

The Combination of Dynamic and Static Typing from a Categorical Perspective

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Gradual typing was first proposed by Siek and Taha in 2006 as a way for a programming language to combine the strengths of both static and dynamic typing. However one question we must ask is, what is gradual typing? This paper contributes to answering this question by providing the first categorical model of gradual typing using the seminal work of Scott and Lambek on the categorical models of the untyped and typed λ -calculus. We then extract a functional programming language, called Grady, from the categorical model using the Curry-Howard-Lambek correspondence that combines both static and dynamic typing, but Grady is an annotated language and not a gradual type system. Finally, we show that Siek and Taha's gradual type system can be translated into Grady, and that their original annotated language is equivalent in expressive power to Grady.

CCS Concepts: •**Theory of computation** → **Denotational semantics**; **Categorical semantics**; *Type theory*; *Functional constructs*; *Type structures*;

Additional Key Words and Phrases: static typing, dynamic typing, gradual typing, categorical semantics, retract, typed lambda-calculus, untyped lambda-calculus, gradual typing, static typing, dynamic typing, categorical model, functional programming

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1 INTRODUCTION

(Scott 1980) showed how to model the untyped λ -calculus within a cartesian closed category, C , with a distinguished object we will call $?$ – read as the type of untyped terms – such that the object¹ $? \rightarrow ?$ is a retract of $?$. That is, there are morphisms $\text{squash} : (? \rightarrow ?) \longrightarrow ?$ and $\text{split} : ? \longrightarrow (? \rightarrow ?)$ where $\text{squash}; \text{split} = \text{id} : (? \rightarrow ?) \longrightarrow (? \rightarrow ?)$ ². For example, taking these morphisms as terms in the typed λ -calculus we can define the prototypical looping term $(\lambda x.x x)(\lambda x.x x)$ by $(\lambda x : ?. (\text{split } x) x)$ ($\text{squash } (\lambda x : ?. (\text{split } x) x)$).

In the same volume as Scott (Lambek 1980) showed that cartesian closed categories also model the typed λ -calculus. Suppose we want to model the typed λ -calculus with pairs and natural numbers. That is, given two types A_1 and A_2 there is a type $A_1 \times A_2$, and there is a type Nat . Furthermore, we have first and second projections, and zero and successor functions. This situation can easily be modeled by a cartesian closed category C – see Section ?? for the details – but also add to C the type of untyped terms $?$, squash , and split . At this point C is a model of both the typed and the untyped λ -calculus. However, the two theories are really just sitting side by side in C and cannot really interact much.

Suppose \mathcal{T} is a discrete category with the objects Nat and Unit (the terminal object or empty product) and $T : \mathcal{T} \longrightarrow C$ is a full and faithful functor. This implies that \mathcal{T} is a subcategory of C , and that \mathcal{T} is the category of atomic types. Then

¹We will use the terms “object” and “type” interchangeably.

²We denote composition of morphisms by $f; g : A \longrightarrow C$ given morphisms $f : A \longrightarrow B$ and $g : B \longrightarrow C$.

for any type A of \mathcal{T} we add to \mathcal{C} the morphisms $\text{box} : TA \rightarrow ?$ and $\text{unbox} : ? \rightarrow TA$ such that $\text{box}; \text{unbox} = \text{id} : TA \rightarrow TA$ making TA a retract of $?$. This is the bridge allowing the typed world to interact with the untyped one. We can think of box as injecting typed data into the untyped world, and unbox as taking it back. Notice that the only time we can actually get the typed data back out is if it were injected into the untyped world initially. In the model this is enforced through composition, but in the language this will be enforced at runtime, and hence, requires the language to contain dynamic typing. Thus, what we have just built up is a categorical model that offers a new perspective of how to combine static and dynamic typing.

(Siek and Taha 2006) define gradual typing to be the combination of both static and dynamic typing that allows for the programmer to program in dynamic style, and thus, annotations should be suppressed. This means that a gradually typed program can utilize both static types which will be enforced during compile time, but may also utilize dynamic typing that will be enforced during runtime. Therefore, gradual typing is the best of both worlds.

Siek and Taha's gradually typed functional language is the typed λ -calculus with the type of untyped terms $?$ and the following rules:

$$\frac{\Gamma \vdash t_1 : ? \quad \Gamma \vdash t_2 : A}{\Gamma \vdash t_1 t_2 : ?} \qquad \frac{\Gamma \vdash t_1 : A_1 \rightarrow B \quad \Gamma \vdash t_2 : A_2 \quad A_1 \sim A_2}{\Gamma \vdash t_1 t_2 : B}$$

The premise $A \sim B$ is read, the type A is consistent with the type B , and is defined in Figure 2. If we squint we can see split, squash, box, and unbox hiding in the definition of the previous rules, but they have been suppressed. We will show that when one uses either of the two typing rules then one is really implicitly using a casting morphism built from split, squash, box, and unbox. In fact, the consistency relation $A \sim B$ can be interpreted as such a morphism. Then the typing above can be read semantically as a saying if a casting morphism exists, then the type really can be converted into the necessary type.

Contributions. This paper offers the following contributions:

- A new categorical model for gradual typing for functional languages. We show how to interpret (Siek and Taha 2006)'s gradual type system in the categorical model outlined above. As far as the authors are aware this is the first categorical model for gradual typing.
- We then extract a functional programming language called Grady from the categorical model via the Curry-Howard-Lambek correspondence. This is not a gradual type system, but can be seen as an alternative annotated language in which Siek and Taha's gradual type system can be translated to.
- A proof that Grady is as expressive as (Siek and Taha 2006)'s annotated language and vice versa. We give a type directed translation of Siek and Taha's annotated language to Grady and vice versa, then we show that these translations preserve evaluation.
- Having the untyped λ -calculus along side the typed λ -calculus can be a lot of fun. We show how to Church encode typed data, utilize the Y-combinator, and even obtain terminating recursion on natural numbers by combining the Y-combinator with a natural number eliminator. Thus, obtaining the expressive power of Gödel's system T (Girard et al. 1989).

Related work. TODO

$\frac{x : A \in \Gamma}{\Gamma \vdash x : A} \text{var}$	$\frac{}{\Gamma \vdash \text{triv} : \text{Unit}} \text{unit}$	$\frac{}{\Gamma \vdash 0 : \text{Nat}} \text{zero}$	$\frac{\Gamma \vdash t : \text{Nat}}{\Gamma \vdash \text{succ } t : \text{Nat}} \text{succ}$	$\frac{\Gamma \vdash t_1 : A_1 \quad \Gamma \vdash t_2 : A_2}{\Gamma \vdash (t_1, t_2) : A_1 \times A_2} \times$
$\frac{\Gamma \vdash t : A_1 \times B \quad A_1 \sim A_2}{\Gamma \vdash \text{fst } t : A_2} \times_{e_1}$	$\frac{\Gamma \vdash t : A \times B_1 \quad B_1 \sim B_2}{\Gamma \vdash \text{snd } t : B_2} \times_{e_2}$	$\frac{\Gamma, x : A \vdash t : B}{\Gamma \vdash \lambda x : A_1. t : A \rightarrow B} \rightarrow$		
$\frac{\Gamma \vdash t_1 : A_1 \rightarrow B \quad \Gamma \vdash t_2 : A_2 \quad A_1 \sim A_2}{\Gamma \vdash t_1 t_2 : B} \rightarrow_e$	$\frac{\Gamma \vdash t : ?}{\Gamma \vdash \text{succ } t : ?} \text{succ}^?$	$\frac{\Gamma \vdash t : ?}{\Gamma \vdash \text{fst } t : ?} \times_{e_1}^?$	$\frac{\Gamma \vdash t : ?}{\Gamma \vdash \text{snd } t : ?} \times_{e_2}^?$	
	$\frac{\Gamma \vdash t_1 : ? \quad \Gamma \vdash t_2 : A}{\Gamma \vdash t_1 t_2 : ?} \rightarrow_e^?$			

Fig. 1. Typing rules for $\lambda_{\rightarrow}^?$

$\frac{}{A \sim A} \text{refl}$	$\frac{}{A \sim ?} \text{box}$	$\frac{}{? \sim A} \text{unbox}$	$\frac{A_1 \sim A_2 \quad B_1 \sim B_2}{A_1 \rightarrow B_1 \sim A_2 \rightarrow B_2} \rightarrow$	$\frac{A_1 \sim A_2 \quad B_1 \sim B_2}{A_1 \times B_1 \sim A_2 \times B_2} \times$
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Fig. 2. Type Consistency for $\lambda_{\rightarrow}^?$

2 GRADUAL TYPING

We begin by introducing a slight variation of (Siek and Taha 2006)'s gradually typed functional language. It has been extended with product types and natural numbers, and instead of a big-step call-by-value operational semantics it uses a single-step type directed full $\beta\eta$ -evaluator. One thing we strive for in this paper is to keep everything as simple as possible so that the underlying structure of these languages shines through. In this vein, the change in evaluation makes it easier to interpret the language into the categorical model.

The syntax of the gradual type system $\lambda_{\rightarrow}^?$ is defined in the following definition.

Definition 2.1. Syntax for $\lambda_{\rightarrow}^?$:

(types)	$A, B ::= \text{Unit} \mid \text{Nat} \mid ? \mid A \times B \mid A_1 \rightarrow A_2$
(terms)	$t ::= x \mid \text{triv} \mid 0 \mid \text{succ } t \mid \lambda x : A. t \mid t_1 t_2 \mid (t_1, t_2) \mid \text{fst } t \mid \text{snd } t$
(contexts)	$\Gamma ::= \cdot \mid x : A \mid \Gamma_1, \Gamma_2$

This definition is the base syntax for every language in this paper. The typing rules are defined in Figure 1 and the type consistency relation is defined in Figure 2. The main changes of the version of $\lambda_{\rightarrow}^?$, defined here from the original due to (Siek and Taha 2006) is that products and natural numbers have been added. The definition of products follows how casting is done for functions. So it allows casting projections of products, for example, it is reasonable for terms like $\lambda x : (? \times ?).(\text{succ } (\text{fst } x))$ to type check.

We can view gradual typing as a surface language feature much like type inference, and we give it a semantics by translating it into an annotated core. (Siek and Taha 2006) do just that and give $\lambda_{\rightarrow}^?$ an operational semantics by translating it to a fully annotated core language called $\lambda_{\rightarrow}^{\langle A \rangle}$. Its syntax is an extension of the syntax of $\lambda_{\rightarrow}^?$ (Definition 2.1) where terms are the only syntactic class that differs, and so we do not repeat the syntax of types or contexts.

$$\begin{array}{c}
\frac{x : A \in \Gamma}{\Gamma \vdash x : A} \text{var} \quad \frac{}{\Gamma \vdash \text{triv} : \text{Unit}} \text{unit} \quad \frac{}{\Gamma \vdash 0 : \text{Nat}} \text{zero} \quad \frac{\Gamma \vdash t : \text{Nat}}{\Gamma \vdash \text{succ } t : \text{Nat}} \text{succ} \quad \frac{\Gamma \vdash t_1 : A_1 \quad \Gamma \vdash t_2 : A_2}{\Gamma \vdash (t_1, t_2) : A_1 \times A_2} \times \\
\frac{\Gamma \vdash t : A_1 \times A_2}{\Gamma \vdash \text{fst } t : A_1} \times_{e_1} \quad \frac{\Gamma \vdash t : A_1 \times A_2}{\Gamma \vdash \text{snd } t : A_2} \times_{e_2} \quad \frac{\Gamma, x : A \vdash t : B}{\Gamma \vdash \lambda x : A_1. t : A \rightarrow B} \rightarrow \quad \frac{\Gamma \vdash t_1 : A \rightarrow B \quad \Gamma \vdash t_2 : A}{\Gamma \vdash t_1 t_2 : B} \rightarrow_e \\
\frac{\Gamma \vdash t : A \quad A \sim B}{\Gamma \vdash \langle B \rangle t : B} \text{cast}
\end{array}$$
Fig. 3. Typing rules for $\lambda_{\rightarrow}^{(A)}$

$$\begin{array}{c}
\frac{\Gamma \vdash v : A}{\Gamma \vdash v \rightsquigarrow v : A} \text{value} \quad \frac{\Gamma \vdash \text{drop-cast } v : C}{\Gamma \vdash \langle C \rangle v \rightsquigarrow \text{drop-cast } v : C} \text{value-cast} \quad \frac{\Gamma \vdash t : ?}{\Gamma \vdash \langle \text{Nat} \rangle (\text{succ } t) \rightsquigarrow \text{succ } \langle \text{Nat} \rangle t : \text{Nat}} \text{Nat-cast} \\
\frac{\Gamma \vdash t : A_1 \rightarrow B_1 \quad (A_1 \rightarrow B_1) \sim (A_2 \rightarrow B_2)}{\Gamma \vdash \langle A_2 \rightarrow B_2 \rangle t \rightsquigarrow \lambda y : A_2. \langle B_2 \rangle (t \langle A_1 \rangle y) : A_2 \rightarrow B_2} \rightarrow\text{-cast} \quad \frac{\Gamma \vdash t : A_1 \times B_1 \quad (A_1 \times B_1) \sim (A_2 \times B_2)}{\Gamma \vdash \langle A_2 \times B_2 \rangle t \rightsquigarrow \langle \langle A_2 \rangle (\text{fst } t), \langle B_2 \rangle (\text{snd } t) \rangle : A_2 \times B_2} \times\text{-cast} \\
\frac{\Gamma \vdash t_1 \rightsquigarrow t_2 : A \quad A \sim B}{\Gamma \vdash \langle B \rangle t_1 \rightsquigarrow \langle B \rangle t_2 : B} \text{cast} \quad \frac{\Gamma, x : A_1 \vdash t_2 : A_2 \quad \Gamma \vdash t_1 : A_1}{\Gamma \vdash (\lambda x : A_1. t_2) t_1 \rightsquigarrow [t_1/x] t_2 : A_2} \beta \quad \frac{\Gamma \vdash t : A_1 \rightarrow A_2 \quad x \notin \text{FV}(t)}{\Gamma \vdash \lambda x : A_1. t x \rightsquigarrow t : A_1 \rightarrow A_2} \eta \\
\frac{\Gamma, x : A_1 \vdash t \rightsquigarrow t' : A_2}{\Gamma \vdash \lambda x : A_1. t \rightsquigarrow \lambda x : A_1. t' : A_1 \rightarrow A_2} \rightarrow \quad \frac{\Gamma \vdash t_1 \rightsquigarrow t'_1 : A_1 \rightarrow A_2 \quad \Gamma \vdash t_2 : A_1}{\Gamma \vdash t_1 t_2 \rightsquigarrow t'_1 t_2 : A_2} \rightarrow_{e_1} \\
\frac{\Gamma \vdash t_1 : A_1 \rightarrow A_2 \quad \Gamma \vdash t_2 \rightsquigarrow t'_2 : A_1}{\Gamma \vdash t_1 t_2 \rightsquigarrow t_1 t'_2 : A_2} \rightarrow_{e_2} \quad \frac{\Gamma \vdash t \rightsquigarrow t' : A_1 \times A_2}{\Gamma \vdash \text{fst } t \rightsquigarrow \text{fst } t' : A_1} \times_{e_1} \quad \frac{\Gamma \vdash t \rightsquigarrow t' : A_1 \times A_2}{\Gamma \vdash \text{snd } t \rightsquigarrow \text{snd } t' : A_2} \times_{e_2} \\
\frac{\Gamma \vdash t : A_1 \times A_2}{\Gamma \vdash (\text{fst } t, \text{snd } t) \rightsquigarrow t : A_1 \times A_2} \times_{\eta} \quad \frac{\Gamma \vdash t_1 \rightsquigarrow t'_1 : A_1 \quad \Gamma \vdash t_2 : A_2}{\Gamma \vdash (t_1, t_2) \rightsquigarrow (t'_1, t_2) : A_1 \times A_2} \times_1 \quad \frac{\Gamma \vdash t_1 : A_1 \quad \Gamma \vdash t_2 \rightsquigarrow t'_2 : A_2}{\Gamma \vdash (t_1, t_2) \rightsquigarrow (t_1, t'_2) : A_1 \times A_2} \times_2
\end{array}$$
Fig. 4. Reduction rules for $\lambda_{\rightarrow}^{(A)}$

Definition 2.2. Syntax for $\lambda_{\rightarrow}^{(A)}$:

(simple values) $s ::= x \mid \text{triv} \mid 0$
 (values) $v ::= s \mid \langle ? \rangle s$
 (terms) $t ::= \dots \mid \langle A \rangle t$

The typing rules for $\lambda_{\rightarrow}^{(A)}$ can be found in Figure 3, and the reduction rules in Figure 4.

The major difference from the formalization of $\lambda_{\rightarrow}^{(A)}$ given here and Siek and Taha's is that it is single step and full β -reduction, but it is based on their original definition. The function $\text{drop-cast } v$ is defined as follows:

$\text{drop-cast } \langle ? \rangle s = s$
 $\text{drop-cast } s = s$

This function is used when casting values to their appropriate type.

Since the formalization of both $\lambda_{\rightarrow}^?$ and $\lambda_{\rightarrow}^{(A)}$ differ from their original definitions we give the definition of cast insertion in Figure 5, but this is only a slightly modified version from the one given by Siek and Taha.

$$\begin{array}{c}
\frac{x : A \in \Gamma}{\Gamma \vdash x \Rightarrow x : A} \text{var} \quad \frac{}{\Gamma \vdash 0 \Rightarrow 0 : A} \text{zero} \quad \frac{}{\Gamma \vdash \text{triv} \Rightarrow \text{triv} : \text{Unit}} \text{unit} \quad \frac{\Gamma \vdash t_1 \Rightarrow t_2 : \text{Nat}}{\Gamma \vdash \text{succ } t_1 \Rightarrow \text{succ } t_2 : \text{Nat}} \text{succ} \\
\\
\frac{\Gamma \vdash t_1 \Rightarrow t_2 : ?}{\Gamma \vdash \text{succ } t_1 \Rightarrow \langle ? \rangle \text{succ } \langle \text{Nat} \rangle t_2 : ?} \text{succ}^? \quad \frac{\Gamma \vdash t_1 \Rightarrow t_3 : A_1 \quad \Gamma \vdash t_2 \Rightarrow t_4 : A_2}{\Gamma \vdash (t_1, t_2) \Rightarrow (t_3, t_4) : A_1 \times A_2} \times \\
\\
\frac{\Gamma \vdash t_1 \Rightarrow t_2 : A_1 \times B \quad A_1 \sim A_2 \quad A_1 \neq A_2}{\Gamma \vdash \text{fst } t_1 \Rightarrow \text{fst } \langle A_2 \times B \rangle t_2 : A_2} \times_{e_1}^{\sim} \quad \frac{\Gamma \vdash t_1 \Rightarrow t_2 : A \times B}{\Gamma \vdash \text{fst } t_1 \Rightarrow \text{fst } t_2 : A} \times_{e_1} \\
\\
\frac{\Gamma \vdash t_1 \Rightarrow t_2 : A \times B_1 \quad B_1 \sim B_2 \quad B_1 \neq B_2}{\Gamma \vdash \text{snd } t_1 \Rightarrow \text{snd } \langle A \times B_2 \rangle t_2 : B_2} \times_{e_2}^{\sim} \quad \frac{\Gamma \vdash t_1 \Rightarrow t_2 : A \times B}{\Gamma \vdash \text{snd } t_1 \Rightarrow \text{snd } t_2 : B} \times_{e_2} \\
\\
\frac{\Gamma \vdash t_1 \Rightarrow t_2 : ?}{\Gamma \vdash \text{fst } t_1 \Rightarrow \text{fst } \langle ? \times ? \rangle t_2 : ?} \times_{e_1}^? \quad \frac{\Gamma \vdash t_1 \Rightarrow t_2 : ?}{\Gamma \vdash \text{snd } t_1 \Rightarrow \text{snd } \langle ? \times ? \rangle t_2 : ?} \times_{e_2}^? \\
\\
\frac{\Gamma, x : A_1 \vdash t_1 \Rightarrow t_2 : A_2}{\Gamma \vdash \lambda x : A_1. t_1 \Rightarrow \lambda x : A_1. t_2 : A_1 \rightarrow A_2} \rightarrow \quad \frac{\Gamma \vdash t_1 \Rightarrow t_3 : ? \quad \Gamma \vdash t_2 \Rightarrow t_4 : A}{\Gamma \vdash t_1 t_2 \Rightarrow \langle \langle A \rightarrow ? \rangle t_3 \rangle t_4 : ?} \rightarrow_e^? \\
\\
\frac{\Gamma \vdash t_1 \Rightarrow t_3 : A_1 \rightarrow B \quad \Gamma \vdash t_2 \Rightarrow t_4 : A_2 \quad A_1 \sim A_2 \quad A_1 \neq A_2}{\Gamma \vdash t_1 t_2 \Rightarrow t_1 \langle A_1 \rangle t : B} \rightarrow_e^{\sim} \quad \frac{\Gamma \vdash t_1 \Rightarrow t_3 : A_1 \rightarrow A_2 \quad \Gamma \vdash t_2 \Rightarrow t_4 : A_1}{\Gamma \vdash t_1 t_2 \Rightarrow t_3 t_4 : A_2} \rightarrow_e
\end{array}$$

Fig. 5. Cast Insertion

3 THE CATEGORICAL PERSPECTIVE

The strength and main motivation for giving a categorical model to a programming language is that it can expose the fundamental structure of the language. This arises because a lot of the language features that often cloud the picture go away, for example, syntactic notions like variables disappear. This can often simplify things and expose the underlying structure. Reynolds (?) was a big advocate for the use of category theory in programming language research for these reasons. For example, when giving the simply typed λ -calculus a categorical model we see that it is a cartesian closed category, but we also know that intuitionistic logic has the same model due to (Lambek 1980); on the syntactic side these two theories are equivalent as well due to (Howard 1980). Thus, the fundamental structure of the simply typed λ -calculus is intuitionistic logic. This also shows a relationship between seemingly unrelated theories. It is quite surprising that these two theories are related. Another more recent example of this can be found in the connection between dependent type theory and homotopy theory (?).

Another major benefit of studying the categorical model of programming languages is that it gives us a powerful tool to study language extensions. For example, purely functional programming in Haskell would not be where it is without the seminal work of Moggi and Wadler (?) on using monads – a purely categorical notion – to add side effects to Haskell. Thus, we believe that developing these types of models for new language designs and features can be hugely beneficial.

Interpreting a programming language into a categorical model requires three steps. First, the types are interpreted as objects. Then programs are interpreted as morphisms in the category, but this is a simplification. Every morphism, f , in a category has a source object and a target object, we usually denote this by $f : A \multimap B$. Thus, in order to interpret

programs as morphisms the program must have a source and target. So instead of interpreting raw terms as morphisms we interpret terms in their typing context. That is, we must show how to interpret every $\Gamma \vdash t : A$ as a morphism $t : \llbracket \Gamma \rrbracket \longrightarrow \llbracket A \rrbracket$. The third step is to show that whenever one program reduces to another their interpretations are isomorphic in the model. This means that whenever $\Gamma \vdash t_1 \rightsquigarrow t_2 : A$, then $\llbracket t_1 \rrbracket = \llbracket t_2 \rrbracket : \llbracket \Gamma \rrbracket \longrightarrow \llbracket A \rrbracket$. This is the reason why we defined our reduction in a typed fashion to aid us in understanding how it relates to the model. For a more thorough introduction see (Crole 1994).

3.1 The Categorical Model

We now give a categorical model for $\lambda_{\rightarrow}^?$ and $\lambda_{\rightarrow}^{(A)}$. The model we develop here builds on the seminal work of (Lambek 1980) and (Scott 1980). (Lambek 1980) showed that the typed λ -calculus can be modeled by a cartesian closed category. In the same volume as Lambek, Scott essentially showed that the untyped λ -calculus is actually typed. That is, typed theories are more fundamental than untyped ones. He accomplished this by adding a single type, $?$, and two functions $\text{squash} : (? \rightarrow ?) \rightarrow ?$ and $\text{split} : ? \rightarrow (? \rightarrow ?)$, such that, $\text{squash}; \text{split} = \text{id} : (? \rightarrow ?) \rightarrow (? \rightarrow ?)$. At this point he was able to translate the untyped λ -calculus into this untyped one. Categorically, he modeled split and squash as the morphisms in a retract within a cartesian closed category – the same model as typed λ -calculus.

Definition 3.1. Suppose C is a category. Then an object A is a **retract** of an object B if there are morphisms $i : A \longrightarrow B$ and $r : B \longrightarrow A$ such that the following diagram commutes:

$$\begin{array}{ccc} A & \xrightarrow{i} & B \\ & \searrow & \downarrow r \\ & & A \end{array}$$

Thus, $? \rightarrow ?$ is a retract of $?$, but we extend this slightly to include $? \times ?$ being a retract of $?$. This is only a slight extension of Scott's model, because our languages will have products where he did not consider products, because he was considering the traditional definition of the untyped λ -calculus.

We can now define our categorical model of the untyped λ -calculus with products.

Definition 3.2. An **untyped λ -model**, $(C, ?, \text{split}, \text{squash})$, is a cartesian closed category C with a distinguished object $?$ and morphisms $\text{squash} : S \longrightarrow ?$ and $\text{split} : ? \longrightarrow S$ making the object S a retract of $?$, where S is either $? \rightarrow ?$ or $? \times ?$.

THEOREM 3.3 (SCOTT (1980)). *An untyped λ -model is a sound and complete model of the untyped λ -calculus.*

Since all of the languages we are studying here contain the natural numbers we must be able to interpret them into our model. We give a novel approach to modeling the natural numbers with their (non-recursive) eliminator using what we call a Scott natural number object. Now the natural number eliminator is not part of $\lambda_{\rightarrow}^?$ or $\lambda_{\rightarrow}^{(A)}$, but we want Grady to contain it, and Grady will directly correspond to the model.

Definition 3.4. Suppose C is a cartesian closed category. A **Scott natural number object (SNNO)** is an object Nat of C and morphisms $z : 1 \longrightarrow \text{Nat}$ and $\text{succ} : \text{Nat} \longrightarrow \text{Nat}$ of C , such that, for any morphisms $f : Y \longrightarrow X$ and

$g : Y \times \text{Nat} \longrightarrow X$ of \mathcal{C} there is a unique morphism $\text{case}_X : Y \times \text{Nat} \longrightarrow X$ making the following diagrams commute:

$$\begin{array}{ccccc}
 Y \times \text{Nat} & \xrightarrow{\text{id}_Y \times z} & Y \times \text{Nat} & \xleftarrow{\text{id}_Y \times \text{succ}} & Y \times \text{Nat} \\
 & \searrow \pi_1; f & \downarrow \text{case}_X & \swarrow g & \\
 & & X & &
 \end{array}$$

Informally, the two diagrams essentially assert that we can define case_X as follows:

$$\begin{aligned}
 \text{case}_X y \ 0 &= f \ y \\
 \text{case}_X y \ (\text{succ } x) &= g \ y \ x
 \end{aligned}$$

This formalization of natural numbers is inspired by the definition of Scott Numerals (?) where the notion of a case distinction is built into the encoding. We can think of Y in the source object of case as the type of additional inputs that will be passed to both f and g , but we can think of Nat in the source object of case as the type of the scrutiny. Thus, since in the base case there is no predecessor, f , will not require the scrutiny, and so it is ignored.

One major difference between SNNOs and the more traditional natural number objects is that in the definition of the latter g is defined by well-founded recursion. However, SNNOs do not allow this, but in the presence of fixpoints we are able to regain this feature without having to bake it into the definition of natural number objects. However, to allow this we have found that when combining fixpoints and case analysis to define terminating functions on the natural numbers it is necessary to uniformly construct the input to both f and g due to the reduction rule of the Y combinator. Thus, we extend the type of f to $Y \times \text{Nat}$, but then ignore the second projection when reaching the base case.

So far we can model the untyped and the typed λ -calculi within a cartesian closed category, but we do not have any way of moving typed data into the untyped part and vice versa. To accomplish this we add two new morphisms $\text{box}_C : C \longrightarrow ?$ and $\text{unbox}_C : ? \longrightarrow C$ such that $\text{box}_C; \text{unbox}_C = \text{id} : C \longrightarrow C$ for every atomic type C . Thus, each atomic type is a retract of $?$. This enforces that the only time we can really consider something as typed is if it were boxed up in the first place. We can look at this from another perspective as well. If the programmer tries to unbox something that is truly untyped, then their program may actually type check, but they will obtain a dynamic type error at runtime, because the unbox will never have been matched up with the correct boxed data. For example, we can cast 3 to type Bool by $\text{unbox}_{\text{Bool}}(\text{box}_{\text{Nat}} 3)$, but if this program is every run, then we will obtain a dynamic type error. Note that we can type the previous program in $\lambda_{\rightarrow}^{(A)}$ as well, but if we run the program it will result in a dynamic type error too.

Now we combine everything we have discussed so far to obtain the categorical model.

Definition 3.5. A **gradual λ -model**, $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$, where \mathcal{T} is a discrete category with at least two objects Nat and Unit , C is a cartesian closed category with a SNN, $(C, ?, \text{split}, \text{squash})$ is an untyped λ -model, $T : \mathcal{T} \longrightarrow C$ is an embedding – a full and faithful functor that is injective on objects – and for every object A of \mathcal{T} there are morphisms $\text{box}_A : TA \longrightarrow ?$ and $\text{unbox}_A : ? \longrightarrow TA$ making TA a retract of $?$.

We call the category \mathcal{T} the category of atomic types. We call an object, A , **atomic** iff there is some object A' in \mathcal{T} such that $A = TA'$. Note that we do not consider $?$ an atomic type. The model really is the cartesian closed category C , but it is extended with the structure of both the typed and the untyped λ -calculus with the ability to cast data.

Interpreting the typing rules for $\lambda_{\rightarrow}^?$, will require the interpretation of type consistency. Thus, we must be able to cast any type A to $?$, but as stated the model only allows atomic types to be casted. It turns out that this can be lifted to any type.

We call any morphism defined completely in terms of id , the functors $- \times -$ and $- \rightarrow -$, split and squash, and box and unbox a **casting morphism**. To cast any type A to $?$ we will build casting morphisms that first take the object A to its skeleton, and then takes the skeleton to $?$.

Definition 3.6. Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model. Then the **skeleton** of an object A of C is an object S that is constructed by replacing each atomic type in A with $?$. Given an object A we denote its skeleton by $\text{skeleton } A$.

One should think of the skeleton of an object as the supporting type structure of the object, but we do not know what kind of data is actually in the structure. For example, the skeleton of the object Nat is $?$, and the skeleton of $(\text{Nat} \times \text{Unit}) \rightarrow \text{Nat} \rightarrow \text{Nat}$ is $(? \times ?) \rightarrow ? \rightarrow ?$.

The next definition defines a means of constructing a casting morphism that casts a type A to its skeleton and vice versa. This definition is by mutual recursion on the input type.

Definition 3.7. Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model. Then for any object A whose skeleton is S we define the morphisms $\widehat{\text{box}}_A : A \rightarrow S$ and $\widehat{\text{unbox}}_A : S \rightarrow A$ by mutual recursion on A as follows:

$$\begin{array}{l|l}
 \widehat{\text{box}}_A = \text{box}_A & \widehat{\text{unbox}}_A = \text{unbox}_A \\
 \text{when } A \text{ is atomic} & \text{when } A \text{ is atomic} \\
 \\
 \widehat{\text{box}}_? = \text{id}_? & \widehat{\text{unbox}}_? = \text{id}_? \\
 \\
 \widehat{\text{box}}_{(A_1 \rightarrow A_2)} = \widehat{\text{unbox}}_{A_1} \rightarrow \widehat{\text{box}}_{A_2} & \widehat{\text{unbox}}_{(A_1 \rightarrow A_2)} = \widehat{\text{box}}_{A_1} \rightarrow \widehat{\text{unbox}}_{A_2} \\
 \\
 \widehat{\text{box}}_{(A_1 \times A_2)} = \widehat{\text{box}}_{A_1} \times \widehat{\text{box}}_{A_2} & \widehat{\text{unbox}}_{(A_1 \times A_2)} = \widehat{\text{unbox}}_{A_1} \times \widehat{\text{unbox}}_{A_2}
 \end{array}$$

The definition of both $\widehat{\text{box}}$ or $\widehat{\text{unbox}}$ uses the functor $- \rightarrow - : C^{\text{op}} \times C \rightarrow C$ which is contravariant in its first argument, and thus, in that contravariant position we must make a recursive call to the opposite function, and hence, they must be mutually defined. Every call to either $\widehat{\text{box}}$ or $\widehat{\text{unbox}}$ in the previous definition is on a smaller object than the input object. Thus, their definitions are well founded. Furthermore, $\widehat{\text{box}}$ and $\widehat{\text{unbox}}$ form a retract between A and S .

LEMMA 3.8 (BOXING AND UNBOXING LIFTED RETRACT). *Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model. Then for any object A ,*

$$\widehat{\text{box}}_A; \widehat{\text{unbox}}_A = \text{id}_A : A \rightarrow A.$$

PROOF. This proof holds by induction on the form A . Please see Appendix A.1 for the complete proof. \square

As an example, suppose we wanted to cast the type $(\text{Nat} \times ?) \rightarrow \text{Nat}$ to its skeleton $(? \times ?) \rightarrow ?$. Then we can obtain a casting morphisms that will do this as follows:

$$\begin{aligned}
 \widehat{\text{box}}_{((\text{Nat} \times ?) \rightarrow \text{Nat})} &= \widehat{\text{unbox}}_{(\text{Nat} \times ?)} \rightarrow \widehat{\text{box}}_{\text{Nat}} \\
 &= (\widehat{\text{unbox}}_{\text{Nat}} \times \widehat{\text{unbox}}_?) \rightarrow \widehat{\text{box}}_{\text{Nat}} \\
 &= (\text{unbox}_{\text{Nat}} \times \text{id}_?) \rightarrow \text{box}_{\text{Nat}}
 \end{aligned}$$

We can also cast a morphism $A \xrightarrow{f} B$ to a morphism

$$S_1 \xrightarrow{\widehat{\text{unbox}}_A} A \xrightarrow{f} B \xrightarrow{\widehat{\text{box}}_B} S_2$$

where $S_1 = \text{skeleton } A$ and $S_2 = \text{skeleton } B$. Now if we have a second

$$S_2 \xrightarrow{\widehat{\text{unbox}}_B} B \xrightarrow{g} C \xrightarrow{\widehat{\text{box}}_C} S_3$$

then their composition reduces to composition at the typed level:

$$\begin{array}{ccccccc} S_1 & \xrightarrow{\widehat{\text{unbox}}_A} & A & \xrightarrow{f} & B & \xrightarrow{\widehat{\text{box}}_B} & S_2 \\ \downarrow & & \downarrow f;g & & \parallel & & \parallel \\ S_3 & \xleftarrow{\widehat{\text{box}}_C} & C & \xleftarrow{g} & B & \xleftarrow{\widehat{\text{unbox}}_B} & S_2 \end{array}$$

The right most diagram commutes because B is a retract of S_2 , and the left unannotated arrow is the composition $\widehat{\text{unbox}}_A; f; g; \widehat{\text{box}}_C$. This tells us that we have a functor $S : C \rightarrow \mathcal{S}$:

$$\begin{aligned} SA &= \text{skeleton } A \\ S(f : A \rightarrow B) &= \widehat{\text{unbox}}_A; f; \widehat{\text{box}}_A \end{aligned}$$

where \mathcal{S} is the full subcategory of C consisting of the skeletons and morphisms between them, that is, \mathcal{S} is a cartesian closed category with one basic object $?$ such that $(\mathcal{S}, ?, \text{split}, \text{squash})$ is an untyped λ -model. The following turns out to be true.

LEMMA 3.9 (S IS FAITHFUL). *Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model, and $(\mathcal{S}, ?, \text{split}, \text{squash})$ is the category of skeletons. Then the functor $S : C \rightarrow \mathcal{S}$ is faithful.*

PROOF. This proof follows from the definition S and Lemma 3.8. For the full proof see Appendix A.2. \square

Thus, we can think of the functor S as an injection of the typed world into the untyped one.

Now that we can cast any type into its skeleton we must show that every skeleton can be cast to $?$. We do this similarly to the above and lift split and squash to arbitrary skeletons.

Definition 3.10. Suppose $(\mathcal{S}, ?, \text{split}, \text{squash})$ is the category of skeletons. Then for any skeleton S we define the morphisms $\widehat{\text{squash}}_S : S \rightarrow ?$ and $\widehat{\text{split}}_S : ? \rightarrow S$ by mutual recursion on S as follows:

$\widehat{\text{squash}}_? = \text{id}_?$	$\widehat{\text{split}}_? = \text{id}_?$
$\widehat{\text{squash}}_{(S_1 \rightarrow S_2)} = (\widehat{\text{split}}_{S_1} \rightarrow \widehat{\text{squash}}_{S_2}); \text{squash}_{? \rightarrow ?}$	$\widehat{\text{split}}_{(S_1 \rightarrow S_2)} = \text{split}_{? \rightarrow ?}; (\widehat{\text{squash}}_{S_1} \rightarrow \widehat{\text{split}}_{S_2})$
$\widehat{\text{squash}}_{(S_1 \times S_2)} = (\widehat{\text{squash}}_{S_1} \times \widehat{\text{squash}}_{S_2}); \text{squash}_{? \times ?}$	$\widehat{\text{split}}_{(S_1 \times S_2)} = \text{split}_{? \times ?}; (\widehat{\text{split}}_{S_1} \times \widehat{\text{split}}_{S_2})$

As an example we will construct the casting morphism that casts the skeleton $(? \times ?) \rightarrow ?$ to $?$:

$$\begin{aligned} \widehat{\text{squash}}_{(? \times ?) \rightarrow ?} &= (\widehat{\text{split}}_{? \times ?} \rightarrow \widehat{\text{squash}}_{?}); \widehat{\text{squash}}_{? \rightarrow ?} \\ &= (\widehat{\text{split}}_{? \times ?}; (\widehat{\text{split}}_{?} \times \widehat{\text{split}}_{?})) \rightarrow \widehat{\text{squash}}_{?}; \widehat{\text{squash}}_{? \rightarrow ?} \\ &= ((\widehat{\text{split}}_{? \times ?}; (\text{id}_{?} \times \text{id}_{?})) \rightarrow \text{id}_{?}); \widehat{\text{squash}}_{? \rightarrow ?} \\ &= (\widehat{\text{split}}_{? \times ?} \rightarrow \text{id}_{?}); \widehat{\text{squash}}_{? \rightarrow ?} \end{aligned}$$

The morphisms $\widehat{\text{split}}_S$ and $\widehat{\text{squash}}_S$ form a retract between S and $?$.

LEMMA 3.11 (SPLITTING AND SQUASHING LIFTED RETRACT). *Suppose $(S, ?, \text{split}, \text{squash})$ is the category of skeletons. Then for any skeleton S ,*

$$\widehat{\text{squash}}_A; \widehat{\text{split}}_A = \text{id}_A : A \longrightarrow A$$

PROOF. The proof is similar to the proof of the boxing and unboxing lifted retract (Lemma 3.8). \square

There is also a faithful functor from S to \mathcal{U} where \mathcal{U} is the full subcategory of S that consists of the single object $?$ and all its morphisms between it:

$$\begin{aligned} \mathcal{U}S &= ? \\ \mathcal{U}(f : S_1 \longrightarrow S_2) &= \widehat{\text{split}}_{S_1}; f; \widehat{\text{squash}}_{S_2} \end{aligned}$$

This finally implies that there is a functor $C : C \longrightarrow \mathcal{U}$ that injects all of C into the object $?$.

LEMMA 3.12 (CASTING TO $?$). *Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model, $(S, ?, \text{split}, \text{squash})$ is the full subcategory of skeletons, and $(\mathcal{U}, ?)$ is the full subcategory containing only $?$ and its morphisms. Then there is a faithful functor $C = C \xrightarrow{S} S \xrightarrow{\mathcal{U}} \mathcal{U}$.*

In a way we can think of $C : C \longrightarrow \mathcal{U}$ as a forgetful functor. It forgets the type information.

Getting back the typed information is harder. There is no nice functor from \mathcal{U} to C , because we need more information. However, given a type A we can always obtain a casting morphism from $?$ to A by $(\widehat{\text{split}}_{(\text{skeleton } A)}); (\widehat{\text{unbox}}_A) : ? \longrightarrow A$. Finally, we have the following result.

LEMMA 3.13 (CASTING MORPHISMS TO $?$). *Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model, and A is an object of C . Then there exists casting morphisms from A to $?$ and vice versa that make A a retract of $?$.*

PROOF. The two morphisms are as follows:

$$\text{Box}_A := \widehat{\text{box}}_A; \widehat{\text{squash}}_{(\text{skeleton } A)} : A \longrightarrow ?$$

$$\text{Unbox}_A := \widehat{\text{split}}_{(\text{skeleton } A)}; \widehat{\text{unbox}}_A : ? \longrightarrow A$$

The fact the these form a retract between A and $?$ holds by Lemma 3.8 and Lemma 3.11. \square

3.2 The Interpretation

In this section we show how to interpret $\lambda_{\rightarrow}^?$ and $\lambda_{\rightarrow}^{(A)}$ into the categorical model given in the previous section. We complete the three steps summarized above. We will show how to interpret the typing of the former into the model, and then show how to do the same for the latter, furthermore, we show that reduction can be interpreted into the model as well, thus concluding soundness for $\lambda_{\rightarrow}^{(A)}$ with respect to our model.

First, we must give the interpretation of types and contexts, but this interpretation is obvious, because we have been making sure to match the names of types and objects throughout this paper.

Definition 3.14. Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model. Then we define the interpretation of types into C as follows:

$$\begin{aligned} \llbracket \text{Unit} \rrbracket &= 1 \\ \llbracket \text{Nat} \rrbracket &= \text{Nat} \\ \llbracket ? \rrbracket &= ? \\ \llbracket A_1 \rightarrow A_2 \rrbracket &= \llbracket A_1 \rrbracket \rightarrow \llbracket A_2 \rrbracket \\ \llbracket A_1 \times A_2 \rrbracket &= \llbracket A_1 \rrbracket \times \llbracket A_2 \rrbracket \end{aligned}$$

We extend this interpretation to typing contexts as follows:

$$\begin{aligned} \llbracket \cdot \rrbracket &= 1 \\ \llbracket \Gamma, x : A \rrbracket &= \llbracket \Gamma \rrbracket \times \llbracket A \rrbracket \end{aligned}$$

Throughout the remainder of this paper we will drop the interpretation symbols around types.

Before we can interpret the typing rules of $\lambda_{\rightarrow}^?$ and $\lambda_{\rightarrow}^{(A)}$ we must show how to interpret the consistency relation from Figure 2. These will correspond to casting morphisms.

LEMMA 3.15 (TYPE CONSISTENCY IN THE MODEL). *Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model, and $A \sim B$ for some types A and B . Then there are two casting morphisms $c_1 : A \longrightarrow B$ and $c_2 : B \longrightarrow A$.*

PROOF. This proof holds by induction on the form $A \sim B$ using the morphisms $\text{Box}_A : A \longrightarrow ?$ and $\text{Unbox}_A : ? \longrightarrow A$. Please see Appendix A.3 for the complete proof. \square

COROLLARY 3.16. *Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model. Then we know the following:*

i. *If $A_1 \rightarrow B_1 \sim A_2 \rightarrow B_2$, then there are casting morphisms:*

$$\begin{aligned} c &= c_1 \rightarrow c_2 : (A_1 \rightarrow B_1) \longrightarrow (A_2 \rightarrow B_2) \\ c' &= c_3 \rightarrow c_4 : (A_2 \rightarrow B_2) \longrightarrow (A_1 \rightarrow B_1) \end{aligned}$$

where $c_1 : A_2 \longrightarrow A_1$ and $c_2 : B_1 \longrightarrow B_2$, and $c_3 : A_1 \longrightarrow A_2$ and $c_4 : B_2 \longrightarrow B_1$.

ii. *If $A_1 \times B_1 \sim A_2 \times B_2$, then there are casting morphisms:*

$$\begin{aligned} c &= c_1 \times c_2 : (A_1 \times B_1) \longrightarrow (A_2 \times B_2) \\ c' &= c_3 \times c_4 : (A_2 \times B_2) \longrightarrow (A_1 \times B_1) \end{aligned}$$

where $c_1 : A_1 \longrightarrow A_2$ and $c_2 : B_1 \longrightarrow B_2$, and $c_3 : A_2 \longrightarrow A_1$ and $c_4 : B_2 \longrightarrow B_1$.

PROOF. This proof holds by the construction of the casting morphisms from the proof of the previous result, and the fact that the type consistency rules are unique for each type. \square

Showing that both c_1 and c_2 exist corresponds to the fact that $A \sim B$ is symmetric. But, this interpretation is an over approximation of type consistency, because type consistency is not transitive, but function composition is. Leaving type consistency implicit in the model just does not make good sense categorically, because it would break composition. For example, if we have morphisms $f : A \longrightarrow ?$ and $g : B \longrightarrow C$, then if we implicitly allowed $?$ to be cast to B , then we could compose these two morphisms, but this does not fit the definition of a category, because it requires the target of f to match the source of g , but this just is not the case. Thus, the explicit cast must be used to obtain $f; \widehat{\text{unbox}_B}; g$.

At this point we have everything we need to show our main result which is that typing in both $\lambda_{\rightarrow}^?$ and $\lambda_{\rightarrow}^{\langle A \rangle}$, and evaluation in $\lambda^{\langle A \rangle}$ can be interpreted into the categorical model.

THEOREM 3.17 (INTERPRETATION OF TYPING). *Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model. If $\Gamma \vdash t : A$ in either $\lambda_{\rightarrow}^?$ or $\lambda_{\rightarrow}^{\langle A \rangle}$, then there is a morphism $\llbracket t \rrbracket : \llbracket \Gamma \rrbracket \longrightarrow A$ in C .*

PROOF. The proof holds by induction on the form of $\Gamma \vdash t : A$ and uses most of the results we have developed up to this point. Please see Appendix A.4 for the complete proof. \square

THEOREM 3.18 (INTERPRETATION OF EVALUATION). *Suppose $(\mathcal{T}, C, ?, T, \text{split}, \text{squash}, \text{box}, \text{unbox})$ is a gradual λ -model. If $\Gamma \vdash t_1 \rightsquigarrow t_2 : A$, then $\llbracket t_1 \rrbracket = \llbracket t_2 \rrbracket : \llbracket \Gamma \rrbracket \longrightarrow A$.*

PROOF. This proof holds by induction on the form of $\Gamma \vdash t_1 \rightsquigarrow t_2 : A$, and uses Theorem 3.17, Lemma 3.15, and Corollary 3.16. Please see Appendix A.5 for the complete proof. \square

4 GRADY

Just as the simply typed λ -calculus corresponds to cartesian closed categories our categorical model has a corresponding type theory we call Grady. It consists of all of the structure found in the model. Its syntax is an extension of the syntax for $\lambda_{\rightarrow}^?$.

Definition 4.1. Syntax for Grady:

(basic skeletons)	$U ::= ? \rightarrow ? \mid ? \times ?$
(skeletons)	$S ::= ? \mid S_1 \times S_2 \mid S_1 \rightarrow S_2$
(atomic types)	$C ::= \text{Unit} \mid \text{Nat}$
(terms)	$t ::= \dots \mid \text{split}_U \mid \text{squash}_U \mid \text{box}_C \mid \text{unbox}_C \mid \text{case } t \text{ of } t_1 \parallel x.t_2$
(natural numbers)	$n ::= 0 \mid \text{succ } n$
(simple values)	$s ::= x \mid \text{triv} \mid n \mid \text{squash}_U \mid \text{split}_U \mid \text{box}_C \mid \text{unbox}_C$

The types of Grady are the same as the types of $\lambda_{\rightarrow}^?$ (Definition 2.1), in addition, it encompasses all the terms of $\lambda_{\rightarrow}^?$, and so we do not repeat either of them here. The typing rules for Grady can be found in Figure 6 and its reduction rules can be found in Figure 7.

Just as we did for the categorical model (Lemma 3.13) we can lift box_C and unbox_C to arbitrary type.

LEMMA 4.2 (SYNTACTIC Box_A AND Unbox_A). *Given any type A there are functions Box_A and Unbox_A such that the following typing rules are admissible:*

$$\frac{}{\Gamma \vdash \text{Box}_A : A \rightarrow ?} \text{Box} \qquad \frac{}{\Gamma \vdash \text{Unbox}_A : ? \rightarrow A} \text{Unbox}$$

Furthermore, the following reduction rule is admissible:

$$\frac{\Gamma \vdash t : A}{\Gamma \vdash \text{Unbox}_A (\text{Box}_A t) \rightsquigarrow t : A} \text{retract}_3$$

PROOF. The functions Box_A and Unbox_A can be defined using the construction from the categorical model, e.g. Definition 3.7, Definition 3.10, and Lemma 3.13. However, the categorical notions of composition, identity, and the functors $- \rightarrow -$ and $- \times -$ must be defined as meta-functions first, but after they are, then the same constructions apply. Please see Appendix A.5 for the constructions. \square

$\frac{x : A \in \Gamma}{\Gamma \vdash x : A}$	var	$\frac{}{\Gamma \vdash \text{box}_C : C \rightarrow ?}$	box	$\frac{}{\Gamma \vdash \text{unbox}_C : ? \rightarrow C}$	unbox	$\frac{}{\Gamma \vdash \text{squash}_U : U \rightarrow ?}$	squash
$\frac{}{\Gamma \vdash \text{split}_U : ? \rightarrow U}$	split	$\frac{}{\Gamma \vdash \text{triv} : \text{Unit}}$	unit	$\frac{}{\Gamma \vdash 0 : \text{Nat}}$	zero	$\frac{\Gamma \vdash t : \text{Nat}}{\Gamma \vdash \text{succ } t : \text{Nat}}$	succ
$\frac{\Gamma \vdash t : \text{Nat} \quad \Gamma \vdash t_1 : A \quad \Gamma, x : \text{Nat} \vdash t_2 : A}{\Gamma \vdash \text{case } t \text{ of } t_1 \parallel x.t_2 : A}$	Nat_e	$\frac{\Gamma \vdash t_1 : A_1 \quad \Gamma \vdash t_2 : A_2}{\Gamma \vdash (t_1, t_2) : A_1 \times A_2}$	\times	$\frac{\Gamma \vdash t : A_1 \times A_2}{\Gamma \vdash \text{fst } t : A_1}$	\times_{e_1}	$\frac{\Gamma \vdash t : A_1 \times A_2}{\Gamma \vdash \text{snd } t : A_2}$	\times_{e_2}
$\frac{\Gamma, x : A \vdash t : B}{\Gamma \vdash \lambda x : A_1. t : A \rightarrow B}$	\rightarrow	$\frac{\Gamma \vdash t_1 : A \rightarrow B \quad \Gamma \vdash t_2 : A}{\Gamma \vdash t_1 t_2 : B}$	\rightarrow_e				

Fig. 6. Typing rules for Grady

Perhaps unsurprisingly, due to our results with respect to the categorical model, we can use the previous result to construct a type-directed translation of both $\lambda_{\rightarrow}^?$ and $\lambda_{\rightarrow}^{(A)}$ into Grady.

LEMMA 4.3 (TRANSLATIONS).

- i. If $\Gamma \vdash t : A$ hold in either $\lambda_{\rightarrow}^?$ or $\lambda_{\rightarrow}^{(A)}$, then there exists a term t' such that $\Gamma \vdash t' : A$ holds in Grady.
- ii. If $\Gamma \vdash t_1 \rightsquigarrow t_2 : A$ holds in $\lambda_{\rightarrow}^{(A)}$, then $\Gamma \vdash t'_1 \rightsquigarrow t'_2 : A$ holds in Grady, where $\Gamma \vdash t'_1 : A$ and $\Gamma \vdash t'_2 : A$ are both the corresponding Grady terms.

PROOF. The proof of this result is similar to the proof that both $\lambda_{\rightarrow}^?$ and $\lambda_{\rightarrow}^{(A)}$ can be interpreted into the categorical model, Theorem 3.17 and Theorem 3.18, and thus, we do not give the full proof. The proof of part one holds by induction on $\Gamma \vdash t : A$, and using the realization that if $A \sim B$ for some types A and B then there are casting terms $\cdot \vdash c_1 : A \rightarrow B$ and $\cdot \vdash c_2 : B \rightarrow A$ following the proof of Lemma 3.15. The proof of part two holds by induction on $\Gamma \vdash t_1 \rightsquigarrow t_2 : A$ making use of part one; it is similar to the proof of Theorem 3.18. \square

4.1 Exploiting the Untyped λ -Calculus

Having the untyped λ -calculus along side the typed λ -calculus can be a lot of fun. This section can be seen from two perspectives: it gives a number of examples in Grady, and shows several ways the typed and untyped fragments can be mixed.

Michael is writing this section.

- Church Encoded Data
- Y-combinator and the natural number eliminator, e.g. terminating recursion on natural numbers
- Scott Encoded data, this is not available in terminating type theories
- Parigot Encoded Data, better efficiency

4.2 Extending Grady with Impredicative Polymorphism

The simplistic nature of Grady makes it well suited to extend with new features. In this section we add polymorphism to Grady extending its typed fragment to System F (Girard et al. 1989).

$\frac{\Gamma \vdash s : A}{\Gamma \vdash s \rightsquigarrow s : A}$ values	$\frac{\Gamma \vdash t : C}{\Gamma \vdash \text{unbox}_C(\text{box}_C t) \rightsquigarrow t : C}$ retract ₁	$\frac{\Gamma \vdash t : U}{\Gamma \vdash \text{split}_U(\text{squash}_U t) \rightsquigarrow t : U}$ retract ₂
$\frac{\Gamma, x : A_1 \vdash t_2 : A_2 \quad \Gamma \vdash t_1 : A_1}{\Gamma \vdash (\lambda x : A_1. t_2) t_1 \rightsquigarrow [t_1/x] t_2 : A_2}$ β	$\frac{\Gamma \vdash t : A_1 \rightarrow A_2 \quad x \notin \text{FV}(t)}{\Gamma \vdash \lambda x : A_1. t x \rightsquigarrow t : A_1 \rightarrow A_2}$ η	$\frac{\Gamma \vdash t_1 : A_1 \quad \Gamma \vdash t_2 : A_2}{\Gamma \vdash \text{fst}(t_1, t_2) \rightsquigarrow t_1 : A_1}$ \times_{e_1}
$\frac{\Gamma \vdash t_1 : A_1 \quad \Gamma \vdash t_2 : A_2}{\Gamma \vdash \text{snd}(t_1, t_2) \rightsquigarrow t_2 : A_2}$ \times_{e_2}	$\frac{\Gamma \vdash t : A_1 \times A_2}{\Gamma \vdash (\text{fst } t, \text{snd } t) \rightsquigarrow t : A_1 \times A_2}$ \times_η	$\frac{\Gamma \vdash t \rightsquigarrow t' : \text{Nat}}{\Gamma \vdash \text{succ } t \rightsquigarrow \text{succ } t' : \text{Nat}}$ succ
$\frac{\Gamma \vdash t_1 : A \quad \Gamma, x : \text{Nat} \vdash t_2 : A}{\Gamma \vdash \text{case } 0 \text{ of } t_1 \parallel x. t_2 \rightsquigarrow t_1 : A}$ Nat _{e₀}	$\frac{\Gamma \vdash t : \text{Nat} \quad \Gamma \vdash t_1 : A \quad \Gamma, x : \text{Nat} \vdash t_2 : A}{\Gamma \vdash \text{case } (\text{succ } t) \text{ of } t_1 \parallel x. t_2 \rightsquigarrow [t/x] t_2 : A}$ Nat _{e₁}	
$\frac{\Gamma \vdash t \rightsquigarrow t' : \text{Nat} \quad \Gamma \vdash t_1 : A \quad \Gamma, x : \text{Nat} \vdash t_2 : A}{\Gamma \vdash \text{case } t \text{ of } t_1 \parallel x. t_2 \rightsquigarrow \text{case } t' \text{ of } t_1 \parallel x. t_2 : A}$ case ₁	$\frac{\Gamma \vdash t : \text{Nat} \quad \Gamma \vdash t_1 \rightsquigarrow t'_1 : A \quad \Gamma, x : \text{Nat} \vdash t_2 : A}{\Gamma \vdash \text{case } t \text{ of } t_1 \parallel x. t_2 \rightsquigarrow \text{case } t \text{ of } t'_1 \parallel x. t_2 : A}$ case ₂	
$\frac{\Gamma \vdash t : \text{Nat} \quad \Gamma \vdash t : A \quad \Gamma, x : \text{Nat} \vdash t_2 \rightsquigarrow t'_2 : A}{\Gamma \vdash \text{case } t \text{ of } t_1 \parallel x. t_2 \rightsquigarrow \text{case } t \text{ of } t_1 \parallel x. t'_2 : A}$ case ₃	$\frac{\Gamma, x : A_1 \vdash t \rightsquigarrow t' : A_2}{\Gamma \vdash \lambda x : A_1. t \rightsquigarrow \lambda x : A_1. t' : A_1 \rightarrow A_2}$ \rightarrow	
$\frac{\Gamma \vdash t_1 \rightsquigarrow t'_1 : A_1 \rightarrow A_2 \quad \Gamma \vdash t_2 : A_1}{\Gamma \vdash t_1 t_2 \rightsquigarrow t'_1 t_2 : A_2}$ \rightarrow_{e_1}	$\frac{\Gamma \vdash t_1 : A_1 \rightarrow A_2 \quad \Gamma \vdash t_2 \rightsquigarrow t'_2 : A_1}{\Gamma \vdash t_1 t_2 \rightsquigarrow t_1 t'_2 : A_2}$ \rightarrow_{e_2}	
$\frac{\Gamma \vdash t \rightsquigarrow t' : A_1 \times A_2}{\Gamma \vdash \text{fst } t \rightsquigarrow \text{fst } t' : A_1}$ fst	$\frac{\Gamma \vdash t \rightsquigarrow t' : A_1 \times A_2}{\Gamma \vdash \text{snd } t \rightsquigarrow \text{snd } t' : A_2}$ snd	$\frac{\Gamma \vdash t_1 \rightsquigarrow t'_1 : A_1 \quad \Gamma \vdash t_2 : A_2}{\Gamma \vdash (t_1, t_2) \rightsquigarrow (t'_1, t_2) : A_1 \times A_2}$ \times_1
$\frac{\Gamma \vdash t_1 : A_1 \quad \Gamma \vdash t_2 \rightsquigarrow t'_2 : A_2}{\Gamma \vdash (t_1, t_2) \rightsquigarrow (t_1, t'_2) : A_1 \times A_2}$ \times_2		

Fig. 7. Reduction rules for Grady

Definition 4.4. The syntax of polymorphic Grady as an extension of Grady's is as follows:

(kinds)	$K ::= \star$
(boxable types)	$C ::= \text{Unit} \mid \text{Nat} \mid X$
(types)	$A, B, D, E ::= \dots \mid X \mid \forall X. A$
(terms)	$t ::= \dots \mid \lambda X. t \mid [A]t$
(contexts)	$\Gamma ::= \cdot \mid \Gamma, X : \star \mid \Gamma, x : A$

We renamed the atomic type syntactic category to boxable types, and added type variables to this category. This will allow generic functions to treat the incoming data as untyped. The rules for kinding and well-formed contexts can be found in Figure 8 and Figure 9. The typing rules can be found in Figure 10, but we recount the entire system, because a number of rules had to be changed to insure that the typing context is always well formed. The evaluation rules require far less changes.

$$\begin{array}{c}
\frac{X : \star \in \Gamma}{\Gamma \vdash X : \star} \text{ var} \quad \frac{}{\Gamma \vdash \text{Unit} : \star} \text{ unit} \quad \frac{}{\Gamma \vdash \text{Nat} : \star} \text{ Nat} \quad \frac{}{\Gamma \vdash ? : \star} ? \quad \frac{\Gamma \vdash A : \star \quad \Gamma \vdash B : \star}{\Gamma \vdash A \rightarrow B : \star} \rightarrow \\
\frac{\Gamma \vdash A : \star \quad \Gamma \vdash B : \star}{\Gamma \vdash A \times B : \star} \times \quad \frac{\Gamma, X : \star \vdash A : \star}{\Gamma \vdash \forall X. A : \star} \forall
\end{array}$$

Fig. 8. Kinding rules for Polymorphic Grady

$$\frac{}{\cdot \text{Ok}} \text{ empty} \quad \frac{\Gamma \text{Ok}}{(\Gamma, X : \star) \text{Ok}} \text{ type-var} \quad \frac{\Gamma \text{Ok} \quad \Gamma \vdash A : \star}{(\Gamma, x : A) \text{Ok}} \text{ var}$$

Fig. 9. Well-formed contexts for Polymorphic Grady

$$\begin{array}{c}
\frac{x : A \in \Gamma \quad \Gamma \text{Ok}}{\Gamma \vdash x : A} \text{ var} \quad \frac{\Gamma \text{Ok}}{\Gamma \vdash \text{box}_C : C \rightarrow ?} \text{ box} \quad \frac{\Gamma \text{Ok}}{\Gamma \vdash \text{unbox}_C : ? \rightarrow C} \text{ unbox} \\
\frac{\Gamma \text{Ok}}{\Gamma \vdash \text{squash}_U : U \rightarrow ?} \text{ squash} \quad \frac{\Gamma \text{Ok}}{\Gamma \vdash \text{split}_U : ? \rightarrow U} \text{ split} \quad \frac{\Gamma \text{Ok}}{\Gamma \vdash \text{triv} : \text{Unit}} \text{ unit} \quad \frac{\Gamma \text{Ok}}{\Gamma \vdash 0 : \text{Nat}} \text{ zero} \\
\frac{\Gamma \vdash t : \text{Nat}}{\Gamma \vdash \text{succ } t : \text{Nat}} \text{ succ} \quad \frac{\Gamma \vdash t : \text{Nat} \quad \Gamma \vdash t_1 : A \quad \Gamma, x : \text{Nat} \vdash t_2 : A}{\Gamma \vdash \text{case } t \text{ of } t_1 \parallel x. t_2 : A} \text{ Nat}_e \quad \frac{\Gamma \vdash t_1 : A_1 \quad \Gamma \vdash t_2 : A_2}{\Gamma \vdash (t_1, t_2) : A_1 \times A_2} \times \\
\frac{\Gamma \vdash t : A_1 \times A_2}{\Gamma \vdash \text{fst } t : A_1} \times_{e_1} \quad \frac{\Gamma \vdash t : A_1 \times A_2}{\Gamma \vdash \text{snd } t : A_2} \times_{e_2} \quad \frac{\Gamma, x : A \vdash t : B}{\Gamma \vdash \lambda x : A. t : A \rightarrow B} \rightarrow \\
\frac{\Gamma \vdash t_1 : A \rightarrow B \quad \Gamma \vdash t_2 : A}{\Gamma \vdash t_1 t_2 : B} \rightarrow_e \quad \frac{\Gamma, X : \star \vdash t : A}{\Gamma \vdash \Lambda X. t : \forall X. A} \forall \quad \frac{\Gamma \vdash t : \forall X. B \quad \Gamma \vdash A : \star}{\Gamma \vdash [A]t : [A/X]B} \forall_e
\end{array}$$

Fig. 10. Typing rules for Polymorphic Grady

Definition 4.5. The evaluation rules of polymorphic Grady (Definition 7) are an extension of Grady's by adding the following rules:

$$\frac{\Gamma, X : \star \vdash t : B}{\Gamma \vdash [A](\Lambda X. t) \rightsquigarrow [A/X]t : [A/X]B} \text{ type}_\beta \quad \frac{\Gamma, X : \star \vdash t_1 \rightsquigarrow t_2 : A}{\Gamma \vdash \Lambda X. t_1 \rightsquigarrow \Lambda X. t_2 : \forall X. A} \forall \quad \frac{\Gamma \vdash t_1 \rightsquigarrow t_2 : \forall X. B}{\Gamma \vdash [A]t_1 \rightsquigarrow [A]t_2 : [A/X]B} \forall_e$$

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4 A PROOFS

6 A.1 Proof of Lifted Retract (Lemma 3.8)

7 This is a proof by induction on the form of A .

8 Case. Suppose A is atomic. Then:

$$10 \quad \widehat{\text{box}}_A; \widehat{\text{unbox}}_A = \text{box}_A; \text{unbox}_A = \text{id}_A$$

11 Case. Suppose A is $?$. Then:

$$\begin{aligned} 12 \quad \widehat{\text{box}}_A; \widehat{\text{unbox}}_A &= \widehat{\text{box}}_?; \widehat{\text{unbox}}_? \\ 13 &= \text{id}_?; \text{id}_? \\ 14 &= \text{id}_? \\ 15 &= \text{id}_A \end{aligned}$$

16 Case. Suppose $A = A_1 \rightarrow A_2$. Then:

$$\begin{aligned} 17 \quad \widehat{\text{box}}_A; \widehat{\text{unbox}}_A &= \widehat{\text{box}}_{(A_1 \rightarrow A_2)}; \widehat{\text{unbox}}_{(A_1 \rightarrow A_2)} \\ 18 &= (\widehat{\text{unbox}}_{A_1} \rightarrow \widehat{\text{box}}_{A_2}); (\widehat{\text{box}}_{A_1} \rightarrow \widehat{\text{box}}_{A_2}) \\ 19 &= (\widehat{\text{box}}_{A_1}; \widehat{\text{unbox}}_{A_1}) \rightarrow (\widehat{\text{box}}_{A_2}; \widehat{\text{unbox}}_{A_2}) \end{aligned}$$

20 By two applications of the induction hypothesis we know the following:

$$21 \quad \widehat{\text{box}}_{A_1}; \widehat{\text{unbox}}_{A_1} = \text{id}_{A_1} \quad \text{and} \quad \widehat{\text{box}}_{A_2}; \widehat{\text{unbox}}_{A_2} = \text{id}_{A_2}$$

22 Thus, we know the following:

$$\begin{aligned} 23 \quad (\widehat{\text{box}}_{A_1}; \widehat{\text{unbox}}_{A_1}) \rightarrow (\widehat{\text{box}}_{A_2}; \widehat{\text{unbox}}_{A_2}) &= \text{id}_{A_1} \rightarrow \text{id}_{A_2} \\ 24 &= \text{id}_{A_1 \rightarrow A_2} \\ 25 &= \text{id}_A \end{aligned}$$

26 Case. Suppose $A = A_1 \times A_2$. Then:

$$\begin{aligned} 27 \quad \widehat{\text{box}}_A; \widehat{\text{unbox}}_A &= \widehat{\text{box}}_{(A_1 \times A_2)}; \widehat{\text{unbox}}_{(A_1 \times A_2)} \\ 28 &= (\widehat{\text{box}}_{A_1} \times \widehat{\text{box}}_{A_2}); (\widehat{\text{unbox}}_{A_1} \times \widehat{\text{unbox}}_{A_2}) \\ 29 &= (\widehat{\text{box}}_{A_1}; \widehat{\text{unbox}}_{A_1}) \times (\widehat{\text{box}}_{A_2}; \widehat{\text{unbox}}_{A_2}) \end{aligned}$$

30 By two applications of the induction hypothesis we know the following:

$$31 \quad \widehat{\text{box}}_{A_1}; \widehat{\text{unbox}}_{A_1} = \text{id}_{A_1} \quad \text{and} \quad \widehat{\text{box}}_{A_2}; \widehat{\text{unbox}}_{A_2} = \text{id}_{A_2}$$

32 Thus, we know the following:

$$\begin{aligned} 33 \quad (\widehat{\text{box}}_{A_1}; \widehat{\text{unbox}}_{A_1}) \times (\widehat{\text{box}}_{A_2}; \widehat{\text{unbox}}_{A_2}) &= \text{id}_{A_1} \times \text{id}_{A_2} \\ 34 &= \text{id}_{A_1 \times A_2} \\ 35 &= \text{id}_A \end{aligned}$$

A.2 Proof of Lemma 3.9

We must show that the function

$$S_{A,B} : \text{Hom}_C(A, B) \longrightarrow \text{Hom}_S(SA, SB)$$

is injective.

So suppose $f \in \text{Hom}_C(A, B)$ and $g \in \text{Hom}_C(A, B)$ such that $Sf = Sg : SA \longrightarrow SB$. Then we can easily see that:

$$\begin{aligned} Sf &= \widehat{\text{unbox}_A; f; \text{box}_B} \\ &= \widehat{\text{unbox}_A; g; \text{box}_B} \\ &= Sg \end{aligned}$$

But, we have the following equalities:

$$\begin{aligned} \widehat{\text{unbox}_A; f; \text{box}_B} &= \widehat{\text{unbox}_A; g; \text{box}_B} \\ \widehat{\text{box}_A; \widehat{\text{unbox}_A; f; \text{box}_B}; \text{unbox}_B} &= \widehat{\text{box}_A; \widehat{\text{unbox}_A; g; \text{box}_B}; \text{unbox}_B} \\ \widehat{\text{id}_A; f; \text{box}_B; \text{unbox}_B} &= \widehat{\text{id}_A; g; \text{box}_B; \text{unbox}_B} \\ \widehat{\text{id}_A; f; \text{id}_B} &= \widehat{\text{id}_A; g; \text{id}_B} \\ f &= g \end{aligned}$$

The previous equalities hold due to Lemma 3.8.

A.3 Proof of Type Consistency in the Model (Lemma 3.15)

This is a proof by induction on the form of $A \sim B$.

Case.

$$\overline{A \sim A}$$

Choose $c_1 = c_2 = \text{id}_A : A \longrightarrow A$.

Case.

$$\overline{A \sim ?}$$

Choose $c_1 = \text{Box}_A : A \longrightarrow ?$ and $c_2 = \text{Unbox}_A : ? \rightarrow A$.

Case.

$$\overline{? \sim A}$$

Choose $c_1 = \text{Unbox}_A : ? \longrightarrow A$ and $c_2 = \text{Box}_A : A \rightarrow ?$.

Case.

$$\frac{A_1 \sim A_2 \quad B_1 \sim B_2}{A_1 \rightarrow B_1 \sim A_2 \rightarrow B_2}$$

By the induction hypothesis there exists four casting morphisms $c'_1 : A_1 \longrightarrow A_2$, $c'_2 : A_2 \longrightarrow A_1$, $c'_3 : B_1 \longrightarrow B_2$, and $c'_4 : B_2 \longrightarrow B_1$. Choose $c_1 = c'_2 \rightarrow c'_3 : (A_1 \rightarrow B_1) \longrightarrow (A_2 \rightarrow B_2)$ and $c_2 = c'_1 \rightarrow c'_4 : (A_2 \rightarrow B_2) \longrightarrow (A_1 \rightarrow B_1)$.

Case.

$$\frac{A_1 \sim A_2 \quad B_1 \sim B_2}{A_1 \times B_1 \sim A_2 \times B_2}$$

By the induction hypothesis there exists four casting morphisms $c'_1 : A_1 \longrightarrow A_2$, $c'_2 : A_2 \longrightarrow A_1$, $c'_3 : B_1 \longrightarrow B_2$, and $c'_4 : B_2 \longrightarrow B_1$. Choose $c_1 = c'_1 \times c'_3 : A_1 \times B_1 \longrightarrow A_2 \times B_2$ and $c_2 = c'_2 \times c'_4 : A_2 \times B_2 \longrightarrow A_1 \times B_1$.

A.4 Proof of Interpretation of Types Theorem 3.17

This is a proof by induction on $\Gamma \vdash t : A$. First, we show how to interpret the rules of $\lambda_{\rightarrow}^?$, and then $\lambda_{\rightarrow}^{\langle A \rangle}$.

Case.

$$\frac{x : A \in \Gamma}{\Gamma \vdash x : A} \text{ var}$$

Suppose with out loss of generality that $\llbracket \Gamma \rrbracket = A_1 \times \cdots \times A_i \times \cdots \times A_j$ where $A_i = A$. We know that $j > 0$ or the assumed typing derivation would not hold. Then take $\llbracket x \rrbracket = \pi_i : \llbracket \Gamma \rrbracket \rightarrow A$.

Case.

$$\overline{\Gamma \vdash \text{triv} : \text{Unit}} \text{ unit}$$

Take $\llbracket \text{triv} \rrbracket = \diamond_{\llbracket \Gamma \rrbracket} : \llbracket \Gamma \rrbracket \rightarrow 1$ where $\diamond_{\llbracket \Gamma \rrbracket}$ is the unique terminal arrow that exists because C is cartesian closed.

Case.

$$\overline{\Gamma \vdash 0 : \text{Nat}} \text{ zero}$$

Take $\llbracket 0 \rrbracket = \diamond_{\llbracket \Gamma \rrbracket} : \llbracket \Gamma \rrbracket \rightarrow \text{Nat}$ where $z : 1 \rightarrow \text{Nat}$ exists because C contains a SNNO.

Case.

$$\frac{\Gamma \vdash t : \text{Nat}}{\Gamma \vdash \text{succ } t : \text{Nat}} \text{ succ}$$

By the induction hypothesis there is a morphism

$\llbracket t \rrbracket : \llbracket \Gamma \rrbracket \rightarrow \text{Nat}$. Then take $\llbracket \text{succ } t \rrbracket = \llbracket t \rrbracket; \text{succ} : \llbracket \Gamma \rrbracket \rightarrow \text{Nat}$, where $\text{succ} : \text{Nat} \rightarrow \text{Nat}$ exists because C has a SNNO.

Case.

$$\frac{\Gamma \vdash t_1 : A_1 \quad \Gamma \vdash t_2 : A_2}{\Gamma \vdash (t_1, t_2) : A_1 \times A_2} \times$$

By two applications of the induction hypothesis there are two morphisms $\llbracket t_1 \rrbracket : \llbracket \Gamma \rrbracket \rightarrow A$ and $\llbracket t_2 \rrbracket : \llbracket \Gamma \rrbracket \rightarrow B$. Then using the fact that C is cartesian we take $\llbracket (t_1, t_2) \rrbracket = \langle \llbracket t_1 \rrbracket, \llbracket t_2 \rrbracket \rangle : \llbracket \Gamma \rrbracket \rightarrow A \times B$.

Case.

$$\frac{\Gamma \vdash t : A_1 \times B \quad A_1 \sim A_2}{\Gamma \vdash \text{fst } t : A_2} \times_{e_1}$$

By the induction hypothesis there is a morphism

$\llbracket t \rrbracket : \llbracket \Gamma \rrbracket \rightarrow A_1 \times B$, and by type consistency in the model (Lemma 3.15) there exists a casting morphism $c_1 : A_1 \rightarrow A_2$. Finally, take $\llbracket \text{fst } t \rrbracket = \llbracket t \rrbracket; \pi_1; c_1 : \llbracket \Gamma \rrbracket \rightarrow A_2$.

Case.

$$\frac{\Gamma \vdash t : A \times B_1 \quad B_1 \sim B_2}{\Gamma \vdash \text{snd } t : B_2} \times_{e_2}$$

By the induction hypothesis there is a morphism

$\llbracket t \rrbracket : \llbracket \Gamma \rrbracket \rightarrow A \times B_1$, and by type consistency in the model (Lemma 3.15) there exists a casting morphism $c_1 : B_1 \rightarrow B_2$. Finally, take $\llbracket \text{snd } t \rrbracket = \llbracket t \rrbracket; \pi_2; c_1 : \llbracket \Gamma \rrbracket \rightarrow B_2$.

Case.

$$\frac{\Gamma, x : A \vdash t : B}{\Gamma \vdash \lambda x : A_1. t : A \rightarrow B} \rightarrow$$

By the induction hypothesis there is a morphism $\llbracket t \rrbracket : \llbracket \Gamma \rrbracket \times A \rightarrow B$. Then take $\llbracket \lambda x : A. t \rrbracket = \text{curry}(\llbracket t \rrbracket) : \llbracket \Gamma \rrbracket \rightarrow (A \rightarrow B)$, where $\text{curry} : \text{Hom}_C(X \times Y, Z) \rightarrow \text{Hom}_C(X, Y \rightarrow Z)$ exists because C is closed.

Case.

$$\frac{\Gamma \vdash t_1 : A_1 \rightarrow B \quad \Gamma \vdash t_2 : A_2 \quad A_1 \sim A_2}{\Gamma \vdash t_1 t_2 : B} \rightarrow_e$$

By two applications of the induction hypothesis there are morphisms $\llbracket t_1 \rrbracket : \llbracket \Gamma \rrbracket \rightarrow (A_1 \rightarrow B)$ and $\llbracket t_2 \rrbracket : \llbracket \Gamma \rrbracket \rightarrow A_2$, and by type consistency in the model (Lemma 3.15) there exists a casting morphism $c_2 : A_2 \rightarrow A_1$. Then take $\llbracket t_1 t_2 \rrbracket = \langle \llbracket t_1 \rrbracket, \llbracket t_2 \rrbracket; c_1 \rangle; \text{app}_{A,B} : \llbracket \Gamma \rrbracket \rightarrow B$. The morphism $\text{app}_{A,B} : (A \rightarrow B) \times A \rightarrow B$ exists because C is closed.

Case.

$$\frac{\Gamma \vdash t : ?}{\Gamma \vdash \text{succ } t : ?} \text{succ}^?$$

By the induction hypothesis there is a morphism $\llbracket t \rrbracket : \llbracket \Gamma \rrbracket \rightarrow ?$. In addition, we know that $? \sim \text{Nat}$ always holds, and its casing morphisms are $\text{unbox}_{\text{Nat}} : ? \rightarrow \text{Nat}$ and $\text{box}_{\text{Nat}} : \text{Nat} \rightarrow ?$. Thus, take $\llbracket \text{succ } t \rrbracket = \llbracket t \rrbracket; \text{unbox}_{\text{Nat}}; \text{succ}; \text{box}_{\text{Nat}}$.

Case.

$$\frac{\Gamma \vdash t : ?}{\Gamma \vdash \text{fst } t : ?} \times_{e_1}^?$$

By the induction hypothesis there is a morphism $\llbracket t \rrbracket : \llbracket \Gamma \rrbracket \rightarrow ?$. In addition, we know that $? \sim ? \times ?$, and its casting morphisms are $\text{split}_{? \times ?} : ? \rightarrow ? \times ?$ and $\text{squash}_{? \times ?} : ? \times ? \rightarrow ?$. Then take $\llbracket \text{fst } t \rrbracket = \llbracket t \rrbracket; \text{split}_{? \times ?}; \pi_1 : \llbracket \Gamma \rrbracket \rightarrow ?$.

Case.

$$\frac{\Gamma \vdash t : ?}{\Gamma \vdash \text{snd } t : ?} \times_{e_2}^?$$

By the induction hypothesis there is a morphism $\llbracket t \rrbracket : \llbracket \Gamma \rrbracket \rightarrow ?$. In addition, we know that $? \sim ? \times ?$, and its casting morphisms are $\text{split}_{? \times ?} : ? \rightarrow ? \times ?$ and $\text{squash}_{? \times ?} : ? \times ? \rightarrow ?$. Then take $\llbracket \text{snd } t \rrbracket = \llbracket t \rrbracket; \text{split}_{? \times ?}; \pi_2 : \llbracket \Gamma \rrbracket \rightarrow ?$.

Case.

$$\frac{\Gamma \vdash t_1 : ? \quad \Gamma \vdash t_2 : A}{\Gamma \vdash t_1 t_2 : ?} \rightarrow_e^?$$

By the induction hypothesis there are morphisms $\llbracket t_1 \rrbracket : \llbracket \Gamma \rrbracket \rightarrow ?$ and $\llbracket t_2 \rrbracket : \llbracket \Gamma \rrbracket \rightarrow A$. In addition, we know that $? \sim ? \rightarrow ?$ and $? \sim A$, and the casting morphisms are $\text{split}_{? \rightarrow ?} : ? \rightarrow (? \rightarrow ?)$ and $\text{squash}_{? \rightarrow ?} : (? \rightarrow ?) \rightarrow ?$, and $\text{Unbox}_A : ? \rightarrow A$ and $\text{Box}_A : A \rightarrow ?$. Then take $\llbracket t_1 t_2 \rrbracket = \langle \llbracket t_1 \rrbracket; \text{split}_{? \rightarrow ?}, \llbracket t_2 \rrbracket; \text{Box}_A \rangle; \text{app}_{?,?}$.

Next we turn to $\lambda_{\rightarrow}^{(A)}$, but we do not show every rule, because it corresponds to the simply typed λ -calculus whose interpretation is similar to what we have already shown above except without casting morphism, and so we only show the case for the cast rule.

Case.

$$\frac{\Gamma \vdash t : A \quad A \sim B}{\Gamma \vdash \langle B \rangle t : B} \text{cast}$$

By the induction hypothesis there is a morphism $\llbracket t \rrbracket : \llbracket \Gamma \rrbracket \rightarrow A$, and by type consistency in the model (Lemma 3.15) there is a casting morphism $c_1 : A \rightarrow B$. So take $\llbracket \langle B \rangle t \rrbracket = \llbracket t \rrbracket; c_1 : \llbracket \Gamma \rrbracket \rightarrow B$.

A.5 Proof of Interpretation of Evaluation (Theorem 3.18)

This proof holds by induction on the form of $\Gamma \vdash t_1 \rightsquigarrow t_2 : A$. We only show the cases for the casting rules, because the others are well-known to hold within any cartesian closed category; see (Lambek 1980) or (Crole 1994). We will routinely use Theorem 3.17 throughout this proof without mention.

Case.

$$\frac{\Gamma \vdash v : A}{\Gamma \vdash v \rightsquigarrow v : A} \text{value}$$

This case is trivial.

Case.

$$\frac{\Gamma \vdash \text{drop-cast } v : C}{\Gamma \vdash \langle C \rangle v \rightsquigarrow \text{drop-cast } v : C} \text{value-cast}$$

Either $\text{drop-cast } v = v$ which implies that v is a simple value, or $v = \langle ? \rangle s$ for some simple value s , and $\text{drop-cast } v = s$.

If the former is true, then it must be the case that $\Gamma \vdash v : C$, and $\Gamma \vdash \langle C \rangle v : C$ holds using the casting rule $C \sim C$, and its casting morphism is $\text{id}_C : C \longrightarrow C$. Thus, $\llbracket \langle C \rangle v \rrbracket = \llbracket v \rrbracket; \text{id}_C = \llbracket v \rrbracket = \llbracket \text{drop-cast } v \rrbracket$.

Now suppose $v = \langle ? \rangle s$ for some simple value s . Then it must be the case that the typing derivation of $\langle C \rangle v$ has the following form:

$$\frac{\frac{\Gamma \vdash s : B \quad \overline{B \sim ?}}{\Gamma \vdash \langle ? \rangle s : ?} \quad \overline{? \sim C}}{\Gamma \vdash \langle C \rangle \langle ? \rangle s : C}$$

This implies that

$$\begin{aligned} \llbracket \langle C \rangle \langle ? \rangle s \rrbracket &= \llbracket s \rrbracket; \text{Box}_B; \text{Unbox}_B \\ &= \llbracket s \rrbracket; \text{id}_B \\ &= \llbracket s \rrbracket \\ &= \llbracket \text{drop-cast } v \rrbracket \end{aligned}$$

The previous equality holds by Lemma 3.13.

Case.

$$\frac{\Gamma \vdash t : ?}{\Gamma \vdash \langle \text{Nat} \rangle (\text{succ } t) \rightsquigarrow \text{succ } \langle \text{Nat} \rangle t : \text{Nat}} \text{Nat-cast}$$

It must be the case that the typing derivation of $\langle \text{Nat} \rangle (\text{succ } t)$ ends as follows:

$$\frac{\frac{\Gamma \vdash t : ?}{\Gamma \vdash \text{succ } t : ?} \quad \overline{? \sim \text{Nat}}}{\Gamma \vdash \langle \text{Nat} \rangle (\text{succ } t) : \text{Nat}}$$

Then

$$\begin{aligned}
 \llbracket \langle \text{Nat} \rangle (\text{succ } t) \rrbracket &= \llbracket \text{succ } t \rrbracket; \text{unbox}_{\text{Nat}} \\
 &= \langle \llbracket t \rrbracket; \text{unbox}_{\text{Nat}}; \text{succ}; \text{box}_{\text{Nat}} \rangle; \text{unbox}_{\text{Nat}} \\
 &= \llbracket t \rrbracket; (\text{unbox}_{\text{Nat}}; \text{succ}; \text{box}_{\text{Nat}}; \text{unbox}_{\text{Nat}}) \\
 &= \llbracket t \rrbracket; (\text{unbox}_{\text{Nat}}; \text{succ}; \text{id}_{\text{Nat}}) \\
 &= \llbracket t \rrbracket; (\text{unbox}_{\text{Nat}}; \text{succ}) \\
 &= \langle \llbracket t \rrbracket; \text{unbox}_{\text{Nat}} \rangle; \text{succ} \\
 &= \llbracket \text{succ } \langle \text{Nat} \rangle t \rrbracket
 \end{aligned}$$

The previous equality holds by Lemma 3.13 and because it must be the case that the typing derivation of $\text{succ } \langle \text{Nat} \rangle t$ ends as follows:

$$\frac{\frac{\Gamma \vdash t : ? \quad ? \sim \text{Nat}}{\Gamma \vdash \langle \text{Nat} \rangle t : \text{Nat}}}{\Gamma \vdash \text{succ } \langle \text{Nat} \rangle t : \text{Nat}}$$

Case.

$$\frac{\Gamma \vdash t : A_1 \rightarrow B_1 \quad (A_1 \rightarrow B_1) \sim (A_2 \rightarrow B_2)}{\Gamma \vdash \langle A_2 \rightarrow B_2 \rangle t \rightsquigarrow \lambda y : A_2. \langle B_2 \rangle (t \langle A_1 \rangle y) : A_2 \rightarrow B_2} \rightarrow\text{-cast}$$

This case requires the following basic results about cartesian closed categories. First, there is a natural bijection:

$$\text{curry} : \text{Hom}_C(X \times Y, Z) \longrightarrow \text{Hom}_C(X, Y \rightarrow Z)$$

This bijection implies the following equalities:

$$\begin{aligned}
 \text{curry}(\text{app}_{X,Y}) &= \text{id}_{X \rightarrow Y} \\
 \text{curry}((f \times g); x; h) &= f; \text{curry}(x); (g \rightarrow h)
 \end{aligned}$$

We know from type consistency in the model (Lemma 3.15 and Corollary 3.16) that there is a casting morphism $c_1 \rightarrow c_2 : (A_1 \rightarrow B_1) \longrightarrow (A_2 \rightarrow B_2)$ where $c_1 : A_2 \longrightarrow A_1$ and $c_2 : B_1 \longrightarrow B_2$.

It suffices to show that:

$$\begin{aligned}
 \llbracket \langle A_2 \rightarrow B_2 \rangle t \rrbracket &= \llbracket t \rrbracket; (c_1 \rightarrow c_2) \\
 &= \text{curry}(\langle \llbracket t \rrbracket \times c_1 \rangle; \text{app}_{A_1, B_1}; c_2) \\
 &= \llbracket \lambda y : A_2. \langle B_2 \rangle (t \langle A_1 \rangle y) \rrbracket
 \end{aligned}$$

We prove this equality from right to left as follows:

$$\begin{aligned}
 \llbracket \lambda y : A_2. \langle B_2 \rangle (t \langle A_1 \rangle y) \rrbracket &= \text{curry}(\langle \llbracket t \rrbracket \times c_1 \rangle; \text{app}_{A_1, B_1}; c_2) \\
 &= \llbracket t \rrbracket; \text{curry}(\text{app}_{A_1, B_1}); (c_1 \rightarrow c_2) \\
 &= \llbracket t \rrbracket; \text{id}_{A_1 \rightarrow B_1}; (c_1 \rightarrow c_2) \\
 &= \llbracket t \rrbracket; (c_1 \rightarrow c_2) \\
 &= \llbracket \langle A_2 \rightarrow B_2 \rangle t \rrbracket
 \end{aligned}$$

Case.

$$\frac{\Gamma \vdash t : A_1 \times B_1 \quad (A_1 \times B_1) \sim (A_2 \times B_2)}{\Gamma \vdash \langle A_2 \times B_2 \rangle t \rightsquigarrow (\langle A_2 \rangle (\text{fst } t), \langle B_2 \rangle (\text{snd } t)) : A_2 \times B_2} \times\text{-cast}$$

We know from type consistency in the model (Lemma 3.15 and Corollary 3.16) that there is a casting morphism $c_1 \times c_2 : (A_1 \times B_1) \longrightarrow (A_2 \times B_2)$ where $c_1 : A_1 \longrightarrow A_2$ and $c_2 : B_1 \longrightarrow B_2$.

It suffices to show that:

$$\begin{aligned} \llbracket \langle A_2 \times B_2 \rangle t \rrbracket &= \llbracket t \rrbracket; (c_1 \times c_2) \\ &= \langle \llbracket t \rrbracket; \pi_1; c_1, \llbracket t \rrbracket; \pi_2; c_2 \rangle \\ &= \llbracket (\langle A_2 \rangle (\text{fst } t), \langle B_2 \rangle (\text{snd } t)) \rrbracket \end{aligned}$$

This equality holds using the following well known fact on cartesian categories:

$$\langle f; g, f; h \rangle = f; (g \times h)$$

Case.

$$\frac{\Gamma \vdash t_1 \rightsquigarrow t_2 : A \quad A \sim B}{\Gamma \vdash \langle B \rangle t_1 \rightsquigarrow \langle B \rangle t_2 : B} \text{cast}$$

By the induction hypothesis we know that $\llbracket t_1 \rrbracket = \llbracket t_2 \rrbracket : \llbracket \Gamma \rrbracket \longrightarrow A$, and by type consistency in the model (Lemma 3.15) there is a casting morphism $c : A \longrightarrow B$.

Then it suffices to show that:

$$\begin{aligned} \llbracket \langle B \rangle t_1 \rrbracket &= \llbracket t_1 \rrbracket; c \\ &= \llbracket t_2 \rrbracket; c \\ &= \llbracket \langle B \rangle t_2 \rrbracket \end{aligned}$$

But, this clearly holds by the fact that $\llbracket t_1 \rrbracket = \llbracket t_2 \rrbracket : \llbracket \Gamma \rrbracket \longrightarrow A$.

A.6 Proof of Lemma 4.2

First, we define the identify meta-function:

$$\text{id}_A := \lambda x : A. x$$

Then composition. Suppose $\Gamma \vdash t_1 : A \rightarrow B$ and $\Gamma \vdash t_2 : B \rightarrow D$ are two terms, then we define:

$$t_1; t_2 := \lambda x : A. t_2 (t_1 x)$$

It is easy to see that the following rule is admissible:

$$\frac{\Gamma \vdash t_1 : A \rightarrow B \quad \Gamma \vdash t_2 : B \rightarrow D}{\Gamma \vdash t_1; t_2 : A \rightarrow D} \text{comp}$$

The functor $- \times -$ requires two morphisms $\Gamma \vdash t_1 : A \rightarrow D$ and $\Gamma \vdash t_2 : B \rightarrow E$, and is defined as follows:

$$t_1 \times t_2 := \lambda x : A \times B. (t_1 (\text{fst } x), t_2 (\text{snd } x))$$

The following rule is admissible:

$$\frac{\Gamma \vdash t_1 : A \rightarrow D \quad \Gamma \vdash t_2 : B \rightarrow E}{\Gamma \vdash t_1 \times t_2 : A \times B \rightarrow D \times E} \text{prod}$$

The functor $- \rightarrow -$ requires two morphisms $\Gamma \vdash t_1 : D \rightarrow A$ and $\Gamma \vdash t_2 : B \rightarrow E$, and is defined as follows:

$$t_1 \rightarrow t_2 := \lambda f : A \rightarrow B. \lambda y : D. t_2 (f (t_1 y))$$

The following rule is admissible:

$$\frac{\Gamma \vdash t_1 : D \rightarrow A \quad \Gamma \vdash t_2 : B \rightarrow E}{\Gamma \vdash t_1 \rightarrow t_2 : (A \rightarrow B) \rightarrow (D \rightarrow E)} \text{prod}$$

At this point it is straightforward to carry out the definition of Box_A and Unbox_A using the definitions from the model. Showing the admissibility of the typing and reduction rules follows by induction on A .