

# Type-Level Computation One Step at a Time

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## Abstract

Many type systems support a conversion rule that allows type-level computation. In such type systems ensuring the *decidability* of type checking requires type-level computation to terminate. For calculi where the syntax of types and terms is the same, the decidability of type-checking is usually dependent on the strong normalization of the calculus, which ensures termination. An unfortunate consequence of this coupling between decidability and strong normalization is that adding (unrestricted) general recursion to such calculi is not possible.

This paper proposes an alternative to the conversion rule that allows the same syntax for types and terms, type-level computation, and preserves decidability of type-checking under the presence of general recursion. The key idea, which is inspired by the traditional treatment of *iso-recursive types*, is to make each type-level computation step explicit. Each beta reduction or expansion at the type-level is introduced by a language construct. This allows control over the type-level computation and ensures decidability of type-checking even in the presence of non-terminating programs at the type-level. We realize this idea by presenting a variant of the calculus of constructions with general recursion and recursive types. Furthermore we show how many advanced programming language features of state-of-the-art functional languages (such as Haskell) can be encoded in our minimalistic core calculus.

**Categories and Subject Descriptors** D.3.1 [Programming Languages]: Formal Definitions and Theory

**General Terms** Languages, Design

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## 1. Introduction

Modern statically typed functional languages (such as ML, Haskell, Scala or OCaml) have increasingly expressive type systems. Often these large source languages are translated into a much smaller typed core language. The choice of the core language is essential to ensure that all the features of the source language can be encoded. For a simple polymorphic functional language it is possible, for example, to pick a variant of System  $F$  as a core language. However, the desire for more expressive type system features puts pressure on

the core languages, often requiring them to be extended to support new features. For example, if the source language supports *higher-kinded types* or *type-level functions* then System  $F$  is not expressive enough and can no longer be used as the core language. Instead another core language that does provide support for higher-kinded types, such as System  $F_\omega$ , needs to be used. However System  $F_\omega$  is significantly more complex than System  $F$  and thus harder to maintain. If later a new feature, such as *kind polymorphism*, is desired the core language may need to be changed again to account for the new feature, introducing at the same time new sources of complexity. Indeed the core language for modern versions of functional languages are quite complex, having multiple syntactic sorts (such as terms, types and kinds), as well as dozens of language constructs [BRUNO:  $F_C$ ].

The more expressive type systems become, the more types become similar to the terms. Therefore a natural idea is to unify terms and types. There are obvious benefits in this approach: only one syntactic level (terms) is needed; and there are much less language constructs, making the core language easier to implement and maintain. At the same time the core language becomes more expressive, giving us for free many useful language features. *Pure type systems* [] build on this observation and they show how a whole family of type systems (including System  $F$  and System  $F_\omega$ ) can be implemented using just a single syntactic form. With the added expressiveness it is even possible to have type-level programs expressed using the same syntax as terms as well as dependently typed programs [].

However having the same syntax for types and terms can also be problematic. If arbitrary type-level computation is allowed then type-level programs can use the same language constructs as terms. Usually type systems have a conversion rule to support type-level computation. In such type systems ensuring the *decidability* of type checking requires type-level computation to terminate. For calculi where the syntax of types and terms is the same, the decidability of type-checking is usually dependent on the strong normalization of the calculus, which ensures termination. An unfortunate consequence of this coupling between decidability and strong normalization is that adding (unrestricted) general recursion to such calculi is not possible. There is a clear tension between decidability of type-checking and allowing general recursion at the type-level.

This paper proposes  $\lambda_*$ : a variant of the calculus of constructions allows the same syntax for types and terms, supports type-level computation, and preserves decidability of type-checking under the presence of general recursion. In  $\lambda_*$ , each type-level computation step is explicit. **BRUNO: emphasis on the advantages: a minimal core language?** The key idea, which is inspired by the traditional treatment of *iso-recursive types*, is to introduce each beta reduction or expansion at the type-level by a *type-safe cast*. The casts allow control over the type-level computation. For example, if a type-level program requires two beta-reductions to reach normal form, then two casts are needed in the program. If a non-terminating program is used at the type-level, it is not pos-

sible to cause non-termination in the type-checker, because that would require a program with an infinite number of casts. Therefore, since single beta-steps are trivially terminating, decidability of type-checking is possible even in the presence of non-terminating programs at the type-level.

Our motivation to develop  $\lambda_*$  is to use it as a simpler alternative to existing core languages for languages such as Haskell. The paper shows how many of programming language features of Haskell, including some of the latest extensions, can be encoded in  $\lambda_*$  via a source language. In particular the source language supports *algebraic datatypes*, *higher-kinded types*, *nested datatypes*, *kind polymorphism* [] and *datatype promotion* []. This result is interesting because  $\lambda_*$  is a minimal calculus with only 8 language constructs and a single syntactic sort. In contrast the latest versions of System  $F_C$  (Haskell's core language) have multiple syntactic sorts and dozens of language constructs. Even if support for equality and coercions, which constitutes a significant part of System  $F_C$ , would be removed the resulting language would still be significantly larger and more complex than  $\lambda_*$ .

We believe  $\lambda_*$  is particularly well-suited as a core for Haskell-like languages. In particular the treatment of type-level computation shares similar ideas with Haskell. Although Haskell's surface language provides a rich set of mechanisms to do type-level computation [], the core language lacks fundamental mechanisms to do type-level computation. In particular, like in  $\lambda_*$ , type equality is purely syntactic (modulo alpha-conversion).

In summary, the contributions of this work are:

- Decidable type checking and managed type-level computation by replacing implicit conversion rule of CoC with generalized fold/unfold semantics.
- A core language based on Calculus of Constructions (CoC) that collapses terms, types and kinds into the same hierarchy, supports general recursion...
- General recursion by introducing recursive types for both terms and types by the same  $\mu$  primitive.
- Surface language that supports datatypes, pattern matching and other language extensions for Haskell, and can be encoded into the core language.

## 2. Overview

This section informally introduces the main features of  $\lambda_*$ . In particular, this section shows how the explicit casts in  $\lambda_*$  can be used instead of the typical conversion rule present in calculi such as the calculus of constructions. The formal details of  $\lambda_*$  are presented in §??.

### 2.1 The Calculus of Constructions and the Conversion Rule

The calculus of constructions ( $\lambda C$ ) [5] is a powerful higher-order typed lambda calculus supporting dependent types (among various other features). A crucial feature of  $\lambda C$  is the so-called *conversion* rule as shown below:

$$\frac{\Gamma \vdash e : \tau_1 \quad \Gamma \vdash \tau_2 : s \quad \tau_1 =_\beta \tau_2}{\Gamma \vdash e : \tau_2} \text{ TCC.CONV}$$

The conversion rule allows one to derive  $e : \tau_2$  from the derivation of  $e : \tau_1$  and the  $\beta$ -equality of  $\tau_1$  and  $\tau_2$ . This rule is important to *automatically* allows terms with equivalent types to be considered type-compatible. To illustrate, let us consider a simple example. Suppose we have a built-in base type  $\text{Int}$  and

$$f \equiv \lambda x : (\lambda y : \star.y) \text{Int}.x$$

Here  $f$  is a simple identity function. However, the type of  $x$  ( $(\lambda y : \star.y) \text{Int}$ ), which is the argument of  $f$ , is interesting: it is an identity

function on types, applied to an integer. Without the conversion rule,  $f$  cannot be applied to, say 3 in  $\lambda C$ . However, given that  $f$  is actually  $\beta$ -convertible to  $\lambda x : \text{Int}.x$ , the conversion rule allows the application of  $f$  to 3 by implicitly converting  $\lambda x : (\lambda y : \star.y) \text{Int}.x$  to  $\lambda x : \text{Int}.x$ .

**Decidability of Type-Checking and Strong Normalization** While the conversion rule in  $\lambda C$  brings a lot of convenience, an unfortunate consequence is that it couples decidability of type-checking with strong normalization of the calculus [4]. However strong normalization does not hold with general recursion. This is simply because due to the conversion rule, any non-terminating term would force the type checker to go into an infinitely loop (by constantly applying the conversion rule without termination), thus rendering the type system undecidable. **BRUNO: show simple example. Explain issue better.**

### 2.2 An Alternative to the Conversion Rule: Explicit Casts

**BRUNO: Mention somewhere that the cast rules do one-step reductions.** In contrast to the implicit reduction rules of  $\lambda C$ ,  $\lambda_*$  makes it explicit as to when and where to convert one type to another. Type conversions are explicit by introducing two language constructs:  $\text{cast}^\uparrow$  (beta expansion) and  $\text{cast}_\downarrow$  (beta reduction). The benefit of this approach is that decidability of type-checking no longer is coupled with strong normalization of the calculus.

**Beta Expansion** The first of the two type conversions  $\text{cast}^\uparrow$ , allows a type conversion provided that the resulting type is a *beta expansion* of the original type of the term. Consider the same example from §2.1. In  $\lambda_*$ ,  $f\ 3$  is an ill-typed application. Instead we must write the application as

$$f (\text{cast}^\uparrow [(\lambda y : \star.y) \text{Int}] 3)$$

**BRUNO: how to put a space before 3?** The intuition is that,  $\text{cast}^\uparrow$  is actually doing a type conversion since the type of 3 is  $\text{Int}$  and  $(\lambda y : \star.y) \text{Int}$  can be reduced to  $\text{Int}$ . **BRUNO: explain why this is a beta expansion** **BRUNO: explain why for beta expansions we need to provide the resulting type as argument.**

**Beta Reduction** The dual operation of  $\text{cast}^\uparrow$  is  $\text{cast}_\downarrow$ , which allows a type conversion provided that the resulting type is a *beta reduction* of the original type of the term. The use of  $\text{cast}_\downarrow$  is better explained by another similar example. Suppose that

$$g \equiv \lambda x : \text{Int}.x$$

and term  $z$  has type

$$(\lambda y : \star.y) \text{Int}$$

$g\ z$  is again an ill-typed application, while  $g (\text{cast}_\downarrow z)$  is type correct because  $\text{cast}_\downarrow$  reduces the type of  $z$  to  $\text{Int}$ . **BRUNO: explain why this is a reduction**

### 2.3 Decidability without Strong Normalization

With explicit type conversion rules the decidability of the type system no longer depends on the normalization property. A nice consequence of this is that the type system remains decidable even in the presence of non-terminating programs at type level. To illustrate, let us consider the following example. Suppose that  $d$  is a “dependent type” where

$$d : \text{Int} \rightarrow \star$$

so that  $d\ 3$  or  $d\ 100$  all yield the same type. With general recursion at hand, we can imagine a term  $z$  that has type

$$d\ \text{loop}$$

where `loop` stands for any diverging computation and of type `Int`. What would happen if we try to type check the following application:

$$(\lambda x : d\ 3.x)\ z$$

Under the normal typing rules of  $\lambda C$ , the type checker would get stuck as it tries to do  $\beta$ -equality on two terms:  $d\ 3$  and  $d\ \text{loop}$ , where the latter is non-terminating.

This is not the case for  $\lambda_*$ :  $\lambda_*$  has no conversion rule, therefore the type checker would do syntactic comparison between the two terms instead of  $\beta$ -equality in the above example and reject the program as ill-typed. Indeed it would be impossible to type-check the program even with the use of  $\text{cast}^\uparrow$  and/or  $\text{cast}_\downarrow$ : one would need to write infinite number of  $\text{cast}_\downarrow$ 's to make the type checker loop forever (e.g.,  $(\lambda x : d\ 3.x)(\text{cast}_\downarrow(\text{cast}_\downarrow \dots z))$ ), but it is impossible to write such program in reality.

In summary,  $\lambda_*$  achieves the decidability of type checking by explicitly controlling type-level computation, which is independent of the normalization property, while supporting general recursion at the same time.

## 2.4 Unifying Recursive Types and Recursion

**BRUNO:** Show how in  $\lambda_*$  recursion and recursive types are unified. Discuss that due to this unification the sensible choice for the evaluation strategy is call-by-name.

Recursive types arise naturally if we want to do general recursion.  $\lambda_*$  differs from other programming languages in that it unifies both recursion and recursive types by the same  $\mu$  primitive.

**Recursive types** In the literature on type systems, there are two approaches to recursive types. One is called *equi-recursive*, the other *iso-recursive*.  $\lambda_*$  takes the latter approach since it is more intuitive to us with regard to recursion. The *iso-recursive* approach treats a recursive type and its unfolding as different, but isomorphic. In  $\lambda_*$ , this is witnessed by first  $\text{cast}^\uparrow$ , then  $\text{cast}_\downarrow$ . A classic example of recursive types is the so-called “hungry” type:  $H = \mu\sigma : \star. \text{Int} \rightarrow \sigma$ . A term  $z$  of type  $H$  can accept any number of numeric arguments and return a new function that is hungry for more, as illustrated below:

$$\begin{aligned} \text{cast}_\downarrow z : \text{Int} \rightarrow H \\ \text{cast}_\downarrow(\text{cast}_\downarrow z) : \text{Int} \rightarrow \text{Int} \rightarrow H \\ \text{cast}_\downarrow(\text{cast}_\downarrow \dots z) : \text{Int} \rightarrow \text{Int} \rightarrow \dots \rightarrow H \end{aligned}$$

**Recursion** The same  $\mu$  primitive can also be used to define recursive functions, e.g., the factorial function:

$$\mu f : \text{Int} \rightarrow \text{Int}. \lambda x : \text{Int}. \text{if } (x == 0) \text{ then } 1 \text{ else } x * f(x - 1)$$

This is reflected by the dynamic semantics of the  $\mu$  primitive:

$$\mu x : T. E \longrightarrow E[x := \mu x : T. E]$$

which is exactly doing recursive unfolding of the same term.

Due to the unification, the *call-by-value* evaluation strategy does not fit in our setting. In call-by-value evaluation, recursion can be expressed by the recursive binder  $\mu$  as  $\mu f : T \rightarrow T. E$  (note that the type of  $f$  is restricted to function types). Since we don't want to pose restrictions on the types, the *call-by-name* evaluation is a sensible choice.

## 2.5 Encoding Datatypes

**BRUNO:** Informally explain how to encode recursive datatypes and recursive functions using datatypes.

With the explicit type conversion rules and the  $\mu$  primitive, it is straightforward to encode recursive datatypes and recursive functions using datatypes. While inductive datatypes can be encoded using either the Church or the Scott encoding, we adopt the Scott

encoding as it bears some resemblance to case analysis, making it more convenient to encode pattern matching. We demonstrate the encoding method using a simple datatype as a running example: the natural numbers.

The datatype declaration for natural numbers is:

**data**  $\text{Nat} = Z \mid S\ \text{Nat};$

In the Scoot encoding, the encoding of the  $\text{Nat}$  type reflects how its two constructors are going to be used. Since  $\text{Nat}$  is a recursive datatype, we have to use recursive types at some point to reflect its recursive nature. As it turns out, the  $\text{Nat}$  type can be simply represented as  $\mu X : \star. \Pi B : \star. B \rightarrow (X \rightarrow B) \rightarrow B$ .

As can be seen, in the function type  $B \rightarrow (X \rightarrow B) \rightarrow B$ ,  $B$  corresponds to the type of the  $Z$  constructor, and  $X \rightarrow B$  corresponds to the type of the  $S$  constructor. The intuition is that any use of the datatype being defined in the constructors is replaced with the recursive type, except for the return type, which is a type variable for use in the recursive functions.

Now its two constructors can be encoded correspondingly as below:

```
let Z : Nat = cast↑ [Nat] (λB : ⋆. λz : B. λf : Nat → B. z)
in
let S : Nat → Nat = λn : Nat.
  cast↑ [Nat] (λB : ⋆. λz : B. λf : Nat → B. f n)
```

Thanks to the explicit type conversion rules, we can make use of the  $\text{cast}^\uparrow$  operation to do type conversion between the recursive type and its unfolding.

As the last example, let us see how we can define recursive functions using the  $\text{Nat}$  datatype. A simple example would be recursively adding two natural numbers, which can be defined as below:

```
let add : Nat → Nat → Nat = μ f : Nat → Nat → Nat.
  λn : Nat. λm : Nat.
    (cast↓ n) Nat m (λn' : Nat. S (f n' m))
```

As we can see, the above definition quite resembles case analysis common in modern functional programming languages. (Actually we formalize the encoding of case analysis in §6.)

Due to the unification of recursive types and recursion, we can use the same  $\mu$  primitive to write both recursive types and recursion with ease.

## 3. Showcase

In this section, we showcase applications, which either Haskell needs non-trivial extensions to do that, or dependently typed languages like Coq and Agda are impossible to do, whereas we can easily achieve in  $\lambda_*$ .

**Conventional datatypes** Conventional datatypes like natural numbers or polymorphic lists can be easily defined in  $\lambda_*$ , as in Haskell. For example, below is the definition of polymorphic lists:

**data**  $\text{List } (a : \star) = \text{Nil} \mid \text{Cons } a\ (\text{List } a);$

Because  $\lambda_*$  is explicitly typed, each type parameter needs to be accompanied with corresponding kind expressions. The use of the above datatype is best illustrated by the *length* function:

```

letrec length : (a :  $\star$ )  $\rightarrow$  List a  $\rightarrow$  nat =
   $\lambda a : \star. \lambda l : \text{List } a. \text{case } l \text{ of}$ 
    | Nil  $\Rightarrow$  0
    | Cons (x : a) (xs : List a)  $\Rightarrow$  1 + length a xs
in
let test : List nat = Cons nat 1 (Cons nat 2 (Nil nat))
in length nat test -- return 2

```

**Higher-kinded types** Higher-kinded types are types that take other types and produce a new type. To support higher-kinded types, languages like Haskell have to extend their existing core languages to account for kind expressions. In  $\lambda_\star$ , since all syntactical constructs are in the same level, we can easily construct higher-kinded types. We show this by an example of encoding the *Functor* class:

```

rcrd Functor (f :  $\star \rightarrow \star$ ) =
  Func {fmap : (a :  $\star$ )  $\rightarrow$  (b :  $\star$ )  $\rightarrow$  f a  $\rightarrow$  f b};

```

A functor is just a record that has only one field *fmap*. A Functor instance of the *Maybe* datatype is simply:

```

let maybeInst : Functor Maybe =
  Func Maybe ( $\lambda a : \star. \lambda b : \star. \lambda f : a \rightarrow b. \lambda x : \text{Maybe } a. \text{case } x \text{ of}$ 
    | Nothing  $\Rightarrow$  Nothing b
    | Just (z : a)  $\Rightarrow$  Just b (f z))

```

**HOAS** Higher-order abstract syntax is a generalization of representing programs where the function space of the meta-language is used to encode the binders of the object language. Because the recursive mention of the datatype can appear in a negative position, systems like Coq and Agda would reject programs using HOAS due to the restrictiveness of their termination checkers. However  $\lambda_\star$  is able to express HOAS in a straightforward way. We show an example of encoding a simple lambda calculus:

```

data Exp = Num nat
  | Lam (Exp  $\rightarrow$  Exp)
  | App Exp Exp;

```

Next we define the evaluator for our lambda calculus. As noted by [6], the evaluation function needs an extra function *reify* to invert the result of evaluation. The code is presented in Figure 1.

The definition of the evaluator is quite straightforward, although it is worth noting that the evaluator is a partial function that can cause run-time errors. Thanks to the flexibility of the  $\mu$  primitive, mutual recursion can be encoded by using records!

Evaluation of a lambda expression proceeds as follows:

```

let test : Exp = App (Lam ( $\lambda f : \text{Exp}. \text{App } f \text{ (Num 42)}$ ))
  (Lam ( $\lambda g : \text{Exp}. g$ ))
in show (eval test) -- return 42

```

**Fix as a datatype** The type-level *Fix* is a good example to demonstrate the expressiveness of  $\lambda_\star$ . The definition is simply:

```

rcrd Fix (f :  $\star \rightarrow \star$ ) = In {out : (f (Fix f))};

```

The record notation also introduces the selector function: *out* : ( $f : \star \rightarrow \star$ )  $\rightarrow$  Fix f  $\rightarrow$  f (Fix f). The *Fix* datatype is interesting in that Coq and Agda would reject this definition because the use of the higher-kinded type parameter *f* could be anywhere (e.g., in

```

data Value = VI nat | VF (Value  $\rightarrow$  Value);
rcrd Eval = Ev {eval' : Exp  $\rightarrow$  Value, reify' : Value  $\rightarrow$  Exp};
let f : Eval =  $\mu f' : \text{Eval}.$ 
  Ev ( $\lambda e : \text{Exp}. \text{case } e \text{ of}$ 
    | Num (n : nat)  $\Rightarrow$  VI n
    | Lam (fun : Exp  $\rightarrow$  Exp)  $\Rightarrow$ 
      VF ( $\lambda e' : \text{Value}. \text{eval}' f' (\text{fun } (\text{reify}' f' e'))$ )
    | App (a : Exp) (b : Exp)  $\Rightarrow$ 
      case eval' f' a of
        | VI (n : nat)  $\Rightarrow$  error
        | VF (fun : Value  $\rightarrow$  Value)  $\Rightarrow$  fun (eval' f' b))
    ( $\lambda v : \text{Value}. \text{case } v \text{ of}$ 
      | VI (n : nat)  $\Rightarrow$  Num n
      | VF (fun : Value  $\rightarrow$  Value)  $\Rightarrow$ 
        Lam ( $\lambda e' : \text{Exp}. (\text{reify}' f' (\text{fun } (\text{eval}' f' e'))$ ))))
in let eval : Exp  $\rightarrow$  Value = eval' f in

```

**Figure 1.** An evaluator for the HOAS-encoded lambda calculus.

a negative position). However in  $\lambda_\star$ , where type-level computation is explicitly controlled, we can safely use *Fix* in the program.

Given *fmap*, many recursive schemes can be defined, for example we can have *catamorphism* or generic function fold:

```

letrec cata : (f :  $\star \rightarrow \star$ )  $\rightarrow$  (a :  $\star$ )  $\rightarrow$ 
  Functor f  $\rightarrow$  (f a  $\rightarrow$  a)  $\rightarrow$  Fix f  $\rightarrow$  a =
   $\lambda f : \star \rightarrow \star. \lambda a : \star. \lambda m : \text{Functor } f. \lambda g : f a \rightarrow a. \lambda t : \text{Fix } f. \lambda$ 
  g (fmap f m (Fix f) a) (cata f a m g) (out f t))

```

**Kind polymorphism for datatypes** In Haskell, System  $F_c^\uparrow$  was proposed to support kind polymorphism. However it separates expressions into terms, types and kinds, which complicates both the implementation and future extensions.  $\lambda_\star$  natively allows datatype definitions to have polymorphic kinds. Here is an example, taken from [16], of a datatype that benefits from kind polymorphism: higher-kinded fixpoint combinator

```

data Mu (k :  $\star$ ) (f : (k  $\rightarrow$   $\star$ )  $\rightarrow$  k  $\rightarrow$   $\star$ ) (a : k) =
  Roll (f (Mu k f) a);

```

*Mu* can be used to construct polymorphic recursive types of any kind, for instance:

```

data Listf (f :  $\star \rightarrow \star$ ) (a :  $\star$ ) = Nil | Cons a (f a);
let List :  $\star \rightarrow \star = \lambda a : \star. \text{Mu } \star \text{ Listf } a$ 
```

**Nested datatypes and polymorphic recursion** A nested datatype, also known as a *non-regular* datatype, is a parametrised datatype whose definition contains different instances of the type parameters. Functions over nested datatypes usually involve polymorphic recursion. We show that  $\lambda_\star$  is capable of defining all useful functions over a nested datatype. A simple example would be the type *Pow* of power trees, whose size is exactly a power of two, declared as follows:

```

data PairT (a :  $\star$ ) = P a a;
data Pow (a :  $\star$ ) = Zero a | Succ (Pow (PairT a));

```

Notice that the recursive mention of *Pow* does not hold *a*, but *PairT a*. This means every time we use a *Succ* constructor, the size of the pairs doubles. In case you are curious about the encoding of *Pow*, here is the one:

let  $Pow : \star \rightarrow \star = \mu X : \star \rightarrow \star.$   
 $\lambda a : \star. (B : \star) \rightarrow (a \rightarrow B) \rightarrow (X (PairT\ a) \rightarrow B) \rightarrow B$

Notice how the higher-kinded type variable  $X : \star \rightarrow \star$  helps encoding nested datatypes. Below is a simple function *toList* that transforms a power tree into a list:

letrec  $toList : (a : \star) \rightarrow Pow\ a \rightarrow List\ a =$   
 $\lambda a : \star. \lambda t : Pow\ a. \text{case } t \text{ of}$   
 $\quad Zero\ (x : a) \Rightarrow Cons\ a\ x\ (Nil\ a)$   
 $\quad | Succ\ (c : Pow\ (PairT\ a)) \Rightarrow$   
 $\quad \quad concatMap\ (PairT\ a)\ a$   
 $\quad \quad (\lambda x : PairT\ a. \text{case } x \text{ of}$   
 $\quad \quad \quad P\ (m : a)\ (n : a) \Rightarrow$   
 $\quad \quad \quad Cons\ a\ m\ (Cons\ a\ n\ (Nil\ a)))$   
 $\quad \quad (toList\ (PairT\ a)\ c)$

#### 4. The core language without recursion

In this section, we present  $\lambda C_{exp}$  calculus, a language based on  $\lambda C$  with explicit type-level computation.  $\lambda C_{exp}$  is a subset of the core language  $\lambda_\star$  without general recursion, which is simple enough for meta-theoretic studies. By explicitly controlling the type-level computation with cast operators,  $\lambda C_{exp}$  can achieve decidable type checking without requiring normalizing property. In rest of this section, we demonstrate the syntax, operational semantics, typing rules and meta-theory of  $\lambda C_{exp}$ .

##### 4.1 Syntax

Figure 2 shows the syntax of  $\lambda C_{exp}$ , including expressions, contexts and values. The syntax follows  $\lambda C$  and fairly straightforward while there are still some differences in  $\lambda C_{exp}$ , including the ‘type-in-type’ axiom and cast operators.

**Unified syntactic levels**  $\lambda C_{exp}$  uses a unified syntactic representation for different levels of expressions by following the *pure type system* (PTS) representation of  $\lambda C$ . Traditionally in  $\lambda C$ , there are two distinct sorts  $\star$  and  $\square$  representing the type of *types* and *sorts* respectively, and an axiom  $\star : \square$  specifying the relation. In  $\lambda C_{exp}$ , we further merge types and kinds together by including only a single sort  $\star$  and an impredicative axiom  $\star : \star$ .

Therefore, there is no syntactic distinction between terms, types or kinds. This design brings the economy for type checking, since one set of rules can cover all syntactic levels. By convention, we use metavariables  $\tau$  and  $\sigma$  for an expression on the type-level position and  $e$  for one on the term level.

**Dependent function types** In the context of  $\lambda C$ , if a term  $x$  has the type  $\tau_1$ , and  $\tau_2$  is a type, i.e.  $x : \tau_1 : \star$  and  $\tau_2 : \square$ , we call the type  $\Pi x : \tau_1. \tau_2$  a *dependent product*.  $\lambda C_{exp}$  follows  $\lambda C$  to use the same  $\Pi$ -notation to represent dependent function types.

However, a higher-kind polymorphic function type such as  $\Pi x : \square. x \rightarrow x$  is not allowed in  $\lambda C$ , because  $\square$  is the highest sort that can not be typed. While  $\Pi$ -notation in  $\lambda C_{exp}$  is more expressive and does not have such limitation because of the axiom  $\star : \star$ . In the source language, we interchangeably use the arrow form  $(x : \tau_1) \rightarrow \tau_2$  of the product for clarity. By convention, we also use the syntactic sugar  $\tau_1 \rightarrow \tau_2$  to represent the product if  $x$  does not occur free in  $\tau_2$ .

**Explicit type conversion** We introduce two new primitives  $\text{cast}^\uparrow$  and  $\text{cast}_\downarrow$  (pronounced as ‘cast up’ and ‘cast down’) to replace implicit conversion rule of  $\lambda C$  with *one-step* explicit type conversion. They represent two directions of type conversion:  $\text{cast}_\downarrow$  stands for

$e, \tau, \sigma$	$::=$	Expressions
$x$		Variable
$\star$		Type of types
$e_1\ e_2$		Application
$\lambda x : \tau. e$		Abstraction
$\Pi x : \tau. \sigma$		Product
$\text{cast}^\uparrow [\tau] e$		Cast up by expansion
$\text{cast}_\downarrow e$		Cast down by reduction
$\Gamma$	$::=$	Contexts
$\emptyset$		Empty
$\Gamma, x : \tau$		Variable binding
$v$	$::=$	Values
$\star$		Type of types
$\lambda x : \tau. e$		Abstraction
$\Pi x : \tau_1. \tau_2$		Product
$\text{cast}^\uparrow [\tau] e$		Cast up by expansion

Figure 2. Syntax of  $\lambda C_{exp}$

the  $\beta$ -reduction of types, while  $\text{cast}^\uparrow$  is the inverse (see examples in §??).

Though cast primitives make the syntax verbose when type conversion is heavily used, the implementation of type checking is simplified because typing rules of  $\lambda C_{exp}$  become type-directed without  $\lambda C$ ’s implicit conversion rule. Considering the core language is compiler-oriented and **source language does not include cast primitives**, end-users will not directly use them. Some type conversions can be generated through the translation of the source language (§??).

##### 4.2 Operational semantics

Figure 3 shows the *call-by-name* operational semantics, defined by one-step reduction. Two base cases include S\_BETA for  $\beta$ -reduction and S\_CASTDOWNUP for cast canceling. Two inductive case, S\_APP and S\_CASTDOWN, define reduction in the head position of an application, and in the  $\text{cast}_\downarrow$  inner expression respectively. The reduction rules are *weak* in the sense that it is not allowed to reduce inside a  $\lambda$ -term or  $\text{cast}^\uparrow$ -term which is viewed as a value (see Figure 2).

To evaluate the value of a term-level expression, we apply the one-step reduction multiple times. The number of evaluation steps is not restricted, which is possible to be infinite. The multi-step reduction can be defined as follows:

**Definition 4.1** (Multi-step reduction). *The relation  $\rightarrow$  is the transitive and reflexive closure of the one-step reduction  $\rightarrow$ .*

But for a type-level expression, the evaluation is driven by cast operators, which is finite. For a consecutive sequence of reductions with  $n$  steps, we use the notation  $\rightarrow_n$  to denote the relation between the initial and final expressions:

**Definition 4.2** ( $n$ -step reduction). *The  $n$ -step reduction is denoted by  $e_0 \rightarrow_n e_n$ , if there exists a sequence of one-step reductions  $e_0 \rightarrow e_1 \rightarrow e_2 \rightarrow \dots \rightarrow e_n$ , where  $n$  is a positive integer and  $e_i$  ( $i = 0, 1, \dots, n$ ) are valid expressions.*

##### 4.3 Typing

Figure 4 gives the *syntax-directed* typing rules of  $\lambda C_{exp}$ , including rules of context well-formedness  $\vdash \Gamma$  and expression typing  $\Gamma \vdash e : \tau$ . Note that there is only a single set of rules for expression typing, because there is no distinction of different syntactic levels.



$$\boxed{e \longrightarrow e'} \quad \text{One-step reduction}$$

$$\begin{array}{c}
\frac{}{(\lambda x : \tau. e_1) e_2 \longrightarrow e_1[x \mapsto e_2]} \text{S\_BETA} \\
\frac{}{\text{cast}_{\downarrow}(\text{cast}_{\uparrow}^{\tau}[e]) \longrightarrow e} \text{S\_CASTDOWNUP} \\
\frac{e_1 \longrightarrow e'_1}{e_1 e_2 \longrightarrow e'_1 e_2} \text{S\_APP} \\
\frac{e \longrightarrow e'}{\text{cast}_{\downarrow} e \longrightarrow \text{cast}_{\downarrow} e'} \text{S\_CASTDOWN}
\end{array}$$

**Figure 3.** Operational semantics of  $\lambda C_{\text{exp}}$

Most typing rules are quite standard. We write  $\vdash \Gamma$  if a context  $\Gamma$  is well-formed. Note that there is only a single sort  $\star$ , we use  $\Gamma \vdash \tau : \star$  to check if  $\tau$  is a well-formed type. Rule T\_AX is the ‘type-in-type’ axiom. Rule T\_VAR checks the type of variable  $x$  from the valid context. Rules T\_APP and T\_LAM check the validity of application and abstraction. Rules T\_PI check the type well-formedness of the dependent function.

We focus on rules T\_CASTUP and T\_CASTDOWN that define the semantics of cast operators and replace the conversion rule of  $\lambda C$  (see ??). The relation between the original and converted type is defined by one-step reduction (see 3). Specifically speaking, if given a judgement  $\Gamma \vdash e : \tau_2$  and relation  $\tau_1 \longrightarrow \tau_2 \longrightarrow \tau_3$ , then  $\text{cast}_{\uparrow}^{\tau_1} e$  expands the type of  $e$  from  $\tau_2$  to  $\tau_1$ , while  $\text{cast}_{\downarrow} e$  reduces the type of  $e$  from  $\tau_2$  to  $\tau_3$ .

Finally, the definition of type equality in  $\lambda C_{\text{exp}}$  differs from  $\lambda C$ . Without  $\lambda C$ ’s conversion rule, the type of a term cannot be converted freely against  $\beta$ -equality, unless using cast operators. Thus, types of expressions are equal only if they are syntactically equal, i.e. satisfy the  $\alpha$ -equality.

#### 4.4 Meta-theory

We now discuss the meta-theory of  $\lambda C_{\text{exp}}$ . We focus on two properties: the decidability of type checking and the type-safety of the language. First, we want to show type checking  $\lambda C_{\text{exp}}$  is decidable without normalizing property. The type checker will not be stuck by type-level non-termination. Second, the language is type safe, proven by standard subject reduction and progress lemmas.

**Decidability of type checking** For the decidability, we need to show there exists a type checking algorithm, which never loops forever and returns a unique type for a well-formed expression  $e$ . This is done by induction on the length of  $e$  and ranging over typing rules. Most expression typing rules, which have only typing judgements in premises, are already decidable by induction hypothesis. Thus, it is straightforward to follow the syntax-directed judgement to derive a unique type checking result.

The critical case is for rules T\_CASTUP and T\_CASTDOWN. Both rules contain a premise that needs to judge if two types  $\tau_1$  and  $\tau_2$  follows the one-step reduction, i.e. if  $\tau_1 \longrightarrow \tau_2$  holds. We need to show such  $\tau_2$  is *unique* with respect to the one-step reduction, or equivalently, reducing  $\tau_1$  by one step will get only a sole result  $\tau_2$ . Otherwise, assume  $e : \tau_1$  and there exists  $\tau'_2$  such that  $\tau_1 \longrightarrow \tau_2$  and  $\tau_1 \longrightarrow \tau'_2$ . Then the type of  $\text{cast}_{\downarrow} e$  can be either  $\tau_2$  or  $\tau'_2$  by rule T\_CASTDOWN, which is not decidable. The property is proven by the following lemma:

**Lemma 4.3** (Uniqueness of one-step reduction). *The relation  $\longrightarrow$ , i.e. one-step reduction, is unique in the sense that given  $e$  there is at most one  $e'$  such that  $e \longrightarrow e'$ .*

$$\boxed{\vdash \Gamma} \quad \text{Well-formed context}$$

$$\begin{array}{c}
\frac{}{\vdash \emptyset} \text{ENV\_EMPTY} \\
\frac{\vdash \Gamma \quad \Gamma \vdash \tau : \star}{\vdash \Gamma, x : \tau} \text{ENV\_VAR} \\
\boxed{\Gamma \vdash e : \tau} \quad \text{Expression typing} \\
\frac{}{\emptyset \vdash \star : \star} \text{T\_AX} \\
\frac{\vdash \Gamma \quad x : \tau \in \Gamma}{\Gamma \vdash x : \tau} \text{T\_VAR} \\
\frac{\Gamma \vdash e_1 : (\Pi x : \tau_2. \tau_1) \quad \Gamma \vdash e_2 : \tau_2}{\Gamma \vdash e_1 e_2 : \tau_1[x \mapsto e_2]} \text{T\_APP} \\
\frac{\Gamma, x : \tau_1 \vdash e : \tau_2 \quad \Gamma \vdash (\Pi x : \tau_1. \tau_2) : \star}{\Gamma \vdash (\lambda x : \tau_1. e) : (\Pi x : \tau_1. \tau_2)} \text{T\_LAM} \\
\frac{\Gamma \vdash \tau_1 : \star \quad \Gamma, x : \tau_1 \vdash \tau_2 : \star}{\Gamma \vdash (\Pi x : \tau_1. \tau_2) : \star} \text{T\_PI} \\
\frac{\Gamma \vdash e : \tau_2 \quad \Gamma \vdash \tau_1 : \star \quad \tau_1 \longrightarrow \tau_2}{\Gamma \vdash (\text{cast}_{\uparrow}^{\tau_1} e) : \tau_1} \text{T\_CASTUP} \\
\frac{\Gamma \vdash e : \tau_1 \quad \Gamma \vdash \tau_2 : \star \quad \tau_1 \longrightarrow \tau_2}{\Gamma \vdash (\text{cast}_{\downarrow} e) : \tau_2} \text{T\_CASTDOWN}
\end{array}$$

**Figure 4.** Typing rules of  $\lambda C_{\text{exp}}$

*Proof.* By induction on the structure of  $e$ .  $\square$

With this result, we show a decidable algorithm to check whether one-step relation  $\tau_1 \longrightarrow \tau_2$  holds. An intuitive algorithm is to reduce the type  $\tau_1$  by one step to obtain  $\tau'_1$  (which is unique by Lemma 4.3), and compare if  $\tau'_1$  and  $\tau_2$  are syntactically equal. Thus, checking if  $\tau_1 \longrightarrow \tau_2$  is decidable and rules T\_CASTUP and T\_CASTDOWN are therefore decidable. We can conclude the decidability of type checking:

**Lemma 4.4** (Decidability of type checking). *There is a decidable algorithm which given  $\Gamma, e$  computes the unique  $\tau$  such that  $\Gamma \vdash e : \tau$  or reports there is no such  $\tau$ .*

*Proof.* By induction on the structure of  $e$ .  $\square$

Note that when proving the decidability of type checking, we do not rely on the normalizing property. Because explicit type conversion rules use one-step reduction, which already has a decidable checking algorithm according to Lemma 4.3. We do not need to further require the normalization of terms. This is different from the proof for  $\lambda C$  which requires the language is normalizing [cite the PTS paper]. Because  $\lambda C$ ’s conversion rule needs to examine the  $\beta$ -equivalence of terms, which is decidable only if every term has a normal form.

**Multi-step cast operator** Write about  $\text{cast}_{\uparrow}^n$  and  $\text{cast}_{\downarrow}^n$ .

**Type-safety** Proof of the type-safety (or soundness) of  $\lambda C_{\text{exp}}$  is fairly standard by subject reduction (or preservation) and progress lemmas. The subject reduction proof relies on the substitution lemma. We give the proof sketch of related lemmas as follows:

$e, \tau, \sigma$	$::=$	Expressions
	$\begin{array}{c} \vdots \\ \dots \\ \mu x : \tau. e \end{array}$	Polymorphic recursion
$e \longrightarrow e'$		One-step reduction
	$\frac{}{\mu x : \tau. e \longrightarrow e[x \mapsto \mu x : \tau. e]}$	S_MU
$\Gamma \vdash e : \tau$		Expression typing
	$\frac{\Gamma, x : \tau \vdash e : \tau \quad \Gamma \vdash \tau : \star}{\Gamma \vdash (\mu x : \tau. e) : \tau}$	T_MU

**Figure 5.** Grammar changes for general recursion

**Lemma 4.5** (Substitution lemma). *If  $\Gamma_1, x : \sigma, \Gamma_2 \vdash e_1 : \tau$  and  $\Gamma_1 \vdash e_2 : \sigma$ , then  $\Gamma_1, \Gamma_2[x \mapsto e_2] \vdash e_1[x \mapsto e_2] : \tau[x \mapsto e_2]$ .*

*Proof.* By induction on the derivation of  $\Gamma_1, x : \sigma, \Gamma_2 \vdash e_1 : \tau$ .  $\square$

**Lemma 4.6** (Subject reduction). *If  $\Gamma \vdash e : \sigma$  and  $e \longrightarrow e'$  then  $\Gamma \vdash e' : \sigma$ .*

*Proof. (Sketch)* We prove the case for one-step reduction, i.e.  $e \longrightarrow e'$ . The lemma can follow by induction on the number of one-step reductions of  $e \longrightarrow e'$ . The proof is by induction with respect to the definition of one-step reduction  $\longrightarrow$ .  $\square$

**Lemma 4.7** (Progress). *If  $e : \sigma$  then either  $e$  is a value  $v$  or there exists  $e'$  such that  $e \longrightarrow e'$ .*

*Proof.* By induction on the derivation of  $\vdash e : \sigma$ .  $\square$

## 5. $\lambda_\star$ : $\lambda C_{\text{exp}}$ with general recursion

In this section, we present the core calculus  $\lambda_\star$ , which is the explicit core  $\lambda C_{\text{exp}}$  extended with general recursion. The general recursion is polymorphic and has a uniform representation on both term level and type level, which works as a fixpoint and recursive type respectively. We have shown that  $\lambda C_{\text{exp}}$  does not rely on normalization for decidable type checking and type-safety. Though general recursion breaks the normalizing property, it is safe for  $\lambda_\star$  to allow non-termination on the type level without breaking the decidability of type checking.

### 5.1 Grammar changes

Figure 5 shows the changes of extending  $\lambda C_{\text{exp}}$  to  $\lambda_\star$  with general recursion. The changes are subtle, since we add only one more primitive, reduction rule and typing rule for general recursion.

For syntax, we add the polymorphic recursion operator  $\mu$  to represent general recursion on both term and type level in the same form  $\mu x : \tau. e$ . For operational semantics, we add the rule S\_MU to define the unrolling operation of a recursion, which results in  $e[x \mapsto \mu x : \tau. e]$ . For typing, we add the rule T\_MU for checking the validity of a polymorphic recursive term. The rule ensures that the polymorphic recursion  $\mu x : \tau. e$  should have the same type  $\tau$  as the binder  $x$  and also the inner expression  $e$ .

### 5.2 Recursion as term and type

**Term-level recursion** Note that the definition of values is not changed and  $\mu$ -term is not treated as a value. This is different from conventional iso-recursive types, because the  $\mu$ -operator can also be regarded as a *fixpoint* on the term level. By rule S\_MU,

evaluating a term  $\mu x : \tau. e$  will substitute all  $x$  in  $e$  with the whole  $\mu$ -term itself, resulting in the unrolling  $e[x \mapsto \mu x : \tau. e]$ . So the  $\mu$ -term is equivalent to a recursive function that should be allowed to unroll without restriction (see example in §??).

**Type-level recursion** On the type level,  $\mu x : \tau. e$  works as a *iso-recursive* type, a kind of recursive type that is not equal but only isomorphic to its unrolling. Normally, we need to add two more primitives fold and unfold for the iso-recursive type to map back and forth between the original and unrolled form. However, in  $\lambda_\star$  we can show that with the rule S\_MU,  $\text{cast}^\uparrow$  and  $\text{cast}_\downarrow$  have the same functionality as fold and unfold.

Assume there exist expressions  $e_1$  and  $e_2$  such that

$$\begin{array}{l} e_1 : \mu x : \tau. e \\ e_2 : e[x \mapsto \mu x : \tau. e] \end{array}$$

Note that  $e_1$  and  $e_2$  have distinct types but the type of  $e_2$  is the unrolling of  $e_1$ 's type, which follows the one-step reduction relation by rule S\_MU:

$$\mu x : \tau. e \longrightarrow e[x \mapsto \mu x : \tau. e]$$

By applying rules T\_CASTUP and T\_CASTDOWN, we can obtain the following typing results:

$$\begin{array}{l} \text{cast}_\downarrow e_1 : e[x \mapsto \mu x : \tau. e] \\ \text{cast}^\uparrow [\mu x : \tau. e] e_2 : \mu x : \tau. e \end{array}$$

Thus,  $\text{cast}^\uparrow$  and  $\text{cast}_\downarrow$  witness the isomorphism between the original recursive type and its unrolling, which behave the same as fold and unfold operations in iso-recursive types.

### 5.3 Decidability of type checking

We give the proof of the decidability of type checking  $\lambda C_{\text{exp}}$  in §4.4 without requiring normalization. **Generalize the proof for recursion.**

## 6. Surface language

BRUNO: Jeremy, I think you should write up this section.

- Expand the core language with datatypes and pattern matching by encoding.
- Give translation rules.
- Encode GADTs and maybe other Haskell extensions? GADTs seems challenging, so perhaps some other examples would be datatypes like *Fixf*, and *Monad* as a record. Could formalize records in Haskell style.

In this section, we present the surface language ( $\lambda C_{\text{suf}}$ ) that supports simple datatypes and case analysis. Due to the expressiveness of  $\lambda_\star$ , all these features can be elaborated into the core language without extending the built-in language constructs of  $\lambda_\star$ . In what follows, we first give the syntax of  $\lambda C_{\text{suf}}$ , followed by the extended typing rules, then we show the formal translation rules that translates  $\lambda C_{\text{suf}}$  expressions into  $\lambda_\star$  expressions. Finally we demonstrate the translation using a simple example.

### 6.1 Extended Syntax

The syntax of  $\lambda C_{\text{suf}}$  is shown in Figure 6. Compared with  $\lambda_\star$ ,  $\lambda C_{\text{suf}}$  has a new syntax category: a program, consisting of a list of datatype declarations, followed by an expression. An *algebraic data type*  $D$  is introduced as a top-level **data** declaration with its *data constructors*. The type of a data constructor  $K$  has the form:

$$K : \Pi \bar{u} : \bar{\kappa}^n. \Pi \bar{x} : \bar{\tau} \rightarrow D \bar{u}^n$$

The first  $n$  quantified type variables  $\bar{u}$  appear in the same order in the return type  $D \bar{u}$ . **Note that the use of  $\Pi$  to tie together the data**

constructor arguments makes it possible to let the types of some data constructor arguments depend on other data constructor arguments. The **case** expression is conventional, used to break up values built with data constructors. The patterns of a case expression are flat (no nested patterns), and bind value variables.

<b>Declarations</b>		
$pgm$	$::= \overline{decl}; e$	Declarations
$decl$	$::= \mathbf{data} \, D \, \overline{u} : \overline{\kappa} = \overline{K} \, \overline{\tau}$	Datatype
<b>Terms</b>		
$u$	$::= x \mid K$	Variables and constructors
$e, \tau, \sigma, v, \kappa$	$::= u$	Term atoms
	$\dots$	
$p$	$::= \begin{array}{l} \mid \mathbf{case} \, e \, \mathbf{of} \, \overline{p} \Rightarrow \overline{e} \\ \mid K \, \overline{x} : \overline{\tau} \end{array}$	Case analysis Pattern
<b>Environments</b>		
$\Gamma$	$::= \emptyset$	Empty
	$\mid \Gamma, u : \tau$	Variable binding

**Figure 6.** Syntax of  $\lambda C_{\text{suf}}$  ( $e$  for terms;  $\tau, \sigma, v$  for types;  $\kappa$  for kinds)

With datatypes, it is easy to encode *records* as syntactic sugar of simple datatypes, as shown in Figure 7.

$$\begin{aligned} \mathbf{data} \, R \, \overline{u} : \overline{\kappa} &= K \{ \overline{S} : \overline{\tau} \} \triangleq \\ \mathbf{data} \, R \, \overline{u} : \overline{\kappa} &= K \, \overline{\tau} \\ \mathbf{let} \, S_i : \Pi \overline{u} : \overline{\kappa}. R \, \overline{u} \rightarrow \tau_i &= \\ \lambda(u : \kappa). \lambda l : R \, \overline{u}. \mathbf{case} \, l \, \mathbf{of} \, K \, \overline{x} : \overline{\tau} &\Rightarrow x_i \\ \mathbf{in} \end{aligned}$$

**Figure 7.** Syntactic sugar for records

## 6.2 Extended Typing Rules

The type system of  $\lambda C_{\text{suf}}$  is shown in Figure 8. To save space, we only show the new typing rules. Furthermore, we sometimes adopt the following syntactic convention:

$$\overline{\tau}^n \rightarrow \tau_r \equiv \tau_1 \rightarrow \dots \rightarrow \tau_n \rightarrow \tau_r$$

Rule (Pgm) type-checks a whole problem. It first type-checks the declarations, which in return gives a new typing environment. Combined with the original environment, it then checks the expression and return the result type. Rule (Data) type-checks datatype declarations by ensuring the well-formedness of the kinds of type constructors and the types of data constructors. Finally rule (Alt) validates the patterns by looking up the the existence of corresponding data constructors in the typing environment, replacing universally quantified type variables with proper concrete types.

## 6.3 Translation Overview

We use a type-directed translation. The typing relations have the form:

$$\Gamma \vdash e : \tau \rightsquigarrow E$$

It states that  $\lambda_*$  expression  $E$  is the translation of  $\lambda C_{\text{suf}}$  expression  $e$  of type  $\tau$ . Figure 9 shows the translation rules, which are the typing rules in Figure 8 extended with the resulting expression  $E$ . In the translation, We require that applications of constructors to be *saturated*.

Among others, Rules (Case), (Alt) and (Data) are of the essence for the translation. Rule (Case) translates case expressions into applications by first type-converting the scrutinee expression, then applying it to the result type and a  $\lambda_*$  expression. Rule (Alt) translate

each pattern into a lambda expression, with each variable in the pattern corresponding to a variable in the lambda expression in the same order. The body in the alternative is recursively translated and taken as the lambda body.

Rule (Data) does the most heavy work and deserves further explanation. First of all, it results in a incomplete expression (as can be seen by the incomplete *let* expressions). The result expression is supposed to be prepended to the translation of the last expression to form a complete  $\lambda_*$  expression, as specified by Rule (Pgm). Furthermore, each type constructor is translated as a lambda expression, with a recursive type as the body. Each data constructor is also translated as a lambda expression. Notice that we use cast operation in the lambda body to restore to the corresponding datatype.

The rest of the translation rules hold few surprises.

## 7. Related Work

- Henk [7] and one of its implementation [9] show the simplicity of the Pure Type System (PTS). [10] also tries to combine recursion with PTS.
- Zombie [3, 11] is a language with two fragments supporting logics with non-termination. It limits the  $\beta$ -reduction for congruence closure [12].
- $\Pi\Sigma$  [1] is a simple, dependently-typed core language for expressing high-level constructions<sup>1</sup>. UHC compiler [8] tries to use a simplified core language with coercion to encode GADTs.
- System  $F_C$  [13] has been extended with type promotion [16] and kind equality [15]. The latter one introduces a limited form of dependent types into the system<sup>2</sup>, which mixes up types and kinds.

## 8. Conclusion

Conclusion and related work.

## Acknowledgments

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<sup>1</sup> But the paper didn't give any meta-theories about the language.

<sup>2</sup> Richard A. Eisenberg is going to implement kind equality [15] into GHC. The implementation is proposed at <https://phabricator.haskell.org/D808> and related paper is at <http://www.cis.upenn.edu/~eir/papers/2015/equalities/equalities-extended.pdf>.



$\Gamma \vdash \text{pgm} : \tau$	
(Pgm)	$\frac{\overline{\Gamma_0 \vdash \text{decl} : \Gamma_d} \quad \Gamma = \Gamma_0, \overline{\Gamma_d} \quad \Gamma \vdash e : \tau}{\Gamma_0 \vdash \overline{\text{decl}}; e : \tau}$
$\Gamma \vdash \text{decl} : \Gamma_d$	
(Data)	$\frac{\Gamma \vdash \bar{\kappa} \rightarrow \star : \square \quad \overline{\Gamma, D : \bar{\kappa} \rightarrow \star, \bar{u} : \bar{\kappa} \vdash \bar{\tau} \rightarrow D \bar{u} : \star}}{\Gamma \vdash (\text{data } D \bar{u} : \bar{\kappa} = \mid K \bar{\tau}) : (D : \bar{\kappa} \rightarrow \star, \bar{K} : \Pi \bar{u} : \bar{\kappa}. \bar{\tau} \rightarrow D \bar{u})}$
$\Gamma \vdash e : \tau$	
(Case)	$\frac{\Gamma \vdash e_1 : \sigma \quad \overline{\Gamma \vdash_p p \Rightarrow e_2 : \sigma \rightarrow \tau}}{\Gamma \vdash \text{case } e_1 \text{ of } \bar{p} \Rightarrow \bar{e}_2 : \tau}$
$\Gamma \vdash_p p \Rightarrow e : \sigma \rightarrow \tau$	
(Alt)	$\frac{\theta = [\bar{u} := \bar{v}] \quad \overline{K : \Pi \bar{u} : \bar{\kappa}. \bar{\sigma} \rightarrow D \bar{u} \in \Gamma} \quad \overline{\Gamma, x : \theta(\sigma) \vdash e : \tau}}{\Gamma \vdash_p K x : \theta(\sigma) \Rightarrow e : D \bar{v} \rightarrow \tau}$

Figure 8. Typing rules of  $\lambda C_{\text{suf}}$

$\Gamma \vdash e : \tau \rightsquigarrow E$	
(Ax)	$\frac{}{\emptyset \vdash \star : \square \rightsquigarrow \star}$
(Var)	$\frac{x : \tau \in \Gamma}{\Gamma \vdash x : \tau \rightsquigarrow x}$
(App)	$\frac{\Gamma \vdash e_1 : (\Pi x : \tau_2. \tau_1) \rightsquigarrow E_1 \quad \Gamma \vdash e_2 : \tau_2 \rightsquigarrow E_2}{\Gamma \vdash e_1 e_2 : \tau_1[x := e_2] \rightsquigarrow E_1 E_2}$
(Lam)	$\frac{\Gamma, x : \tau_1 \vdash e : \tau_2 \rightsquigarrow E \quad \Gamma \vdash (\Pi x : \tau_1. \tau_2) : t}{\Gamma \vdash (\lambda x : \tau_1. e) : (\Pi x : \tau_1. \tau_2) \rightsquigarrow \lambda x : \tau_1. E} \quad t \in \{\star, \square\}$
(Pi)	$\frac{\Gamma \vdash \tau_1 : s \quad \Gamma, x : \tau_1 \vdash \tau_2 : t}{\Gamma \vdash (\Pi x : \tau_1. \tau_2) : t \rightsquigarrow \Pi x : \tau_1. \tau_2} \quad (s, t) \in \mathcal{R}$
(Mu)	$\frac{\Gamma, x : \tau \vdash e : \tau \rightsquigarrow E \quad \Gamma \vdash \tau : s}{\Gamma \vdash (\mu x : \tau. e) : \tau \rightsquigarrow \mu x : \tau. E} \quad s \in \{\star, \square\}$
(Fold)	$\frac{\Gamma \vdash e : \tau_2 \rightsquigarrow E \quad \Gamma \vdash \tau_1 : s \quad \tau_1 \longrightarrow \tau_2}{\Gamma \vdash (\text{cast}^\uparrow_{[\tau_1]} e) : \tau_1 \rightsquigarrow \text{cast}^\uparrow_{[\tau_1]} E}$
(Unfold)	$\frac{\Gamma \vdash e : \tau_1 \rightsquigarrow E \quad \Gamma \vdash \tau_2 : s \quad \tau_1 \longrightarrow \tau_2}{\Gamma \vdash (\text{cast}_\downarrow e) : \tau_2 \rightsquigarrow \text{cast}_\downarrow E}$
(Case)	$\frac{\Gamma \vdash e_1 : \sigma \rightsquigarrow E_1 \quad \overline{\Gamma \vdash_p p \Rightarrow e_2 : \sigma \rightarrow \tau \rightsquigarrow E_2}}{\Gamma \vdash \text{case } e_1 \text{ of } \bar{p} \Rightarrow \bar{e}_2 : \tau \rightsquigarrow (\text{cast}_\downarrow E_1) \tau \bar{E}_2}$
$\Gamma \vdash_p p \Rightarrow e : \sigma \rightarrow \tau \rightsquigarrow E$	
(Alt)	$\frac{\theta = [\bar{u} := \bar{v}]}{K : \Pi \bar{u} : \bar{\kappa}. \bar{\sigma} \rightarrow D \bar{u} \in \Gamma \quad \overline{\Gamma, x : \theta(\sigma) \vdash e : \tau \rightsquigarrow E}} \quad \Gamma \vdash_p K x : \theta(\sigma) \Rightarrow e : D \bar{v} \rightarrow \tau \rightsquigarrow \lambda(x : \theta(\sigma)). E$
$\Gamma \vdash \text{decl} : \Gamma_d \rightsquigarrow E$	
(Data)	$\frac{\Gamma \vdash \bar{\kappa} \rightarrow \star : \square \quad \overline{\Gamma, D : \bar{\kappa} \rightarrow \star, \bar{u} : \bar{\kappa} \vdash \bar{\tau} \rightarrow D \bar{u} : \star}}{\Gamma \vdash (\text{data } D \bar{u} : \bar{\kappa} = \mid K \bar{\tau}) : (D : \bar{\kappa} \rightarrow \star, \bar{K} : \Pi \bar{u} : \bar{\kappa}. \bar{\tau} \rightarrow D \bar{u}) \rightsquigarrow E}$
$E ::=$	$\begin{aligned} & \text{let } D : \bar{\kappa} \rightarrow \star = \lambda \bar{u} : \bar{\kappa}. \mu X : \star. \Pi b : \star. (\bar{\tau}[D \bar{u} := X] \rightarrow b) \rightarrow b \text{ in} \\ & \text{let } K_i : \Pi \bar{u} : \bar{\kappa}. \bar{\tau} \rightarrow D \bar{u} = \lambda(u : \kappa). \lambda(x : \tau). \\ & \quad \text{cast}^\uparrow_{[D \bar{u}]} (\lambda(b : \star) (c : \bar{\tau} \rightarrow b). c_i \bar{x}) \text{ in} \end{aligned}$
$\Gamma \vdash \text{pgm} : \tau \rightsquigarrow E$	
(Pgm)	$\frac{\overline{\Gamma_0 \vdash \text{decl} : \Gamma_d \rightsquigarrow E_1} \quad \Gamma = \Gamma_0, \overline{\Gamma_d} \quad \Gamma \vdash e : \tau \rightsquigarrow E}{\Gamma_0 \vdash \overline{\text{decl}}; e : \tau \rightsquigarrow E_1 \oplus E}$

Figure 9. Type-directed translation from  $\lambda C_{\text{suf}}$  to  $\lambda_\star$

ISBN 0897917693. . URL <http://dl.acm.org/citation.cfm?id=237721.237792>.

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## A. Full specification of core language

### A.1 Syntax

$e, \tau, \sigma$	$::=$	Expressions
$x$		Variable
$\star$		Type of types
$e_1 e_2$		Application
$\lambda x : \tau. e$		Abstraction
$\Pi x : \tau. \sigma$		Product
$\text{cast}^\uparrow[\tau] e$		Cast up by expansion
$\text{cast}_\downarrow e$		Cast down by reduction
$\mu x : \tau. e$		General recursion
$\Gamma$	$::=$	Contexts
$\emptyset$		Empty
$\Gamma, x : \tau$		Variable binding
$v$	$::=$	Values
$\star$		Type of types
$\lambda x : \tau. e$		Abstraction
$\Pi x : \tau_1. \tau_2$		Product
$\text{cast}^\uparrow[\tau] e$		Cast up by expansion

### A.2 Operational semantics

$e \longrightarrow e'$  One-step reduction

$$\frac{}{(\lambda x : \tau. e_1) e_2 \longrightarrow e_1[x \mapsto e_2]} \text{S\_BETA}$$

$$\frac{}{\text{cast}_\downarrow(\text{cast}^\uparrow[\tau] e) \longrightarrow e} \text{S\_CASTDOWNUP}$$

$$\frac{e_1 \longrightarrow e'_1}{e_1 e_2 \longrightarrow e'_1 e_2} \text{S\_APP}$$

$$\frac{e \longrightarrow e'}{\text{cast}_\downarrow e \longrightarrow \text{cast}_\downarrow e'} \text{S\_CASTDOWN}$$

$$\frac{}{\mu x : \tau. e \longrightarrow e[x \mapsto \mu x : \tau. e]} \text{S\_MU}$$

### A.3 Expression typing

$\Gamma \vdash e : \tau$  Expression typing

$$\frac{}{\emptyset \vdash \star : \star} \text{T\_AX}$$

$$\frac{\vdash \Gamma \quad x : \tau \in \Gamma}{\Gamma \vdash x : \tau} \text{T\_VAR}$$

$$\frac{\Gamma \vdash e_1 : (\Pi x : \tau_2. \tau_1) \quad \Gamma \vdash e_2 : \tau_2}{\Gamma \vdash e_1 e_2 : \tau_1[x \mapsto e_2]} \text{T\_APP}$$

$$\frac{\Gamma, x : \tau_1 \vdash e : \tau_2 \quad \Gamma \vdash (\Pi x : \tau_1. \tau_2) : \star}{\Gamma \vdash (\lambda x : \tau_1. e) : (\Pi x : \tau_1. \tau_2)} \text{T\_LAM}$$

$$\frac{\Gamma \vdash \tau_1 : \star \quad \Gamma, x : \tau_1 \vdash \tau_2 : \star}{\Gamma \vdash (\Pi x : \tau_1. \tau_2) : \star} \text{T\_PI}$$

$$\frac{\Gamma \vdash e : \tau_2 \quad \Gamma \vdash \tau_1 : \star \quad \tau_1 \longrightarrow \tau_2}{\Gamma \vdash (\text{cast}^\uparrow[\tau_1] e) : \tau_1} \text{T\_CASTUP}$$

$$\frac{\Gamma \vdash e : \tau_1 \quad \Gamma \vdash \tau_2 : \star \quad \tau_1 \longrightarrow \tau_2}{\Gamma \vdash (\text{cast}_\downarrow e) : \tau_2} \text{T\_CASTDOWN}$$

$$\frac{\Gamma, x : \tau \vdash e : \tau \quad \Gamma \vdash \tau : \star}{\Gamma \vdash (\mu x : \tau. e) : \tau} \text{T\_MU}$$

## B. Proofs about core language

### B.1 Properties

**Lemma B.1** (Free variable lemma). *If  $\Gamma \vdash e : \tau$ , then  $\text{FV}(e) \subseteq \text{dom}(\Gamma)$  and  $\text{FV}(\tau) \subseteq \text{dom}(\Gamma)$ .*

*Proof.* By induction on the derivation of  $\Gamma \vdash e : \tau$ . We only treat cases T\_MU, T\_CASTUP and T\_CASTDOWN (since proofs of other cases are the same as  $\lambda C$  [2]):

**Case T\_MU:** From premises of  $\Gamma \vdash (\mu x : \tau. e_1) : \tau$ , by induction hypothesis, we have  $\text{FV}(e_1) \subseteq \text{dom}(\Gamma) \cup \{x\}$  and  $\text{FV}(\tau) \subseteq \text{dom}(\Gamma)$ . Thus the result follows by  $\text{FV}(\mu x : \tau. e_1) = \text{FV}(e_1) \setminus \{x\} \subseteq \text{dom}(\Gamma)$  and  $\text{FV}(\tau) \subseteq \text{dom}(\Gamma)$ .

**Case T\_CASTUP:** Since  $\text{FV}(\text{cast}^\uparrow[\tau] e_1) = \text{FV}(e_1)$ , the result follows directly by the induction hypothesis.

**Case T\_CASTDOWN:** Since  $\text{FV}(\text{cast}_\downarrow e_1) = \text{FV}(e_1)$ , the result follows directly by the induction hypothesis.  $\square$

**Lemma B.2** (Substitution lemma). *If  $\Gamma_1, x : \sigma, \Gamma_2 \vdash e_1 : \tau$  and  $\Gamma_1 \vdash e_2 : \sigma$ , then  $\Gamma_1, \Gamma_2[x \mapsto e_2] \vdash e_1[x \mapsto e_2] : \tau[x \mapsto e_2]$ .*

*Proof.* By induction on the derivation of  $\Gamma_1, x : \sigma, \Gamma_2 \vdash e_1 : \tau$ . Let  $e^* \equiv e[x \mapsto e_2]$ . Then the result can be written as  $\Gamma_1, \Gamma_2^* \vdash e_1^* : \tau^*$ . We only treat cases T\_MU, T\_CASTUP and T\_CASTDOWN. Consider the last step of derivation of the following cases:

$$\text{Case T\_MU: } \frac{\Gamma_1, x : \sigma, \Gamma_2 \vdash e_1 : \tau \quad \Gamma_1, x : \sigma, \Gamma_2 \vdash \tau : \star}{\Gamma_1, x : \sigma, \Gamma_2 \vdash (\mu y : \tau. e_1) : \tau}$$

By induction hypothesis, we have  $\Gamma_1, \Gamma_2^* \vdash e_1^* : \tau^*$  and  $\Gamma_1, \Gamma_2^* \vdash \tau^* : \star$ . Then by the deviation rule,  $\Gamma_1, \Gamma_2^* \vdash (\mu y : \tau^*. e_1^*) : \tau^*$ . Thus we have  $\Gamma_1, \Gamma_2^* \vdash (\mu y : \tau. e_1)^* : \tau^*$  which is just the result.

$$\text{Case T\_CASTUP: } \frac{\Gamma_1, x : \sigma, \Gamma_2 \vdash e_1 : \tau_2 \quad \Gamma_1, x : \sigma, \Gamma_2 \vdash \tau_1 : \star \quad \tau_1 \longrightarrow \tau_2}{\Gamma_1, x : \sigma, \Gamma_2 \vdash (\text{cast}^\uparrow[\tau_1]e_1) : \tau_1}$$

By induction hypothesis, we have  $\Gamma_1, \Gamma_2^* \vdash e_1^* : \tau_2^*$ ,  $\Gamma_1, \Gamma_2^* \vdash \tau_1^* : \star$  and  $\tau_1 \longrightarrow \tau_2$ . By the definition of substitution, we can obtain  $\tau_1^* \longrightarrow \tau_2^*$  by  $\tau_1 \longrightarrow \tau_2$ . Then by the deviation rule,  $\Gamma_1, \Gamma_2^* \vdash (\text{cast}^\uparrow[\tau_1^*]e_1^*) : \tau_1^*$ . Thus we have  $\Gamma_1, \Gamma_2^* \vdash (\text{cast}^\uparrow[\tau_1]e_1)^* : \tau_1^*$  which is just the result.

$$\text{Case T\_CASTDOWN: } \frac{\Gamma_1, x : \sigma, \Gamma_2 \vdash e_1 : \tau_1 \quad \Gamma_1, x : \sigma, \Gamma_2 \vdash \tau_2 : \star \quad \tau_1 \longrightarrow \tau_2}{\Gamma_1, x : \sigma, \Gamma_2 \vdash (\text{cast}_\downarrow e_1) : \tau_2}$$

By induction hypothesis, we have  $\Gamma_1, \Gamma_2^* \vdash e_1^* : \tau_1^*$ ,  $\Gamma_1, \Gamma_2^* \vdash \tau_2^* : \star$  and  $\tau_1 \longrightarrow \tau_2$  thus  $\tau_1^* \longrightarrow \tau_2^*$ . Then by the deviation rule,  $\Gamma_1, \Gamma_2^* \vdash (\text{cast}_\downarrow e_1^*) : \tau_2^*$ . Thus we have  $\Gamma_1, \Gamma_2^* \vdash (\text{cast}_\downarrow e_1)^* : \tau_2^*$  which is just the result.  $\square$

**Lemma B.3** (Generation lemma).

- (1) If  $\Gamma \vdash x : \sigma$ , then there exist an expression  $\tau$  such that  $\tau \equiv \sigma$ ,  $\Gamma \vdash \tau : \star$  and  $x : \tau \in \Gamma$ .
- (2) If  $\Gamma \vdash e_1 e_2 : \sigma$ , then there exist expressions  $\tau_1$  and  $\tau_2$  such that  $\Gamma \vdash e_1 : (\Pi x : \tau_1. \tau_2)$ ,  $\Gamma \vdash e_2 : \tau_1$  and  $\sigma \equiv \tau_2[x \mapsto e_2]$ .
- (3) If  $\Gamma \vdash (\lambda x : \tau_1. e) : \sigma$ , then there exist an expression  $\tau_2$  such that  $\sigma \equiv \Pi x : \tau_1. \tau_2$  where  $\Gamma \vdash (\Pi x : \tau_1. \tau_2) : \star$  and  $\Gamma, x : \tau_1 \vdash e : \tau_2$ .
- (4) If  $\Gamma \vdash (\Pi x : \tau_1. \tau_2) : \sigma$ , then  $\sigma \equiv \star$ ,  $\Gamma \vdash \tau_1 : \star$  and  $\Gamma, x : \tau_1 \vdash \tau_2 : \star$ .
- (5) If  $\Gamma \vdash (\mu x : \tau. e) : \sigma$ , then  $\Gamma \vdash \tau : \star$ ,  $\sigma \equiv \tau$  and  $\Gamma, x : \tau \vdash e : \tau$ .
- (6) If  $\Gamma \vdash (\text{cast}^\uparrow[\tau_1]e) : \sigma$ , then there exist an expression  $\tau_2$  such that  $\Gamma \vdash e : \tau_2$ ,  $\Gamma \vdash \tau_1 : \star$ ,  $\tau_1 \longrightarrow \tau_2$  and  $\sigma \equiv \tau_1$ .
- (7) If  $\Gamma \vdash (\text{cast}_\downarrow e) : \sigma$ , then there exist expressions  $\tau_1, \tau_2$  such that  $\Gamma \vdash e : \tau_1$ ,  $\Gamma \vdash \tau_2 : \star$ ,  $\tau_1 \longrightarrow \tau_2$  and  $\sigma \equiv \tau_2$ .

*Proof.* Consider a derivation of  $\Gamma \vdash e : \sigma$  for one of cases in the lemma. Note that rule T.WEAK does not change  $e$ , then we can follow the process of derivation until expression  $e$  is introduced the first time. The last step of derivation can be done by

- rule T\_VAR for case 1;
- rule T\_APP for case 2;
- rule T\_LAM for case 3;
- rule T\_PI for case 4;
- rule T\_MU for case 5;
- rule T\_CASTUP for case 6;
- rule T\_CASTDOWN for case 7.

In each case, assume the conclusion of the rule is  $\Gamma' \vdash e : \tau'$  where  $\Gamma' \subseteq \Gamma$  and  $\tau' \equiv \sigma$ . Then by inspection of used derivation rules, it can be shown that the statement of the lemma holds and is the only possible case.  $\square$

**Lemma B.4** (Correctness of types). If  $\Gamma \vdash e : \tau$  then  $\tau \equiv \star$  or  $\Gamma \vdash \tau : \star$ .

*Proof.* Trivial induction on the derivation of  $\Gamma \vdash e : \tau$  using Lemma B.3.  $\square$

## B.2 Decidability of type checking

**Lemma B.5** (Uniqueness of one-step reduction). The relation  $\longrightarrow$ , i.e. one-step reduction, is unique in the sense that given  $e$  there is at most one  $e'$  such that  $e \longrightarrow e'$ .

*Proof.* By induction on the structure of  $e$ :

**Case  $e = v$ :**  $e$  has one of the following forms: (1)  $\star$ , (2)  $\lambda x : \tau. e$ , (3)  $\Pi x : \tau_1. \tau_2$ , (4)  $\text{cast}^\uparrow[\tau]e$ , which cannot match any rules of  $\longrightarrow$ . Thus there is no  $e'$  such that  $e \longrightarrow e'$ .

**Case  $e = (\lambda x : \tau. e_1) e_2$ :** There is a unique  $e' = e_1[x \mapsto e_2]$  by rule S\_BETA.

**Case  $e = \text{cast}_\downarrow(\text{cast}^\uparrow[\tau]e)$ :** There is a unique  $e' = e$  by rule S\_CASTDOWNUP.

**Case  $e = \mu x : \tau. e$ :** There is a unique  $e' = e[x \mapsto \mu x : \tau. e]$  by rule S\_MU.

**Case  $e = e_1 e_2$  and  $e_1$  is not a  $\lambda$ -term:** If  $e_1 = v$ , there is no  $e'_1$  such that  $e_1 \longrightarrow e'_1$ . Since  $e_1$  is not a  $\lambda$ -term, there is no rule to reduce  $e$ . Thus there is no  $e'$  such that  $e \longrightarrow e'$ .

Otherwise, there exists some  $e'_1$  such that  $e_1 \longrightarrow e'_1$ . By the induction hypothesis,  $e'_1$  is unique reduction of  $e_1$ . Thus by rule S\_APP,  $e' = e'_1 e_2$  is the unique reduction for  $e$ .

**Case  $e = \text{cast}_\downarrow e_1$  and  $e_1$  is not a  $\text{cast}^\uparrow$ -term:** If  $e_1 = v$ , there is no  $e'_1$  such that  $e_1 \longrightarrow e'_1$ . Since  $e_1$  is not a  $\text{cast}^\uparrow$ -term, there is no rule to reduce  $e$ . Thus there is no  $e'$  such that  $e \longrightarrow e'$ .

Otherwise, there exists some  $e'_1$  such that  $e_1 \longrightarrow e'_1$ . By the induction hypothesis,  $e'_1$  is unique reduction of  $e_1$ . Thus by rule S\_CASTDOWN,  $e' = \text{cast}_\downarrow e'_1$  is the unique reduction for  $e$ .  $\square$

**Lemma B.6** (Uniqueness of  $n$ -step reduction). The  $n$ -step reduction  $\longrightarrow_n$  is unique in the sense that given  $e$  there is at most one  $e'$  such that  $e \longrightarrow_n e'$ .

*Proof.* Immediate from Lemma B.5, by induction on the number of reduction steps.  $\square$

**Lemma B.7** (Decidability of type checking). There is a decidable algorithm which given  $\Gamma, e$  computes the unique  $\tau$  such that  $\Gamma \vdash e : \tau$  or reports there is no such  $\tau$ .

*Proof.* By induction on the structure of  $e$ :

**Case  $e = \star$ :** Trivial by applying T\_AX and  $\tau \equiv \star$ .

**Case  $e = x$ :** Trivial by rule T\_S\_VAR and  $\tau$  is the unique type of  $x$  if  $x : \tau \in \Gamma$ .

**Case  $e = e_1 e_2$ , or  $\lambda x : \tau_1. e_1$ , or  $\Pi x : \tau_1. \tau_2$ , or  $\mu x : \tau. e_1$ :** Trivial according to Lemma B.3 by using rule T\_APP, T\_LAM, T\_PI, or T\_MU respectively.

**Case  $e = \text{cast}^\uparrow[\tau_1]e_1$ :** From the premises of rule T\_CASTUP, by induction hypothesis, we can derive the type of  $e_1$  as  $\tau_2$ , and check whether  $\tau_1$  is legal, i.e. its sorts is  $\star$ . If  $\tau_1$  is legal, by Lemma B.5, there is at most one  $\tau'_1$  such that  $\tau_1 \longrightarrow \tau'_1$ . If such  $\tau'_1$  does not exist, then we report the type checking is failed. Otherwise, we examine if  $\tau'_1$  is syntactically equal to  $\tau_2$ , i.e.  $\tau'_1 \equiv \tau_2$ . If the equality holds, we obtain the unique type of  $e$  which is  $\tau_1$ . Otherwise, we report  $e$  fails to type check.

**Case  $e = \text{cast}_\downarrow e_1$ :** From the premises of rule T\_CASTDOWN, by induction hypothesis, we can derive the type of  $e_1$  as  $\tau_1$ . By Lemma B.5, there is at most one  $\tau_2$  such that  $\tau_1 \longrightarrow \tau_2$ . If such  $\tau_2$  exists and its sorts is  $\star$ , we have found the unique type of  $e$  is  $\tau_2$ . Otherwise, we report  $e$  fails to type check.  $\square$

### B.3 Soundness

**Definition B.8** (Multi-step reduction). *The relation  $\rightarrow$  is the transitive and reflexive closure of  $\rightarrow$ .*

**Lemma B.9** (Subject reduction). *If  $\Gamma \vdash e : \sigma$  and  $e \rightarrow e'$  then  $\Gamma \vdash e' : \sigma$ .*

*Proof.* We prove the case for one-step reduction, i.e.  $e \rightarrow e'$ . The lemma can follow by induction on the number of one-step reductions of  $e \rightarrow e'$ . The proof is by induction with respect to the definition of one-step reduction  $\rightarrow$  as follows:

**Case**  $\frac{(\lambda x : \tau. e_1) e_2 \rightarrow e_1[x \mapsto e_2]}{\text{S\_BETA:}}$

Suppose  $\Gamma \vdash (\lambda x : \tau_1. e_1) e_2 : \sigma$  and  $\Gamma \vdash e_1[x \mapsto e_2] : \sigma'$ . By Lemma B.3(2), there exist expressions  $\tau'_1$  and  $\tau_2$  such that

$$\begin{aligned} \Gamma \vdash (\lambda x : \tau_1. e_1) : (\Pi x : \tau'_1. \tau_2) \quad (1) \\ \Gamma \vdash e_2 : \tau'_1 \\ \sigma \equiv \tau_2[x \mapsto e_2] \end{aligned}$$

By Lemma B.3(3), the judgement (1) implies that there exists an expression  $\tau'_2$  such that

$$\begin{aligned} \Pi x : \tau'_1. \tau_2 \equiv \Pi x : \tau_1. \tau'_2 \quad (2) \\ \Gamma, x : \tau_1 \vdash e_1 : \tau'_2 \end{aligned}$$

Hence, by (2) we have  $\tau_1 \equiv \tau'_1$  and  $\tau_2 \equiv \tau'_2$ . Then we can obtain  $\Gamma, x : \tau_1 \vdash e_1 : \tau_2$  and  $\Gamma \vdash e_2 : \tau_1$ . By Lemma B.2, we have  $\Gamma \vdash e_1[x \mapsto e_2] : \tau_2[x \mapsto e_2]$ . Therefore, we conclude with  $\sigma' \equiv \tau_2[x \mapsto e_2] \equiv \sigma$ .

**Case**  $\frac{e_1 \rightarrow e'_1}{e_1 e_2 \rightarrow e'_1 e_2} \text{S\_APP:}$

Suppose  $\Gamma \vdash e_1 e_2 : \sigma$  and  $\Gamma \vdash e'_1 e_2 : \sigma'$ . By Lemma B.3(2), there exist expressions  $\tau_1$  and  $\tau_2$  such that

$$\begin{aligned} \Gamma \vdash e_1 : (\Pi x : \tau_1. \tau_2) \\ \Gamma \vdash e_2 : \tau_1 \\ \sigma \equiv \tau_2[x \mapsto e_2] \end{aligned}$$

By induction hypothesis, we have  $\Gamma \vdash e'_1 : (\Pi x : \tau_1. \tau_2)$ . By rule T\_APP, we obtain  $\Gamma \vdash e'_1 e_2 : \tau_2[x \mapsto e_2]$ . Therefore,  $\sigma' \equiv \tau_2[x \mapsto e_2] \equiv \sigma$ .

**Case**  $\frac{e \rightarrow e'}{\text{cast}_{\downarrow} e \rightarrow \text{cast}_{\downarrow} e'} \text{S\_CASTDOWN:}$

Suppose  $\Gamma \vdash \text{cast}_{\downarrow} e : \sigma$  and  $\Gamma \vdash \text{cast}_{\downarrow} e' : \sigma'$ . By Lemma B.3(7), there exist expressions  $\tau_1, \tau_2$  such that

$$\begin{aligned} \Gamma \vdash e : \tau_1 \quad \Gamma \vdash \tau_2 : \star \\ \tau_1 \rightarrow \tau_2 \quad \sigma \equiv \tau_2 \end{aligned}$$

By induction hypothesis, we have  $\Gamma \vdash e' : \tau_1$ . By rule T\_CASTDOWN, we obtain  $\Gamma \vdash \text{cast}_{\downarrow} e' : \tau_2$ . Therefore,  $\sigma' \equiv \tau_2 \equiv \sigma$ .

**Case**  $\frac{\text{cast}_{\downarrow} (\text{cast}^{\uparrow} [\tau] e) \rightarrow e}{\text{S\_CASTDOWNUP:}}$

Suppose  $\Gamma \vdash \text{cast}_{\downarrow} (\text{cast}^{\uparrow} [\tau] e) : \sigma$  and  $\Gamma \vdash e : \sigma'$ . By Lemma B.3(7), there exist expressions  $\tau'_1, \tau_2$  such that

$$\begin{aligned} \Gamma \vdash (\text{cast}^{\uparrow} [\tau_1] e) : \tau'_1 \quad (3) \\ \tau'_1 \rightarrow \tau_2 \quad (4) \\ \sigma \equiv \tau_2 \quad (5) \end{aligned}$$

By Lemma B.3(6), the judgement (3) implies that there exists an expression  $\tau'_2$  such that

$$\Gamma \vdash e : \tau'_2 \quad (6)$$

$$\tau_1 \rightarrow \tau'_2 \quad (7)$$

$$\tau'_1 \equiv \tau_1 \quad (8)$$

By (4, 7, 8) and Lemma B.5 we obtain  $\tau_2 \equiv \tau'_2$ . From (6) we have  $\sigma' \equiv \tau'_2$ . Therefore, by (5),  $\sigma' \equiv \tau'_2 \equiv \tau_2 \equiv \sigma$ .

**Case**  $\frac{\mu x : \tau. e \rightarrow e[x \mapsto \mu x : \tau. e]}{\text{S\_MU:}}$

Suppose  $\Gamma \vdash (\mu x : \tau. e) : \sigma$  and  $\Gamma \vdash e[x \mapsto \mu x : \tau. e] : \sigma'$ . By Lemma B.3(5), we have  $\sigma \equiv \tau$  and  $\Gamma, x : \tau \vdash e : \tau$ . Then we obtain  $\Gamma \vdash (\mu x : \tau. e) : \tau$ . Thus by Lemma B.2, we have  $\Gamma \vdash e[x \mapsto \mu x : \tau. e] : \tau[x \mapsto \mu x : \tau. e]$ .

Note that  $x : \tau$ , i.e. the type of  $x$  is  $\tau$ , then  $x \notin \text{FV}(\tau)$  holds implicitly. Hence, by the definition of substitution, we obtain  $\tau[x \mapsto \mu x : \tau. e] \equiv \tau$ . Therefore,  $\sigma' \equiv \tau[x \mapsto \mu x : \tau. e] \equiv \tau \equiv \sigma$ .

□

**Lemma B.10** (Progress). *If  $\vdash e : \sigma$  then either  $e$  is a value  $v$  or there exists  $e'$  such that  $e \rightarrow e'$ .*

*Proof.* By induction on the derivation of  $\vdash e : \sigma$  as follows:

**Case**  $e = x$ : Impossible, because the context is empty.

**Case**  $e = v$ : Trivial, since  $e$  is already a value that has one of the following forms: (1)  $\star$ , (2)  $\lambda x : \tau. e$ , (3)  $\Pi x : \tau_1. \tau_2$ , (4)  $\text{cast}^{\uparrow} [\tau] e$ .

**Case**  $e = e_1 e_2$ : By Lemma B.3(2), there exist expressions  $\tau_1$  and  $\tau_2$  such that  $\vdash e_1 : (\Pi x : \tau_1. \tau_2)$  and  $\vdash e_2 : \tau_1$ . Consider whether  $e_1$  is a value:

- If  $e_1 = v$ , by Lemma B.3(3), it must be a  $\lambda$ -term such that  $e_1 \equiv \lambda x : \tau_1. e'_1$  for some  $e'_1$  satisfying  $\vdash e'_1 : \tau_2$ . Then by rule S\_BETA, we have  $(\lambda x : \tau_1. e'_1) e_2 \rightarrow e'_1[x \mapsto e_2]$ . Thus, there exists  $e' \equiv e'_1[x \mapsto e_2]$  such that  $e \rightarrow e'$ .
- Otherwise, by induction hypothesis, there exists  $e'_1$  such that  $e_1 \rightarrow e'_1$ . Then by rule S\_APP, we have  $e_1 e_2 \rightarrow e'_1 e_2$ . Thus, there exists  $e' \equiv e'_1 e_2$  such that  $e \rightarrow e'$ .

**Case**  $e = \text{cast}_{\downarrow} e_1$ : By Lemma B.3(7), there exist expressions  $\tau_1$  and  $\tau_2$  such that  $\vdash e_1 : \tau_1$  and  $\tau_1 \rightarrow \tau_2$ . Consider whether  $e_1$  is a value:

- If  $e_1 = v$ , by Lemma B.3(6), it must be a  $\text{cast}^{\uparrow}$ -term such that  $e_1 \equiv \text{cast}^{\uparrow} [\tau_1] e'_1$  for some  $e'_1$  satisfying  $\vdash e'_1 : \tau_2$ . Then by rule S\_CASTDOWNUP, we can obtain  $\text{cast}_{\downarrow} (\text{cast}^{\uparrow} [\tau_1] e'_1) \rightarrow e'_1$ . Thus, there exists  $e' \equiv e'_1$  such that  $e \rightarrow e'$ .
- Otherwise, by induction hypothesis, there exists  $e'_1$  such that  $e_1 \rightarrow e'_1$ . Then by rule S\_CASTDOWN, we have  $\text{cast}_{\downarrow} e_1 \rightarrow \text{cast}_{\downarrow} e'_1$ . Thus, there exists  $e' \equiv \text{cast}_{\downarrow} e'_1$  such that  $e \rightarrow e'$ .

**Case**  $e = \mu x : \tau. e_1$ : By rule S\_MU, there always exists  $e' \equiv e_1[x \mapsto \mu x : \tau. e_1]$ .

□

## C. Full specification of source language

### C.1 Syntax

See Figure 10.

### C.2 Expression typing

See Figure 11.

$pgm$	$::=$ $\mid \overline{decl}; E$	Program Declarations
$decl$	$::=$ $\mid \mathbf{data} \overline{Du : \kappa} = \overline{K\overline{T}}$	Declarations Datatype declaration
$u, v$	$::=$ $\mid x$ $\mid K$	Atoms Variable Data constructor
$p$	$::=$ $\mid K \overline{x : T}$	Patterns Pattern
$E, T, S, \kappa$	$::=$ $\mid u$ $\mid \star$ $\mid E_1 E_2$ $\mid \lambda x : T. E$ $\mid \Pi x : T_1. T_2$ $\mid \mu x : T. E$ $\mid \mathbf{case} E_1 \text{ of } \overline{p \Rightarrow E_2}$ $\mid \mathbf{let} x : T = E_2 \text{ in } E_1$	Expressions Atom Type of types Application Abstraction Product General recursion Case analysis Let binding
$\Sigma$	$::=$ $\mid \emptyset$ $\mid \Sigma, u : T$	Contexts Empty Atom binding
$T_1 \rightarrow T_2$	$\triangleq \Pi x : T_1. T_2$	$x \notin \text{FV}(T_2)$ Syntactic Sugar Function type
$(x : T_1) \rightarrow T_2$	$\triangleq \Pi x : T_1. T_2$	$x \in \text{FV}(T_2)$ Dependent function type
$\mathbf{data} \overline{Ru : \kappa} = K\{\overline{N : T}\}$	$\triangleq \mathbf{data} \overline{Ru : \kappa} = K\overline{T};$ $\mathbf{let} N_i : (\overline{u : \kappa}) \rightarrow \overline{Ru} \rightarrow T_i =$ $\lambda \overline{u : \kappa}. \lambda y : \overline{Ru}. \mathbf{case} y \text{ of } K \overline{x : T} \Rightarrow x_i \text{ in}$	$\forall i$ Record

**Figure 10.** Syntax of source language

### C.3 Translation to the core

See Figure 12.



$\boxed{\vdash_{\bar{s}} \Sigma}$  Well-formed context

$$\frac{}{\vdash_{\bar{s}} \emptyset} \text{ENVS\_EMPTY}$$

$$\frac{\vdash_{\bar{s}} \Sigma \quad \Sigma \vdash_{\bar{s}} T : \star}{\vdash_{\bar{s}} \Sigma, x : T} \text{ENVS\_VAR}$$

$\boxed{\Sigma \vdash_{\bar{s}} \text{pgm} : T}$  Program context

$$\frac{\overline{\Sigma_0 \vdash_{\bar{s}} \text{decl} : \Sigma'} \quad \Sigma = \Sigma_0, \overline{\Sigma'} \quad \Sigma \vdash_{\bar{s}} E : T}{\Sigma_0 \vdash_{\bar{s}} (\overline{\text{decl}}; E) : T} \text{TS_PGM\_PGM}$$

$\boxed{\Sigma \vdash_{\bar{s}} \text{decl} : \Sigma'}$  Datatype declaration

$$\frac{\Sigma \vdash_{\bar{s}} (\overline{u : \kappa}) \rightarrow \star : \star \quad \overline{\Sigma, D : (\overline{u : \kappa}) \rightarrow \star, \overline{u : \kappa} \vdash_{\bar{s}} \overline{T} \rightarrow D\overline{u} : \star}}{\Sigma \vdash_{\bar{s}} (\text{data } D\overline{u} : \kappa = \overline{K\overline{T}}) : (D : (\overline{u : \kappa}) \rightarrow \star, K : (\overline{u : \kappa}) \rightarrow \overline{T} \rightarrow D\overline{u})} \text{TSDECL\_DATA}$$

$\boxed{\Sigma \vdash_{\bar{s}} p \Rightarrow E : S \rightarrow T}$  Pattern typing

$$\frac{K : (\overline{u : \kappa}) \rightarrow \overline{S} \rightarrow D\overline{u} \in \Sigma \quad \Sigma, \overline{x : S[\overline{u \mapsto v}]} \vdash_{\bar{s}} E : T \quad \Sigma \vdash_{\bar{s}} S[\overline{u \mapsto v}] : \star}{\Sigma \vdash_{\bar{s}} Kx : S[\overline{u \mapsto v}] \Rightarrow E : D\overline{v} \rightarrow T} \text{TPAT\_ALT}$$

$\boxed{\Sigma \vdash_{\bar{s}} E : T}$  Expression typing

$$\frac{}{\emptyset \vdash_{\bar{s}} \star : \star} \text{TS\_AX}$$

$$\frac{\vdash_{\bar{s}} \Sigma \quad x : T \in \Sigma}{\Sigma \vdash_{\bar{s}} x : T} \text{TS\_VAR}$$

$$\frac{\Sigma \vdash_{\bar{s}} E_1 : (\Pi x : T_2. T_1) \quad \Sigma \vdash_{\bar{s}} E_2 : T_2}{\Sigma \vdash_{\bar{s}} E_1 E_2 : T_1[x \mapsto E_2]} \text{TS\_APP}$$

$$\frac{\Sigma, x : T_1 \vdash_{\bar{s}} E : T_2 \quad \Sigma \vdash_{\bar{s}} (\Pi x : T_1. T_2) : \star}{\Sigma \vdash_{\bar{s}} (\lambda x : T_1. E) : (\Pi x : T_1. T_2)} \text{TS\_LAM}$$

$$\frac{\Sigma \vdash_{\bar{s}} T_1 : \star \quad \Sigma, x : T_1 \vdash_{\bar{s}} T_2 : \star}{\Sigma \vdash_{\bar{s}} (\Pi x : T_1. T_2) : \star} \text{TS\_PI}$$

$$\frac{\Sigma, x : T \vdash_{\bar{s}} E : T \quad \Sigma \vdash_{\bar{s}} T : \star}{\Sigma \vdash_{\bar{s}} (\mu x : T. E) : T} \text{TS\_MU}$$

$$\frac{\Sigma \vdash_{\bar{s}} E_1 : S \quad \overline{\Sigma \vdash_{\bar{s}} p \Rightarrow E_2 : S \rightarrow T} \quad \overline{\Sigma \vdash_{\bar{s}} S \rightarrow T : \star}}{\Sigma \vdash_{\bar{s}} \text{case } E_1 \text{ of } \overline{p \Rightarrow E_2} : T} \text{TS\_CASE}$$

$$\frac{\Sigma \vdash_{\bar{s}} E_2 : T \quad \Sigma \vdash_{\bar{s}} E_1[x \mapsto E_2] : S}{\Sigma \vdash_{\bar{s}} \text{let } x : T = E_2 \text{ in } E_1 : S} \text{TS\_LET}$$

**Figure 11.** Typing rules of source language

$\boxed{\Sigma \vdash_{\text{tr}} \text{pgm} : T \rightsquigarrow e}$	Program translation
$\frac{\overline{\Sigma_0 \vdash_{\text{tr}} \text{decl} : \Sigma' \rightsquigarrow e_1} \quad \Sigma = \Sigma_0, \overline{\Sigma'} \quad \Sigma \vdash_{\text{tr}} E : T \rightsquigarrow e}{\Sigma_0 \vdash_{\text{tr}} (\overline{\text{decl}}; E) : T \rightsquigarrow \overline{e_1} \uplus e} \quad \text{TRPGM\_PGM}$	
$\boxed{\Sigma \vdash_{\text{tr}} \text{decl} : \Sigma' \rightsquigarrow e}$	Datatype translation
$\frac{\Sigma \vdash_{\text{tr}} (\overline{u} : \overline{\kappa}) \rightarrow \star : \star \rightsquigarrow (\overline{u} : \overline{\sigma}) \rightarrow \star \quad \overline{\Sigma, D : (\overline{u} : \overline{\kappa}) \rightarrow \star, \overline{u} : \overline{\kappa} \vdash_{\text{tr}} \overline{T} \rightarrow D\overline{u} : \star \rightsquigarrow \overline{\tau} \rightarrow D\overline{u}}}{\Sigma \vdash_{\text{tr}} (\text{data } D\overline{u} : \overline{\kappa} = \overline{KT}) : (D : (\overline{u} : \overline{\kappa}) \rightarrow \star, K : (\overline{u} : \overline{\kappa}) \rightarrow \overline{T} \rightarrow D\overline{u}) \rightsquigarrow e} \quad \text{TRDECL\_DATA}$ $e \triangleq \text{let } D : (\overline{u} : \overline{\sigma}) \rightarrow \star = \mu X : (\overline{u} : \overline{\sigma}) \rightarrow \star. \lambda \overline{u} : \overline{\sigma}. (\alpha : \star) \rightarrow (\overline{\tau}[D \mapsto X] \rightarrow \alpha) \rightarrow \alpha \text{ in}$ $\text{let } K_i : (\overline{u} : \overline{\sigma}) \rightarrow \overline{\tau} \rightarrow D\overline{u} = \lambda \overline{u} : \overline{\sigma}. \lambda \overline{x} : \overline{\tau}. \text{cast}_{\uparrow}^n [D\overline{u}] (\lambda \alpha : \star. \lambda \overline{b} : \overline{\tau} \rightarrow \alpha. b_i \overline{x}) \text{ in}$	
$\boxed{\Sigma \vdash_{\text{tr}} p \Rightarrow E : S \rightarrow T \rightsquigarrow e}$	Pattern translation
$\frac{K : (\overline{u} : \overline{\kappa}) \rightarrow \overline{S} \rightarrow D\overline{u} \in \Sigma \quad \Sigma, x : S[\overline{u} \mapsto \overline{v}] \vdash_{\text{tr}} E : T \rightsquigarrow e \quad \Sigma \vdash_{\text{tr}} S[\overline{u} \mapsto \overline{v}] : \star \rightsquigarrow \sigma}{\Sigma \vdash_{\text{tr}} Kx : S[\overline{u} \mapsto \overline{v}] \Rightarrow E : D\overline{v} \rightarrow T \rightsquigarrow \lambda \overline{x} : \overline{\sigma}. e} \quad \text{TRPAT\_ALT}$	
$\boxed{\Sigma \vdash_{\text{tr}} E : T \rightsquigarrow e}$	Expression translation
$\frac{}{\emptyset \vdash_{\text{tr}} \star : \star \rightsquigarrow \star} \quad \text{TR\_AX}$ $\frac{\vdash_{\text{S}} \Sigma \quad x : T \in \Sigma}{\Sigma \vdash_{\text{tr}} x : T \rightsquigarrow x} \quad \text{TR\_VAR}$ $\frac{\Sigma \vdash_{\text{tr}} E_1 : (\Pi x : T_2. T_1) \rightsquigarrow e_1 \quad \Sigma \vdash_{\text{tr}} E_2 : T_2 \rightsquigarrow e_2}{\Sigma \vdash_{\text{tr}} E_1 E_2 : T_1[x \mapsto E_2] \rightsquigarrow e_1 e_2} \quad \text{TR\_APP}$ $\frac{\Sigma, x : T_1 \vdash_{\text{tr}} E : T_2 \rightsquigarrow e \quad \Sigma \vdash_{\text{tr}} (\Pi x : T_1. T_2) : \star \rightsquigarrow \Pi x : \tau_1. \tau_2}{\Sigma \vdash_{\text{tr}} (\lambda x : T_1. E) : (\Pi x : T_1. T_2) \rightsquigarrow \lambda x : \tau_1. e} \quad \text{TR\_LAM}$ $\frac{\Sigma \vdash_{\text{tr}} T_1 : \star \rightsquigarrow \tau_1 \quad \Sigma, x : T_1 \vdash_{\text{tr}} T_2 : \star \rightsquigarrow \tau_2}{\Sigma \vdash_{\text{tr}} (\Pi x : T_1. T_2) : \star \rightsquigarrow \Pi x : \tau_1. \tau_2} \quad \text{TR\_PI}$ $\frac{\Sigma, x : T \vdash_{\text{tr}} E : T \rightsquigarrow e \quad \Sigma \vdash_{\text{tr}} T : \star \rightsquigarrow \tau}{\Sigma \vdash_{\text{tr}} (\mu x : T. E) : T \rightsquigarrow \mu x : \tau. e} \quad \text{TR\_MU}$ $\frac{\Sigma \vdash_{\text{tr}} E_1 : S \rightsquigarrow e_1 \quad \overline{\Sigma \vdash_{\text{tr}} p \Rightarrow E_2 : S \rightarrow T \rightsquigarrow e_2} \quad \overline{\Sigma \vdash_{\text{tr}} S \rightarrow T : \star \rightsquigarrow \sigma \rightarrow \tau}}{\Sigma \vdash_{\text{tr}} \text{case } E_1 \text{ of } p \Rightarrow E_2 : T \rightsquigarrow (\text{cast}_{\downarrow}^n e_1) \tau e_2} \quad \text{TR\_CASE}$ $\frac{\Sigma \vdash_{\text{tr}} E_2 : T \rightsquigarrow e_2 \quad \Sigma \vdash_{\text{tr}} E_1[x \mapsto E_2] : S \rightsquigarrow e_1[x \mapsto e_2]}{\Sigma \vdash_{\text{tr}} \text{let } x : T = E_2 \text{ in } E_1 : S \rightsquigarrow e_1[x \mapsto e_2]} \quad \text{TR\_LET}$	

**Figure 12.** Translation rules of source language