n \rangle = \langle \lambda x. t_1, \dots, \lambda x. t_n \rangle \quad (\gamma_2)$$

similar to our typed λ^U -calculus.

1.9.3 Exception analyses

To do: Update to this paper! Several exception analyses have been described in the literature, primarily targeting the detection of uncaught exceptions in ML. The exception analysis in Yi (1994) is based on abstract interpretation. Guzmán and Suárez (1994) and Fahndrich et al. (1998) describe type-based exception analyses. The row-based type system for exception analysis described in Leroy and Pessaux (2000) containing a data-flow analysis component that is specialized towards tracking value-carrying exceptions instead of value-dependent exceptions and thus employs a less precise unification method instead of subtyping. Glynn et al. (2002) developed the first exception analysis for a non-strict language. It is a type-based analysis using Boolean constraints and, although it does not take data flow into account, has a similar flavor to our system.

To do: Koot and Hage 2015

1.10 Conclusion

We show that it is feasible to extend non-strict higher-order languages with exception-annotated types, as is already done in some strict first-order languages. We argue that the use of higher-ranked exception polymorphism and exception operators *à la* System F_ω , not only increase the precision of the analysis, but are also more readable and understandable for the programmer *vis-à-vis* constrained type schemes, as the exception terms in the the annotations more closely mirror what is happening at the term level than a constraint set would.

$$\begin{aligned}
\mathcal{R} &:: \mathbf{TyEnv} \times \mathbf{KiEnv} \times \mathbf{Tm} \rightarrow \mathbf{ExnTy} \times \mathbf{Exn} \\
\mathcal{R} \Gamma \Delta x &= \Gamma_x \\
\mathcal{R} \Gamma \Delta c_\tau &= \langle \perp_\tau; \emptyset \rangle \\
\mathcal{R} \Gamma \Delta \frac{\ell}{\tau} &= \langle \perp_\tau; \{\ell\} \rangle \\
\mathcal{R} \Gamma \Delta (\lambda x : \tau. t) &= \\
&\quad \mathbf{let} \langle \widehat{\tau}_1; e_1; \overline{e_i : \kappa_i} \rangle = \mathcal{C} \emptyset \tau \\
&\quad \langle \widehat{\tau}_2; \xi_2 \rangle = \mathcal{R} (\Gamma, x : \widehat{\tau}_1 \ \& \ e_1) (\Delta, \overline{e_i : \kappa_i}) t \\
&\quad \mathbf{in} \langle \overline{\forall e_i : \kappa_i. \widehat{\tau}_1} \langle e_1 \rangle \rightarrow \widehat{\tau}_2 \langle \xi_2 \rangle; \emptyset \rangle \\
\mathcal{R} \Gamma \Delta (t_1 \ t_2) &= \\
&\quad \mathbf{let} \langle \widehat{\tau}_1; \xi_1 \rangle = \mathcal{R} \Gamma \Delta t_1 \\
&\quad \langle \widehat{\tau}_2; \xi_2 \rangle = \mathcal{R} \Gamma \Delta t_2 \\
&\quad \langle \widehat{\tau}'_2 \langle e'_2 \rangle \rightarrow \widehat{\tau}' \langle \xi' \rangle; \overline{e_i : \kappa_i} \rangle = \mathcal{I} \widehat{\tau}_1 \\
&\quad \theta = [e'_2 \mapsto \xi_2] \circ \mathcal{M} \emptyset \widehat{\tau}_2 \widehat{\tau}'_2 \\
&\quad \mathbf{in} \langle \llbracket \theta \widehat{\tau}' \rrbracket_\Delta; \llbracket \theta \xi' \cup \xi_1 \rrbracket_\Delta \rangle \\
\mathcal{R} \Gamma \Delta (\mathbf{fix} \ t) &= \\
&\quad \mathbf{let} \langle \widehat{\tau}; \xi \rangle = \mathcal{R} \Gamma \Delta t \\
&\quad \langle \widehat{\tau}' \langle e' \rangle \rightarrow \widehat{\tau}'' \langle \xi'' \rangle; \overline{e_i : \kappa_i} \rangle = \mathcal{I} \widehat{\tau} \\
&\quad \mathbf{in} \langle \widehat{\tau}_0; \xi_0; i \rangle \leftarrow \langle \perp_{[\widehat{\tau}']}; \emptyset; 0 \rangle \\
&\quad \mathbf{do} \ \theta \leftarrow [e' \mapsto \xi_i] \circ \mathcal{M} \emptyset \widehat{\tau}_i \widehat{\tau}' \\
&\quad \langle \widehat{\tau}_{i+1}; \xi_{i+1}; i \rangle \leftarrow \langle \llbracket \theta \widehat{\tau}'' \rrbracket_\Delta; \llbracket \theta \xi'' \rrbracket_\Delta; i+1 \rangle \\
&\quad \mathbf{until} \langle \widehat{\tau}_i; \xi_i \rangle \equiv \langle \widehat{\tau}_{i-1}; \xi_{i-1} \rangle \\
&\quad \mathbf{return} \langle \widehat{\tau}_i; \llbracket \xi \cup \xi_i \rrbracket_\Delta \rangle \\
\mathcal{R} \Gamma \Delta (t_1 \oplus t_2) &= \\
&\quad \mathbf{let} \langle \widehat{\mathbf{int}}; \xi_1 \rangle = \mathcal{R} \Gamma \Delta t_1 \\
&\quad \langle \widehat{\mathbf{int}}; \xi_2 \rangle = \mathcal{R} \Gamma \Delta t_2 \\
&\quad \mathbf{in} \langle \widehat{\mathbf{bool}}; \llbracket \xi_1 \cup \xi_2 \rrbracket_\Delta \rangle \\
\mathcal{R} \Gamma \Delta (t_1 \ \mathbf{seq} \ t_2) &= \\
&\quad \mathbf{let} \langle \widehat{\tau}_1; \xi_1 \rangle = \mathcal{R} \Gamma \Delta t_1 \\
&\quad \langle \widehat{\tau}_2; \xi_2 \rangle = \mathcal{R} \Gamma \Delta t_2 \\
&\quad \mathbf{in} \langle \widehat{\tau}_2; \llbracket \xi_1 \cup \xi_2 \rrbracket_\Delta \rangle \\
\mathcal{R} \Gamma \Delta (\mathbf{if} \ t_1 \ \mathbf{then} \ t_2 \ \mathbf{else} \ t_3) &= \\
&\quad \mathbf{let} \langle \widehat{\mathbf{bool}}; \xi_1 \rangle = \mathcal{R} \Gamma \Delta t_1 \\
&\quad \langle \widehat{\tau}_2; \xi_2 \rangle = \mathcal{R} \Gamma \Delta t_2 \\
&\quad \langle \widehat{\tau}_3; \xi_3 \rangle = \mathcal{R} \Gamma \Delta t_3 \\
&\quad \mathbf{in} \langle \llbracket \widehat{\tau}_2 \sqcup \widehat{\tau}_3 \rrbracket_\Delta; \llbracket \xi_1 \cup \xi_2 \cup \xi_3 \rrbracket_\Delta \rangle \\
\mathcal{R} \Gamma \Delta \llbracket \perp_\tau \rrbracket &= \langle \llbracket \perp_\tau \langle \emptyset \rangle \rrbracket; \emptyset \rangle \\
\mathcal{R} \Gamma \Delta (t_1 :: t_2) &= \\
&\quad \mathbf{let} \langle \widehat{\tau}_1; \xi_1 \rangle = \mathcal{R} \Gamma \Delta t_1 \\
&\quad \langle \llbracket \widehat{\tau}_2 \langle \xi'_2 \rangle \rrbracket; \xi_2 \rangle = \mathcal{R} \Gamma \Delta t_2 \\
&\quad \mathbf{in} \langle \llbracket \llbracket (\widehat{\tau}_1 \sqcup \widehat{\tau}_2) \langle \xi_1 \cup \xi'_2 \rangle \rrbracket_\Delta; \xi_2 \rangle \\
\mathcal{R} \Gamma \Delta (\mathbf{case} \ t_1 \ \mathbf{of} \ \{\llbracket \cdot \rrbracket \mapsto t_2; x_1 :: x_2 \mapsto t_3\}) &= \\
&\quad \mathbf{let} \langle \llbracket \widehat{\tau}_1 \langle \xi'_1 \rangle \rrbracket; \xi_1 \rangle = \mathcal{R} \Gamma \Delta t_1 \\
&\quad \langle \widehat{\tau}_2; \xi_2 \rangle = \mathcal{R} (\Gamma \ x_1 : \widehat{\tau}_1 \ \& \ \xi'_1 \ x_2 : \llbracket \widehat{\tau}_1 \langle \xi'_1 \rangle \rrbracket \ \& \ \xi_1) \ \Delta \ t_2
\end{aligned}$$

$$\begin{array}{c}
\frac{}{\Delta \vdash \widehat{\tau} \leq \widehat{\tau}} [\text{S-REFL}] \quad \frac{\Delta \vdash \widehat{\tau}_1 \leq \widehat{\tau}_2 \quad \Delta \vdash \widehat{\tau}_2 \leq \widehat{\tau}_3}{\Delta \vdash \widehat{\tau}_1 \leq \widehat{\tau}_3} [\text{S-TRANS}] \\
\\
\frac{}{\Delta \vdash \widehat{\mathbf{bool}} \leq \widehat{\mathbf{bool}}} [\text{S-BOOL}] \quad \frac{}{\Delta \vdash \widehat{\mathbf{int}} \leq \widehat{\mathbf{int}}} [\text{S-INT}] \\
\\
\frac{\Delta \vdash \widehat{\tau}'_1 \leq \widehat{\tau}_1 \quad \Delta \vdash \zeta'_1 \leq \zeta_1 \quad \Delta \vdash \widehat{\tau}_2 \leq \widehat{\tau}'_2 \quad \Delta \vdash \zeta_2 \leq \zeta'_2}{\Delta \vdash \widehat{\tau}_1 \langle \zeta_1 \rangle \rightarrow \widehat{\tau}_2 \langle \zeta_2 \rangle \leq \widehat{\tau}'_1 \langle \zeta'_1 \rangle \rightarrow \widehat{\tau}'_2 \langle \zeta'_2 \rangle} [\text{S-ARR}] \\
\\
\frac{\Delta \vdash \widehat{\tau} \leq \widehat{\tau}' \quad \Delta \vdash \zeta \leq \zeta'}{\Delta \vdash [\widehat{\tau} \langle \zeta \rangle] \leq [\widehat{\tau}' \langle \zeta' \rangle]} [\text{S-LIST}] \\
\\
\frac{\Delta, e : \kappa \vdash \widehat{\tau}_1 \leq \widehat{\tau}_2}{\Delta \vdash \forall e : \kappa. \widehat{\tau}_1 \leq \widehat{\tau}_2} [\text{S-FORALL}]
\end{array}$$

Figure 1.14: Subtyping

$$\begin{array}{ll}
e[\zeta/e] \equiv \zeta & \\
e'[\zeta/e] \equiv e' & \text{if } e \neq e' \\
\{\ell\}[\zeta/e] \equiv \{\ell\} & \\
\emptyset[\zeta/e] \equiv \emptyset & \\
(\lambda e' : \kappa. \zeta') [\zeta/e] \equiv \lambda e' : \kappa. \zeta'[\zeta/e] & \text{if } e \neq e' \text{ and } e' \notin \text{fv}(\zeta) \\
(e_1 \ e_2) [\zeta/e] \equiv (e_1[\zeta/e]) \ (e_2[\zeta/e]) & \\
(e_1 \cup e_2) [\zeta/e] \equiv e_1[\zeta/e] \cup e_2[\zeta/e] &
\end{array}$$

Figure 1.15: Annotation substitution

$$\begin{aligned}
x[t/x] &\equiv t \\
x'[t/x] &\equiv x' && \text{if } x \neq x' \\
c_\tau[t/x] &\equiv c_\tau \\
(\lambda x' : \widehat{\tau}.t') [t/x] &\equiv \lambda x' : \widehat{\tau}.t'[t/x] && \text{if } x \neq x' \text{ and } x' \notin \text{fv}(t) \\
&\dots
\end{aligned}$$

Figure 1.16: Term substitution

Appendix A

Old stuff

A.1 λ^{\cup} -calculus

Values Values v are terms of the form

$$\lambda x_1 : \tau_1. \dots \lambda x_i : \tau_i. \{c_1\} \cup (\dots \cup (\{c_j\} \cup (x_1 v_{11} \dots v_{1m} \cup (\dots \cup x_k v_{k1} \dots v_{kn}))))))$$

A.1.1 Reduction relation (wrong!)

- **To do.** Do not match the rules in the prototype (those are sensitive to the order in which they are tried).
- **To do.** In the second rule only one term is applied; contrast this with the other rules involving applications.
- **To do.** Should make use of the fact the everything is fully applied (and η -expanded/-long?): all atoms are of the form $k \bar{t}_i$, where k is c or x and the number of arguments fixed by the arity of k . Then try to factor out the commutativity rules by taking “sets” of these atoms. That might simplify stuff a whole lot...
- **To do.** Can we restrict the typing rule T-Union to only allow sets and not functions on both sides? This would remove the 2nd and 3rd rewrite rules and make the system a more traditional higher-order rewrite system: it’s “just” higher-order pattern E-unification (decidable), boolean

rings are easy to integrate, and higher-ranked dimension types becomes higher-order E-unification (semi-decidable). Open question: how to represent e.g. $U(e_2(e_1), e_1) = [e_2 \mapsto \lambda e_1. e_1]$ without abstractions? (Reinterpret e_1 as $f(e_1)$ with $f = id$?)

Definition 2. Let \prec be a strict total order on $\mathbf{Con} \cup \mathbf{Var}$, with $c \prec x$ for all $c \in \mathbf{Con}$ and $x \in \mathbf{Var}$.

$$\begin{aligned}
& (\lambda x : \tau. t_1) \ t_2 \longrightarrow t_1[t_2/x] && (\beta\text{-reduction}) \\
& (t_1 \cup t_2) \ t_3 \longrightarrow t_1 \ t_3 \cup t_2 \ t_3 \\
& (\lambda x : \tau. t_1) \cup (\lambda x : \tau. t_2) \longrightarrow \lambda x : \tau. (t_1 \cup t_2) && (\text{congruences}) \\
& x \ t_1 \cdots t_n \cup x \ t'_1 \cdots t'_n \longrightarrow x \ (t_1 \cup t'_1) \cdots (t_n \cup t'_n) \\
& (t_1 \cup t_2) \cup t_3 \longrightarrow t_1 \cup (t_2 \cup t_3) && (\text{associativity}) \\
& \emptyset \cup t \longrightarrow t && (\text{unit}) \\
& t \cup \emptyset \longrightarrow t \\
& x \cup x \longrightarrow x \\
& x \cup (x \cup t) \longrightarrow x \cup t && (\text{idempotence}) \\
& \{c\} \cup \{c\} \longrightarrow \{c\} \\
& \{c\} \cup (\{c\} \cup t) \longrightarrow \{c\} \cup t \\
& x \ t_1 \cdots t_n \cup \{c\} \longrightarrow \{c\} \cup x \ t_1 \cdots t_n && (\text{A.1}) \\
& x \ t_1 \cdots t_n \cup (\{c\} \cup t) \longrightarrow \{c\} \cup (x \ t_1 \cdots t_n \cup t) && (\text{A.2}) \\
& x \ t_1 \cdots t_n \cup x \ t'_1 \cdots t'_n \longrightarrow x \ t'_1 \cdots t'_n \cup x \ t_1 \cdots t_n && \text{if } x' \prec x \quad (\text{A.3}) \\
& x \ t_1 \cdots t_n \cup (x \ t'_1 \cdots t'_n \cup t) \longrightarrow x \ t'_1 \cdots t'_n \cup (x \ t_1 \cdots t_n \cup t) && \text{if } x' \prec x \quad (\text{A.4}) \\
& \{c\} \cup \{c'\} \longrightarrow \{c'\} \cup \{c\} && \text{if } c' \prec c \quad (\text{A.5}) \\
& \{c\} \cup (\{c'\} \cup t) \longrightarrow \{c'\} \cup (\{c\} \cup t) && \text{if } c' \prec c \quad (\text{A.6})
\end{aligned}$$

Conjecture 1. The reduction relation \longrightarrow preserves meaning.

Conjecture 2. The reduction relation \longrightarrow is strongly normalizing.

Conjecture 3. The reduction relation \longrightarrow is locally confluent.

Corollary 1. *The reduction relation \longrightarrow is confluent.*

Proof. Follows from SN, LC and Newman's Lemma. \square

Corollary 2. *The λ^\cup -calculus has unique normal forms.*

Corollary 3. *Equality of λ^\cup -terms can be decided by normalization.*

A.1.2 Semantics

- **To do.** Combine the lemma and the theorem and make the “extensionally” explicit.
- **To do.** Is the case for applications rigorous? Relies on the monotonicity of $\varphi : V_{\tau_1} \rightarrow V_{\tau_2}$ (separate lemma, require in the denotational semantics?); this might fail for anything other than set union?

Lemma 6. $\llbracket t \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket t \rrbracket_{\rho[x \mapsto v_2]} \subseteq \llbracket t \rrbracket_{\rho[x \mapsto v_1 \cup v_2]}$ (*extensionally*).

Proof. By induction on the term t .

Case “ $t = x$ ”: $\llbracket x \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket x \rrbracket_{\rho[x \mapsto v_2]} = \rho[x \mapsto v_1](x) \cup \rho[x \mapsto v_2](x) = v_1 \cup v_2 = \rho[x \mapsto v_1 \cup v_2](x) = \llbracket x \rrbracket_{\rho[x \mapsto v_1 \cup v_2]}.$

Case “ $t = y$ ($y \neq x$)”: $\llbracket y \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket y \rrbracket_{\rho[x \mapsto v_2]} = \rho[x \mapsto v_1](y) \cup \rho[x \mapsto v_2](y) = \rho(y) \cup \rho(y) = \rho(y) = \rho[x \mapsto v_1 \cup v_2](y) = \llbracket y \rrbracket_{\rho[x \mapsto v_1 \cup v_2]}.$

Case “ $t = t_1 \ t_2$ ”:

$$\begin{aligned}
 & \llbracket t_1 \ t_2 \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket t_1 \ t_2 \rrbracket_{\rho[x \mapsto v_2]} \\
 &= \bigcup \left\{ \varphi(\llbracket t_2 \rrbracket_{\rho[x \mapsto v_1]}) \mid \varphi \in \llbracket t_1 \rrbracket_{\rho[x \mapsto v_1]} \right\} \cup \bigcup \left\{ \varphi(\llbracket t_2 \rrbracket_{\rho[x \mapsto v_2]}) \mid \varphi \in \llbracket t_1 \rrbracket_{\rho[x \mapsto v_2]} \right\} \\
 &\stackrel{!}{\subseteq} \bigcup \left\{ \varphi(\llbracket t_2 \rrbracket_{\rho[x \mapsto v_1]}) \cup \varphi(\llbracket t_2 \rrbracket_{\rho[x \mapsto v_2]}) \mid \varphi \in \llbracket t_1 \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket t_1 \rrbracket_{\rho[x \mapsto v_2]} \right\} \\
 &\stackrel{!}{\subseteq} \bigcup \left\{ \varphi(\llbracket t_2 \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket t_2 \rrbracket_{\rho[x \mapsto v_2]}) \mid \varphi \in \llbracket t_1 \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket t_1 \rrbracket_{\rho[x \mapsto v_2]} \right\} \\
 &\stackrel{\text{i.h.}}{\subseteq} \bigcup \left\{ \varphi(\llbracket t_2 \rrbracket_{\rho[x \mapsto v_1 \cup v_2]}) \mid \varphi \in \llbracket t_1 \rrbracket_{\rho[x \mapsto v_1 \cup v_2]} \right\} \\
 &= \llbracket t_1 \ t_2 \rrbracket_{\rho[x \mapsto v_1 \cup v_2]}
 \end{aligned}$$

Case “ $t = \emptyset$ ”: $\llbracket \emptyset \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket \emptyset \rrbracket_{\rho[x \mapsto v_2]} = \emptyset = \llbracket \emptyset \rrbracket_{\rho[x \mapsto v_1 \cup v_2]}.$

Case “ $t = \{c\}$ ”: $\llbracket \{c\} \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket \{c\} \rrbracket_{\rho[x \mapsto v_2]} = \{c\} = \llbracket \{c\} \rrbracket_{\rho[x \mapsto v_1 \cup v_2]}.$

$$\begin{aligned} \text{Case } "t = t_1 \cup t_2": \quad \llbracket t_1 \cup t_2 \rrbracket_{\rho[x \mapsto v_1]} \llbracket t_1 \cup t_2 \rrbracket_{\rho[x \mapsto v_2]} &= \llbracket t_1 \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket t_1 \rrbracket_{\rho[x \mapsto v_2]} \cup \\ \llbracket t_2 \rrbracket_{\rho[x \mapsto v_1]} \cup \llbracket t_2 \rrbracket_{\rho[x \mapsto v_2]} &\stackrel{\text{i.h.}}{\subseteq} \llbracket t_1 \rrbracket_{\rho[x \mapsto v_1 \cup v_2]} \cup \llbracket t_2 \rrbracket_{\rho[x \mapsto v_1 \cup v_2]} = \llbracket t_1 \cup t_2 \rrbracket_{\rho[x \mapsto v_1 \cup v_2]}. \end{aligned}$$

□

The inequality of Lemma 6 is not an equality.

Counterexample 1. Let $\text{app} = \lambda f. \lambda x. f \ x$, then $\text{app} \ (\lambda x. \emptyset) \ \{C\} \cup \text{app} \ (\lambda x. x) \ \emptyset \rightsquigarrow \emptyset$, but $\text{app} \ ((\lambda x. \emptyset) \cup (\lambda x. x)) \ (\{C\} \cup \emptyset) \rightsquigarrow \{C\}$.

A.1.3 Normalization (with widening)

- **To do:** We can make union only work on base types (as we not longer need to distribute unions over applications)? Then the denotation of the function space would be simpler and might generalize to other structures..

To reduce λ^\cup -terms to a normal form we combine the β -reduction rule of the simply typed λ -calculus with rewrite rules that deal with the associativity, commutativity, idempotence and identity (ACI1) properties of set-union operator.

If a term t is η -long it can be written in the form

$$t = \lambda x_1 \cdots x_n. \{f_1(t_{11}, \dots, t_{1q_1}), \dots, f_p(t_{p1}, \dots, t_{pq_p})\}$$

where f_i can be a free or bound variable, a singleton-set constant, or another η -long term; and q_i is equal to the arity of f_i (for all $1 \leq i \leq p$). The notation $\{f_1(t_{11}, \dots, t_{1q_1}), \dots, f_p(t_{p1}, \dots, t_{pq_p})\}$ is a shorthand for $f_1(t_{11}, \dots, t_{1q_1}) \cup \dots \cup f_p(t_{p1}, \dots, t_{pq_p})$, where we forget the associativity of the set-union operator and any empty-set constants. Note that despite the suggestive notation, this is not a true set, as there may still be duplicate elements $f_i(t_{i1}, \dots, t_{iq_i})$.

A normal form v of a term t can be written as

$$v = \lambda x_1 \cdots x_n. \{k_1(v_{11}, \dots, v_{1q_1}), \dots, k_p(v_{p1}, \dots, v_{pq_p})\}$$

where k_i can be a free or bound variable, or a singleton-set constant, but not a term as this would form a β -redex.¹ For each k_i, k_j with $i < j$ we must also

¹Technically, terms that bind at least one variable would form a β -redex. Terms that do not bind any variables do not occur either as they merely form a subsequence of $k_1(v_{11}, \dots, v_{1q_1}), \dots, k_p(v_{p1}, \dots, v_{pq_p})$ in this notation.

have that $k_i < k_j$ for some total order on $\mathbf{Var} \cup \mathbf{Con}$. Not only does this imply that each term $k_i(v_{i1}, \dots, v_{iq_i})$ occurs only once in $k_1(v_{11}, \dots, v_{1q_1}), \dots, k_p(v_{p1}, \dots, v_{pq_p})$, but also the stronger condition that $k_i \neq k_j$ for all $i \neq j$.

```

-- normalization of terms
 $\llbracket \cdot \rrbracket :: \mathbf{Tm} \rightarrow \mathbf{Nf}$ 
 $\llbracket \lambda x_1 \dots x_n. T \rrbracket =$ 
   $\lambda x_1 \dots x_n. \{ \llbracket f_i(\llbracket t_{i1} \rrbracket, \dots, \llbracket t_{iq_i} \rrbracket) \rrbracket \mid f_i(t_{i1}, \dots, t_{iq_i}) \in T \}$ 
--  $\beta$ -reduction
 $\llbracket k(v_1, \dots, v_q) \rrbracket$ 
   $= k(v_1, \dots, v_q)$ 
 $\llbracket (\lambda y_1 \dots y_q. T) (v_1, \dots, v_q) \rrbracket$ 
   $= \text{SUBST } x \ y \ z$ 
-- set-rewriting
 $\{ \dots, k.i(\dots), \dots, k.j(\dots), \dots \}$ 
   $\mid k.j < k.i = \{ \dots, k.i(\dots), \dots, k.j(\dots), \dots \}$ 
 $\{ \dots, k(\dots), k(\dots), \dots \}$ 
   $= \{ \dots, k(\dots), \dots \}$ 
 $\{ T \}$ 
   $= T$ 

```

Figure A.1: Normalization algorithm for λ^U -terms.

A.2 Completion

The completion procedure as an algorithm:

```

 $\mathcal{C} : \mathbf{Env} \times \mathbf{Ty} \rightarrow \widehat{\tau} \times \xi \times \mathbf{Env}$ 
 $\mathcal{C} \ \bar{e}_i :: \bar{\kappa}_i \ \mathbf{bool} =$ 
  let  $e$  be fresh
  in  $\langle \widehat{\mathbf{bool}}; e \ \bar{e}_i; e : \bar{\kappa}_i \Rightarrow \mathbf{EXN} \rangle$ 

```

A.3 TODO

- standard polyrec examples (DHM + GSM)
- polyrec does not “come for free”
- type inference: we have a fixpoint a la DHM
- complexity (HM = DEXP, STLC in Statman)
- widening
- “algebraic” effects?
- unexpected decidability
- ack: Vincent + Femke; Andrew + Jeremy + Stephanie + Andres; ST-RC
- check wiki and folder for notes
- exception type of twice (and other h-o funs)
- no slanted-greek for lambda and Lambda
- typeset System F_ω correctly
- re-enable more fields in bibliography, full first names
- Stefan’s terminology: fully parametric vs. fully flexible
- roll Metatheory into earlier sections, add new section Analysis (also add to Overview)
- Untracked exceptions can break information flow security.
- Elaborate in the subsection “Contributions”. Mention prototype?

A.3.1 Introduction

- Why not $\alpha\langle e_1 \rangle \xrightarrow{e_3} \beta\langle e_2 \rangle$?! Give some examples why higher-rankedness is needed. The example on the poster/*map* isn’t sufficient. Postpone to a later section?

A.3.2 The λ^{\cup} -calculus

- Prove semantics is ACI1. We have a different unit for each type!
- $\mathcal{P}(V_{\tau_1} \rightarrow V_{\tau_2}) \simeq V_{\tau_1} \rightarrow \mathcal{P}(V_{\tau_2})$? Cardinality suggests not: $2^{(\beta^\alpha)} \neq (2^\beta)^\alpha$.
- If we don't distribute unions over applications, can we ever get them deep inside terms?
- If we don't *and* the outermost lambdas are not there because is always of kind star, can we get non-trivial terms? I.e. something other than $e_1(e_{11}, \dots, e_{1n_1}) \cup \dots \cup e_k(e_{k1}, \dots, e_{kn_k})$ (note: e and not t as arguments).

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