

CPL: A Core Language for Cloud Computing

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Abstract

Running distributed applications in the cloud involves deployment. That is, distribution and configuration of application services and middleware infrastructure. The considerable complexity of these tasks resulted in the emergence of declarative JSON-based domain-specific *deployment languages* to develop *deployment programs*. However, existing deployment programs unsafely compose artifacts written in different languages, leading to bugs that are hard to detect before run time. Furthermore, deployment languages do not provide extension points for custom implementations of existing cloud services such as application-specific load balancing policies.

To address these shortcomings, we propose CPL (*Cloud Platform Language*), a statically-typed core language for programming both distributed applications as well as their deployment on a cloud platform. In CPL, application services and deployment programs interact through statically typed, extensible interfaces, and an application can trigger further deployment at run time. We provide a formal semantics of CPL and demonstrate that it enables type-safe, composable and extensible libraries of *service combinators*, such as load balancing and fault tolerance.

1. Introduction

Cloud computing [30] has emerged as the reference infrastructure for concurrent distributed services with high availability, resilience and quick response times, providing access to on-demand and location-transparent computing resources. Companies develop and run distributed applications on specific cloud platforms, e.g., Amazon AWS¹ or Google Cloud

Platform.² Services are bought as needed from the cloud provider in order to adapt to customer demand,

An important and challenging task in the development process of cloud applications is *deployment*. Especially, deployment involves the distribution, configuration and composition of (1) virtual machines that implement the application and its services, and of (2) virtual machines that provide middleware infrastructure such as load balancing, key-value stores, and MapReduce. Deploying a cloud application can go wrong and cause the application to malfunction. Possible causes are software bugs in the application itself, but also wrong configurations, such as missing library dependencies or inappropriate permissions for a shell script. Fixing mistakes *after* deployment causes high costs and loss of reputation. For example, in 2012, Knight Capital lost over \$440 Million over the course of 30 minutes due to a bug in its deployed trading software,³ causing the disappearance of the company from the market.

Considering that cloud applications can have deployment sizes in the hundreds or thousands of virtual machines, manual configuration is error-prone and does not scale. Cloud platforms address this issue with domain-specific languages (DSLs) such as Amazon CloudFormation or Google Cloud Deployment Manager. The purpose of these DSLs is to write reusable *deployment programs*, which instruct the cloud platform to perform deployment steps automatically. A typical deployment program specifies the required virtual machines for the application infrastructure, how these virtual machines are connected with each other, and how the application infrastructure connects to the pre-existing or newly created middleware infrastructure of the cloud platform.

However, the modularity of current cloud deployment DSLs is insufficient (detailed discussion in Section 2):

Unsafe Composition: Application services and deployment programs are written in different languages. Deployment DSLs configure application services by lexically expanding configuration parameters into application source code before its execution. This approach is similar to a lex-

¹<https://aws.amazon.com>

²<https://cloud.google.com>

³<http://www.bloomberg.com/bw/articles/2012-08-02/knight-shows-how-to-lose-440-million-in-30-minutes>.

ical macro system and makes deployment programs unsafe because of unintentional code injection and lexical incompatibilities.

No Extensibility: Middleware cloud services (e.g., elastic load balancing, which may dynamically allocate new virtual machines) are pre-defined in the cloud platform and only referenced by the deployment program through external interfaces. As such, there is no way to customize those services during deployment or extend them with additional features.

Stage Separation: Current deployment DSLs finish their execution before the application services are started. Therefore, it is impossible to change the deployment once the application stage is active. Thus, applications cannot self-adjust their own deployment, e.g., to react to time-varying customer demand.

We propose CPL (Cloud Platform Language), a statically-typed core language for programming cloud applications and deployments. CPL employs techniques from programming language design and type systems to overcome the issues outlined above. Most importantly, CPL unifies the programming of deployments and applications into a single language. This avoids unsafe composition because deployments and applications can exchange values directly via statically typed interfaces. For extensibility, CPL supports higher-order service combinators with statically typed interfaces using bounded polymorphism. Finally, CPL programs run at a single stage where an application service can trigger further deployment.

To demonstrate how CPL solves the problems of deployment languages, we implemented a number of case studies. First, we demonstrate type-safe composition through generically typed worker and thunk abstractions. Second, on top of the worker abstraction, we define composable and reusable *service combinators* in CPL, which add new features, such as elastic load balancing and fault tolerance. Finally, we demonstrate how to model MapReduce as a deployment program in CPL and apply our combinators, obtaining different MapReduce variants, which safely deploy at run time.

In summary, we make the following contributions:

- We analyze the problems with current cloud deployment DSLs.
- We define the formal syntax and semantics of CPL to model cloud platforms as distributed, asynchronous message passing systems. Our design is inspired by the Join Calculus [12].
- We define the type system of CPL as a variant of System F with bounded quantification [24].
- We formalize CPL in PLT Redex [9] and we provide a concurrent implementation in Scala.
- We evaluated CPL with case studies, including a library of typed service combinators that model elastic load balancing and fault tolerance mechanisms. Also, we apply the combinators to a MapReduce deployment specification.

```

1 { //...
2 "Parameters": {
3   "InstanceType": {
4     "Description": "WebServer EC2 instance type",
5     "Type": "String",
6     "Default": "t2.small",
7     "AllowedValues": [ "t2.micro", "t2.small" ],
8     "ConstraintDescription": "a valid EC2 instance type."
9   } //...
10 },
11 "Resources": {
12   "WebServer": {
13     "Type": "AWS::EC2::Instance",
14     "Properties": {
15       "InstanceType": { "Ref" : "InstanceType" } ,
16       "UserData": { "Fn::Base64" : { "Fn::Join" : ["", [
17         "#!/bin/bash -xe\n",
18         "yum update -y aws-cfn-bootstrap\n",
19
20         "/opt/aws/bin/cfn-init -v ",
21         "  --stack ", { "Ref" : "AWS::StackName" },
22         "  --resource WebServer ",
23         "  --configsets wordpress_install ",
24         "  --region ", { "Ref" : "AWS::Region" }, "\n"
25       ]}}, //...
26     }},
27   }
28 },
29 "Outputs": {
30   "WebsiteURL": {
31     "Value":
32     { "Fn::Join" :
33       [ "", [ "http://", { "Fn::GetAtt" :
34         [ "WebServer", "PublicDnsName" ] }, "/wordpress" ] ] },
35     "Description": "WordPress Website"
36   }
37 }
38 }

```

Figure 1. A deployment program in CloudFormation (details omitted, full version: https://s3.eu-central-1.amazonaws.com/cloudformation-templates-eu-central-1/WordPress_Multi_AZ.template).

The source code of the PLT Redex and Scala implementations and of all case studies is available online: <https://github.com/seba--/djc-lang>.

2. Motivation

In this section, we analyze the issues that programmers encounter with current configuration and deployment languages on cloud platforms by a concrete example.

2.1 State of the Art

Figure 1 shows an excerpt of a deployment program in CloudFormation, a JSON-based DSL for Amazon AWS. The example is from the CloudFormation documentation. We summarize the main characteristics of the deployment language below.

- Input parameters capture varying details of a configuration (Lines 2-10). For example, the program receives the virtual machine instance type that should host the web server for a user blog (Line 3). This enables reuse of the program with different parameters.
- CloudFormation programs specify *named resources* to be created in the deployment (Lines 11-28), e.g., deployed virtual machines, database instances, load balancers and

even other programs as modules. The program in Figure 1 allocates a "WebServer" resource (Line 12), which is a virtual machine instance. The type of the virtual machine references a parameter (Line 15), that the program declared earlier on (Line 3). Resources can refer to each other, for example, configuration parameters of a web server may refer to tables in a database resource.

- Certain configuration phases require to execute application code inside virtual machine instances after the deployment stage. Application code is often directly specified in resource bodies (Lines 17-24). In the example, a bash script defines the list of software packages to install on the new machine instance (in our case a WordPress⁴ blog). In principle, arbitrary programs in any language can be specified.
- Deployment programs specify output parameters (Lines 29-37), which may depend on input parameters and resources. Output parameters are returned to the caller after executing the deployment program. In this example, it is a URL pointing to the new WordPress blog.
- The deployment program is interpreted at run time by the cloud platform which performs the deployment steps according to the specification.

2.2 Problems with Deployment Programs

In the following we discuss the issues with the CloudFormation example described above.

Internal Safety Type safety for deployment programs is limited. Developers define "types" for resources of the cloud platform, e.g., `AWS::EC2::Instance` (Line 13) represents an Amazon EC2 instance. However, the typing system of current cloud deployment languages is primitive and only relies on the JSON types.

Cross-language Safety Even more problematic are issues caused by cross-language interaction between the deployment language and the language(s) of the deployed application services. For example, the `AWS::Region` variable is passed from the JSON specification to the bash script (Line 24). However, the sharing mechanism is just syntactic replacement of the current value of `AWS::Region` inside the script. Neither are there type-safety checks nor syntactic checks before the script is executed. More generally, there is no guarantee that the data types of the deployment language are compatible with the types of the application language nor that the resulting script is syntactically correct. This problem makes cloud applications susceptible to hygiene-related bugs and injection attacks [4].

Low Abstraction Level Deployment languages typically are Turing-complete but the abstractions are low-level and not deployment-specific. For example, (1) deployment

programs receive parameters and return values similar to procedures and (2) deployment programs can be instantiated from inside other deployment programs, which resembles modules. Since deployment is a complex engineering task, advanced language features are desirable to facilitate programming in the large, e.g., higher-order combinators, rich data types and strong interfaces.

Two-phase Staging Deployment programs in current DSLs execute before the actual application services, that is, information flows from the deployment language to the deployed application services, but not the other way around. As a result, an application service cannot reconfigure a deployment based on the run time state. Recent scenarios in reactive and big data computations demonstrate that this is a desirable feature [10].

Lack of Extensibility Resources and service references in deployment programs refer to pre-defined abstractions of the cloud platform, which have rigid interfaces. Cloud platforms determine the semantics of the services. Programmers cannot implement their own variants of services that plug into the deployment language with the same interfaces as the native services.

Informal Specification The behavior of JSON deployment scripts is only informally defined. The issue is exacerbated by the mix of different languages. As a result, it is hard to reason about properties of systems implemented using deployment programs.

The issues above demand for a radical change in the way programmers deploy cloud applications and in the way application and deployment configuration code relate to each other.

3. The Cloud Platform Language

A solution to the problems identified in the previous section requires an holistic approach where cloud abstractions are explicitly represented in the language. Programmers should be able to modularly specify application behavior as well as reconfiguration procedures. Run time failures should be prevented at compilation time through type checking.

These requirements motivated the design of CPL. In this section, we present its syntax and the operational semantics.

3.1 Language Features in a Nutshell

Simple Meta-Theory: CPL should serve as the basis for investigating high-level language features and type systems designed for cloud computations. To this end, it is designed as a core language with a simple meta-theory. Established language features and modeling techniques, such as lexical scoping and a small-step operational semantics, form the basis of CPL.

Concurrency: CPL targets distributed concurrent computations. To this end, it allows the definition of independent computation units, which we call *servers*.

⁴<http://wordpress.org>

Asynchronous Communication: Servers can receive parameterized *service requests* from other servers. To realistically model low-level communication within a cloud, the language only provides asynchronous end-to-end communication, where the success of a service request is not guaranteed. Other forms of communication, such as synchronous, multicast, or error-checking communication, can be defined on top of the asynchronous communication.

Local Synchronization: Many useful concurrent and asynchronous applications require synchronization. We adopt *join patterns* from the Join Calculus [12]. Join patterns are *declarative* synchronization primitives for machine-local synchronization.

First-Class Server Images: Cloud platforms employ virtualization to spawn, suspend and duplicate virtual machines. That is, virtual machines are data that can be stored and send as payload in messages. This is the core idea behind CPL’s design and enables programs to change their deployment at run time. Thus, CPL features servers as values, called *first-class servers*. Active *server instances* consist of an address, which points to a *server image* (or *snapshot*). The server image embodies the current run time state of a server and a description of the server’s functionality, which we call *server template*. At run time, a server instance may be overwritten by a new server image, thus changing the behavior for subsequent service requests to that instance.

Transparent Placement: Cloud platforms can reify new machines physically (on a new network node) or virtually (on an existing network node). Since this difference does not influence the semantics of a program but only its non-functional properties (such as performance), our semantics is transparent with respect to placement of servers. Thus, actual languages based on our core language can freely employ user-defined placement definitions and automatic placement strategies. Also, we do not require that CPL-based languages map servers to virtual machines, which may be inefficient for short-lived servers. Servers may as well represent local computations executing on a virtual machine.

3.2 Core Syntax

Figure 2 displays the core syntax of CPL. An expression e is either a value or one of the following syntactic forms:⁵

- A *variable* x is from the countable set \mathcal{N} . Variables identify services and parameters of their requests.
- A *server template* $(\text{srv } \bar{r})$ is a first-class value that describes the behavior of a server as a sequence of reaction rules \bar{r} . A reaction rule takes the form $\bar{p} \triangleright e$, where \bar{p} is a

sequence of joined service patterns and e is the body. A *service pattern* $x_0 \langle \bar{x} \rangle$ in \bar{p} declares a service named x_0 with parameters \bar{x} and a rule can only fire if all service patterns are matched simultaneously. The same service pattern can occur in multiple rules of a server.

- A *server spawn* $(\text{spwn } e)$ creates a new running *server instance* at a freshly allocated *server address* i from a given *server image* $(\text{srv } \bar{r}, \bar{m})$ represented by e . A *server image* is a description of a server behavior plus a server state – a *buffer* of pending messages. A real-world equivalent of server images are e.g., virtual machine snapshots. A special case of a server image is the value $\mathbf{0}$, which describes an inactive or shut down server.
- A fully qualified *service reference* $e \# x$, where e denotes a server address and x is the name of a service provided by the server instance at e . Service references to server instances are themselves values.
- A *self-reference* **this** refers to the address of the lexically enclosing server template, which, e.g., allows one service to call upon other services of the same server instance.
- An asynchronous *service request* $e_0 \langle \bar{e} \rangle$, where e_0 represents a service reference and \bar{e} the arguments of the requested service.
- A *parallel expression* $(\text{par } \bar{e})$ of service requests \bar{e} to be executed independently. The empty parallel expression $(\text{par } \varepsilon)$ acts as a noop expression, unit value, or null process and is a value.
- A *snapshot* **snap** e yields an image of the server instance which resides at the address denoted by e .
- A *replacement* **repl** $e_1 e_2$ of the server instance at address e_1 with the server image e_2 .

Notation: In examples, p & p denotes pairs of join patterns and $e \parallel e$ denotes pairs of parallel expressions. We sometimes omit empty buffers when spawning servers, i.e., we write **spwn** $(\text{srv } \bar{r})$ for **spwn** $(\text{srv } \bar{r}, \varepsilon)$. To improve readability in larger examples, we use curly braces to indicate the lexical scope of syntactic forms. We write service names and meta-level definitions in typewriter font, e.g., **this**#foo and **MyServer** = **srv** { }. We write bound variables in italic font, e.g., **srv** { *left* $\langle x \rangle$ & *right* $\langle y \rangle$ \triangleright *pair* $\langle x, y \rangle$ }.

Example. For illustration, consider the following server template **Fact** for computing factorials, which defines three rules with 5 services.⁶

```

1 Fact = srv {
2   main $\langle n, k \rangle \triangleright //\text{initialization}$ 
3     this#fac $\langle n \rangle \parallel \text{this}\#\text{acc}\langle 1 \rangle \parallel \text{this}\#\text{out}\langle k \rangle$ 
4
5   fac $\langle n \rangle$  & acc $\langle a \rangle \triangleright //\text{recursive fac computation}$ 

```

⁵ We write \bar{a} to denote the finite sequence $a_1 \dots a_n$ and we write ε to denote the empty sequence.

⁶ For the sake of presentation, we use ordinary notation for numbers, arithmetics and conditionals, all of which is church-encodable on top of CPL (cf. Section 3.5).

$e ::= v \mid x \mid \mathbf{this} \mid \mathbf{srv} \bar{r} \mid \mathbf{spwn} e \mid e \# x \mid e(\bar{e}) \mid \mathbf{par} \bar{e} \mid \mathbf{snap} e \mid \mathbf{repl} e e$	(Expressions)	$i \in \mathbb{N}$	(Server Addresses)
$v ::= \mathbf{srv} \bar{r} \mid i \mid i \# x \mid \mathbf{par} \varepsilon \mid (\mathbf{srv} \bar{r}, \bar{m}) \mid \mathbf{0}$	(Values)	$r ::= \bar{p} \triangleright e$	(Reaction Rules)
$E ::= [\] \mid \mathbf{spwn} E \mid E \# x \mid E(\bar{e}) \mid e(\bar{e} E \bar{e}) \mid \mathbf{par} \bar{e} E \bar{e} \mid \mathbf{snap} E \mid \mathbf{repl} E e \mid \mathbf{repl} e E$	(Evaluation Contexts)	$p ::= x(\bar{x})$	(Join Patterns)
$x, y, z \in \mathcal{N}$	(Variable Names)	$m ::= x(\bar{v})$	(Message Values)
		$\mu ::= \emptyset \mid \mu; i \mapsto (\mathbf{srv} \bar{r}, \bar{m}) \mid \mu; i \mapsto \mathbf{0}$	(Routing Tables)

Figure 2. Expression Syntax of CPL.

$\frac{e \mid \mu \longrightarrow e' \mid \mu'}{E[e] \mid \mu \longrightarrow E[e'] \mid \mu'}$	(CONG)	$\frac{\mu(i) = s \quad (s = \mathbf{0} \vee s = (\mathbf{srv} \bar{r}, \bar{m}))}{\mathbf{snap} i \mid \mu \longrightarrow s \mid \mu}$	(SNAP)
$\overline{\mathbf{par} \bar{e}_1 (\mathbf{par} \bar{e}_2) \bar{e}_3 \mid \mu \longrightarrow \mathbf{par} \bar{e}_1 \bar{e}_2 \bar{e}_3 \mid \mu}$	(PAR)	$\frac{i \in \text{dom}(\mu) \quad (s = \mathbf{0} \vee s = (\mathbf{srv} \bar{r}, \bar{m}))}{\mathbf{repl} i s \mid \mu \longrightarrow \mathbf{par} \varepsilon \mid \mu; i \mapsto s}$	(REPL)
$\frac{\mu(i) = (s, \bar{m})}{i \# x(\bar{v}) \mid \mu \longrightarrow \mathbf{par} \varepsilon \mid \mu; i \mapsto (s, \bar{m} \cdot x(\bar{v}))}$	(RCV)	Matching Rules:	
$\frac{\mu(i) = (s, \bar{m}) \quad s = \mathbf{srv} \bar{r}_1 (\bar{p} \triangleright e_b) \bar{r}_2 \quad \text{match}(\bar{p}, \bar{m}) \Downarrow (\bar{m}', \sigma) \quad \sigma_b = \sigma \cup \{\mathbf{this} := i\}}{\mathbf{par} e \mid \mu \longrightarrow \mathbf{par} e \sigma_b(e_b) \mid \mu; i \mapsto (s, \bar{m}')}$	(REACT)	$\overline{\text{match}(\varepsilon, \bar{m}) \Downarrow (\bar{m}, \emptyset)}$	(MATCH ₀)
$\frac{i \notin \text{dom}(\mu) \quad (s = \mathbf{0} \vee s = (\mathbf{srv} \bar{r}, \bar{m}))}{\mathbf{spwn} s \mid \mu \longrightarrow i \mid \mu; i \mapsto s}$	(SPWN)	$\frac{\bar{m} = \bar{m}_1 (x(v_1 \dots v_k)) \bar{m}_2 \quad \sigma = \{x_i := v_i \mid 1 \leq i \leq k\} \quad \text{match}(\bar{p}, \bar{m}_1 \bar{m}_2) \Downarrow (\bar{m}_r, \sigma_r) \quad \text{dom}(\sigma) \cap \text{dom}(\sigma_r) = \emptyset}{\text{match}(x(x_1 \dots x_k) \bar{p}, \bar{m}) \Downarrow (\bar{m}_r, \sigma \cup \sigma_r)}$	(MATCH ₁)

Figure 3. Small-step Operational Semantics of CPL.

$e ::= \dots \mid \Lambda \alpha <: T. e \mid e [T]$	(Extended Expressions)	$T ::= \mathbf{Top} \mid \mathbf{Unit} \mid \alpha \mid \langle \bar{T} \rangle \mid \mathbf{srv} \overline{x: \bar{T}} \mid \mathbf{srv} \perp \mid \mathbf{inst} T$	(Types)
$v ::= \dots \mid \Lambda \alpha <: T. e$	(Extended Values)	$\mid \mathbf{img} T \mid \forall \alpha <: T. T$	
$p ::= x(x: \bar{T})$	(Typed Join Patterns)	$\Gamma ::= \emptyset \mid \Gamma, \alpha <: T \mid \Gamma, x: T \mid \Gamma, \mathbf{this}: T$	(Type Contexts)
$\alpha, \beta, \gamma \dots$	(Type-Variables)	$\Delta ::= \emptyset \mid \Delta, i: T$	(Location Typings)

Figure 4. Expression Syntax of CPL with Types

```

6   if (n ≤ 1)
7   then this#res⟨a⟩
8   else (this#fac⟨n - 1⟩ || this#acc⟨a * n⟩)
9
10  res⟨n⟩ & out⟨k⟩ ▷ k⟨n⟩ //send result to k
11 }

```

The first rule defines a service `main` with two arguments, an integer n and a continuation k . The continuation is necessary because service requests are asynchronous and thus, the factorial server must notify the caller when the computation finishes. Upon receiving a `main` request, the server sends itself three requests: `fac` represents the outstanding factorial computation, `acc` is used as an accumulator for the ongoing computation, and `out` stores the continuation provided by the caller.

The second rule of `Fact` implements the factorial function and synchronously matches and consumes requests `fac` and `acc` using join patterns. Upon termination, the second rule sends a request `res` to the running server instance, otherwise it decreases the argument of `fac` and updates the accumulator. Finally, the third rule of `Fact` retrieves the user-provided

continuation k from the request out and the result `res`. The rule expects the continuation to be a service reference and sends a request to it with the final result as argument.

To compute a factorial, we create a server instance from the template `Fact` and request service `main`:

```
(spwn Fact)#main⟨5, k⟩.
```

An example reduction trace is in the appendix.

3.3 Operational Semantics

We define the semantics of CPL as a small-step structural operational semantics using reduction contexts E (Figure 2) in the style of Felleisen and Hieb [8].

Figure 3 shows the reduction rules for CPL expressions. Reduction steps are atomic and take the form $e \mid \mu \longrightarrow e' \mid \mu'$. A pair $e \mid \mu$ represents a distributed cloud application, where expression e describes its current behavior and μ describes its current *distributed* state. We intend e as a description of the software components and resources that execute and reside at the cloud provider and do not model client devices. We call the component μ a *routing table*, which is a finite map. Intuitively, μ records which addresses a cloud provider

$\frac{\Gamma(x) = T \quad x \in \mathcal{N} \cup \{\mathbf{this}\}}{\Gamma \mid \Sigma \vdash x : T}$	(T-VAR)	$\frac{\forall i. \Gamma \mid \Sigma \vdash e_i : \mathbf{Unit}}{\Gamma \mid \Sigma \vdash \mathbf{par} \ \bar{e} : \mathbf{Unit}}$	(T-PAR)
$\begin{array}{l} r_i = \bar{p}_i \triangleright e_i \quad p_{i,j} = x_{i,j} \langle \overline{y_{i,j} : T_{i,j}} \rangle \quad S_{i,j} = \langle \overline{T_{i,j}} \rangle \\ T = \mathbf{srv} \ x_{i,j} : S_{i,j} \quad (\forall i, j, k. j \neq k \rightarrow \overline{y_{i,j}} \cap \overline{y_{i,k}} = \emptyset) \\ (\forall i, j, k, l. x_{i,j} = x_{k,l} \rightarrow T_{i,j} = T_{k,l}) \quad \text{ftv}(T) \subseteq \text{ftv}(\Gamma) \\ \forall i. \Gamma, \overline{y_{i,j} : T_{i,j}}, \mathbf{this} : T \mid \Sigma \vdash e_i : \mathbf{Unit} \end{array}$			(T-SRV)
$\frac{}{\Gamma \mid \Sigma \vdash \mathbf{0} : \mathbf{img} \ \mathbf{srv} \ \perp}$			(T-0)
$\frac{\Gamma \mid \Sigma \vdash \mathbf{srv} \ \bar{r} : T \quad r_i = \bar{p}_i \triangleright e_i \quad p_{i,j} = x_{i,j} \langle \overline{y_{i,j} : T_{i,j}} \rangle \quad (\forall k. \exists i, j. (m_k = x_{i,j} \langle \overline{v_{i,j}} \rangle \wedge \Gamma \mid \Sigma \vdash v_{i,j} : T_{i,j}))}{\Gamma \mid \Sigma \vdash (\mathbf{srv} \ \bar{r}, \bar{m}) : \mathbf{img} \ T}$			(T-IMG)
$\frac{\Gamma \mid \Sigma \vdash e : \mathbf{inst} \ T}{\Gamma \mid \Sigma \vdash \mathbf{snap} \ e : \mathbf{img} \ T}$	(T-SNAP)	$\frac{\Gamma \mid \Sigma \vdash e : \langle T_1 \dots T_n \rangle \quad \forall i. \Gamma \mid \Sigma \vdash e_i : T_i}{\Gamma \mid \Sigma \vdash e \langle e_1 \dots e_n \rangle : \mathbf{Unit}}$	(T-REQ)
$\frac{\Gamma \mid \Sigma \vdash e_1 : \mathbf{inst} \ T \quad \Gamma \mid \Sigma \vdash e_2 : \mathbf{img} \ T}{\Gamma \mid \Sigma \vdash \mathbf{repl} \ e_1 \ e_2 : \mathbf{Unit}}$	(T-REPL)	$\frac{\Gamma, \alpha <: T \mid \Sigma \vdash e : U}{\Gamma \mid \Sigma \vdash \Lambda \alpha <: T. e : \forall \alpha <: T. U}$	(T-TABS)
$\frac{\Gamma \mid \Sigma \vdash e : \mathbf{img} \ T}{\Gamma \mid \Sigma \vdash \mathbf{spwn} \ e : \mathbf{inst} \ T}$	(T-SPWN)	$\frac{\Gamma \mid \Sigma \vdash e : \forall \alpha <: T_2. T \quad \Gamma \vdash T_1 <: T_2 \quad \text{ftv}(T_1) \subseteq \text{ftv}(\Gamma)}{\Gamma \mid \Sigma \vdash e [T_1] : T\{\alpha := T_1\}}$	(T-TAPP)
$\frac{\Sigma(i) = \mathbf{img} \ T}{\Gamma \mid \Sigma \vdash i : \mathbf{inst} \ T}$	(T-INST)	$\frac{\Gamma \mid \Sigma \vdash e : T \quad \Gamma \vdash T <: U}{\Gamma \vdash e : U}$	(T-SUB)
$\frac{\Gamma \mid \Sigma \vdash e : \mathbf{inst} \ (\mathbf{srv} \ \overline{x : T})}{\Gamma \mid \Sigma \vdash e \# x_i : T_i}$	(T-SVC)		

Figure 5. Typing rules of CPL.

$\frac{}{\Gamma \vdash T <: \mathbf{Top}}$	(S-TOP)	$\frac{\Gamma \vdash T <: U}{\Gamma \vdash \mathbf{img} \ T <: \mathbf{img} \ U}$	(S-IMG)
$\frac{\alpha <: T \in \Gamma}{\Gamma \vdash \alpha <: T}$	(S-TVAR)	$\frac{\forall i. \Gamma \vdash U_i <: T_i}{\Gamma \vdash \langle T_1, \dots, T_n \rangle <: \langle U_1, \dots, U_n \rangle}$	(S-SVC)
$\frac{\forall j. \exists i. (y_j = x_i \wedge \Gamma \vdash T_i <: U_j)}{\Gamma \vdash \mathbf{srv} \ \overline{x : T} <: \mathbf{srv} \ \overline{y : U}}$	(S-SRV)	$\frac{\Gamma, \alpha_1 <: T \vdash U_1 <: U_2\{\alpha_2 := \alpha_1\}}{\Gamma \vdash (\forall \alpha_1 <: T. U_1) <: (\forall \alpha_2 <: T. U_2)}$	(S-UNIV)
$\frac{\Gamma \vdash T <: U}{\Gamma \vdash \mathbf{inst} \ T <: \mathbf{inst} \ U}$	(S-INST)	$\frac{}{\Gamma \vdash T <: T}$	(S-REFL)
$\frac{}{\Gamma \vdash \mathbf{srv} \ \perp <: \mathbf{srv} \ T}$	(S-SRV \perp)	$\frac{\Gamma \vdash T_1 <: T_2 \quad \Gamma \vdash T_2 <: T_3}{\Gamma \vdash T_1 <: T_3}$	(S-TRANS)

Figure 6. Subtyping rules of CPL.

assigns to the server instances that the cloud application creates during its execution.⁷ We abstract over technical details, such as the underlying network.

The first reduction rule (CONG) defines the congruence rules of the language and is standard. The second rule (PAR) is technical. It flattens nested parallel expressions in order to have a simpler representation of parallel computations. The third rule (RCV) lets a server instance receive an asynchronous service request, where the request is added to the

instance's buffer for later processing. Our semantics abstracts over the technicalities of network communication. That is, we consider requests $i \# x \langle \bar{v} \rangle$ that occur in a CPL expression to be in transit, until a corresponding (RCV) step consumes them. The fourth rule (REACT) fires reaction rules of a server. It selects a running server instance (s, \bar{m}) , selects a reaction rule $(\bar{p} \triangleright e_b)$ from it, and tries to match its join patterns \bar{p} against the pending service requests in the buffer \bar{m} . A successful match consumes the service requests, instantiates the body e_b of the selected reaction rule and executes it independently in parallel.

⁷This bears similarity to lambda calculi enriched with references and a store [31].

Finally, let us consider the rules for **spwn**, **snap** and **repl**, which manage server instances and images. Reduction rule (SPWN) creates a new server instance from a server image, where a fresh unique address is assigned to the server instance. This is the only rule that allocates new addresses in μ . One can think of this rule as a request to the cloud provider to create a new virtual machine and return its IP address. Importantly, the address that **spwn** yields is only visible to the caller. The address can only be accessed by another expression if it shares a common lexical scope with the caller. Thus, lexical scope restricts the visibility of addresses. This also means that the map μ is not a shared memory, but a combined, flat view of disjoint distributed information.⁸

Reduction rule (SNAP) yields a copy of the server image at address i , provided the address is in use. Intuitively, it represents the invocation of a cloud management API to create a virtual machine snapshot. Reduction rule (REPL) replaces the server image at address i with another server image.

We define **spwn**, **snap** and **repl** as atomic operations. At the implementation level, each operation may involve multiple communication steps with the cloud provider, taking noticeable time to complete and thus block execution for too long, especially when the operation translates to booting a new OS-level virtual machine. On the other hand, as we motivated at the beginning of this section, servers may not necessarily map to virtual machines, but in-memory computations. In this case, we expect our three atomic operations to be reasonably fast. Also, we do not impose any synchronization mechanism on a server addresses, which may result in data races if multiple management operations access it in parallel. Instead, programmers have to write their own synchronization mechanisms on top of CPL if required. Matching satisfies the following property:

Proposition 1 (Match soundness and completeness). *Let \bar{p} be a sequence of join patterns with $p_i = x_i \langle \bar{y}_i \rangle$, \bar{m} and \bar{m}_r sequences of service request values, and σ a substitution. $\text{match}(\bar{p}, \bar{m}) \Downarrow (\bar{m}_r, \sigma)$ if and only if it exists a sequence \bar{m}_c such that:*

1. Sequence \bar{m}_c represents the requests values consumed from \bar{m} , that is, $\bar{m}_r \bar{m}_c = \bar{m}$ modulo permutation.
2. All consumed requests \bar{m}_c match the join patterns \bar{p} , that is, \bar{m}_c and \bar{p} have the same length and $m_{c,i} = x_i \langle \bar{v}_i \rangle$, where \bar{y}_i and \bar{v}_i have the same length.
3. σ substitutes the parameters of the matched join patterns with the actual arguments, that is,

$$\sigma = \{ \bar{y}_i := \bar{v}_i \mid 1 \leq i \leq k \}$$
 where k is the length of \bar{p} .

Proof. Soundness (\Rightarrow): Straightforward induction on the derivation of the judgment $\text{match}(\bar{p}, \bar{m}_1) \Downarrow (\bar{m}_2, \sigma)$. Com-

⁸This approach is comparable to sets of definitions in the chemical soup of the Join Calculus [12].

pleteness (\Leftarrow): Straightforward by induction on the number k of service patterns in p . \square

Our semantics is nondeterministic along 3 dimensions:

- If multiple server instances can fire a rule, (REACT) selects one of them nondeterministically. This models concurrent execution of servers that can react to incoming service requests independently.
- If multiple rules of a server instance can fire, (REACT) selects one of them nondeterministically. This is of lesser importance and languages building on ours may fix a specific order for firing rules (e.g., in the order of definition).
- If multiple service request values can satisfy a join pattern, (MATCH₁) selects one of them nondeterministically. This models asynchronous communication in distributed systems, i.e., the order in which a server serves requests is independent of the order in which services are requested. More concrete languages based on CPL may employ stricter ordering (e.g., to preserve the order of requests that originate from a single server).

3.4 Placement of Servers.

We intentionally designed the semantics of CPL with transparency of server *placement* in mind. That is, a single abstraction in the language, the server instance, models all computations, irrespective of whether the instance runs on its own physical machine or as a virtual machine hosted remotely – indeed, placement transparency is a distinguishing feature of cloud applications.

However, despite the behavior of servers being invariant to placement, placement has a significant impact in real-world scenarios and influences communication and computation performance [2, 19]. The need to account for placement in an implementation is critical considering that – servers being the only supported abstraction – every single let binding and lambda abstraction desugars to a server spawn (cf. Section 3.5). In our concurrent Scala implementation, we support an extended syntax for server spawns that allows programmers to declare whether a server instance runs in a new thread or in the thread that executes the spawn. This provides a simple mechanism for manually implementing placement strategies.

A viable alternative to manual specification of placement are automatic placement strategies. Together with server migration, automatic placement strategies can adapt the server layout to changing conditions. Based on our language, a management system for a cloud infrastructure can formally reason about optimal placement strategies. In future work, we plan to implement these ideas in a distributed run-time system for CPL (cf. Section 5.4).

3.5 Derived syntax and base operations

Our core language is expressive enough to encode a wide range of typical language constructs. To illustrate its expres-

siveness and for convenience in expressing example computations in the rest of the paper, we define derived syntax for let-expressions, first-class functions, thunks, and base operations, all of which can be desugared to the core syntax introduced above.

Let bindings. The derived syntax for **let** bindings desugars to the core syntax of the CPL as follows:

$\text{let } x = e_1 \text{ in } e_2 \rightsquigarrow (\text{spwn } (\text{srv let } \langle x \rangle \triangleright e_2)) \# \text{let } \langle e_1 \rangle$. Evaluating **let** amounts to (a) spawning a new server instance that offers a service called **let** that will run e_2 when requested and (b) requesting this service with the bound expression e_1 as an argument.

We also define derived syntax for a variant of **let** called **letk** for cases in which the bound expression provides its result through a continuation. This is to account for the fact that often expressions in the CPL involve asynchronous service calls that, instead of returning a value, pass it to a continuation. The definition of **letk** is as follows:

$\text{letk } x = e_1 \langle \bar{e} \rangle \text{ in } e_2 \rightsquigarrow e_1 \langle \bar{e} \rangle, (\text{spwn } (\text{srv k } \langle x \rangle \triangleright e_2)) \# \text{k}$. Here, we bind the variable x via continuation that we add to the service request $e_1 \langle \bar{e} \rangle$, assuming e_1 takes a continuation as final argument. When e_1 terminates, it calls the continuation and thus triggers execution of e_2 . For example, we can use **letk** to bind and use the result of the **Fact** server shown above:

$\text{letk } n = (\text{spwn Fact}) \# \text{main} \langle 5 \rangle \text{ in Log} \# \text{write} \langle n \rangle$

Note that the desugaring for both variants of **let** wrap the body e_2 in a server template, which changes the meaning of the self reference **this** in e_2 . To counter this effect and to make the derived syntax transparent, the desugaring that we actually implemented substitutes free occurrences of **this** in e_2 to the server instance surrounding the **let**.

First-class functions. We can encode first-class functions as server instances with a single service **app**:

$\lambda x. e \rightsquigarrow \text{spwn } (\text{srv app } \langle x, k \rangle \triangleright T(e, k))$, where k is fresh. Recall that service requests in CPL are asynchronous. In order to correctly propagate argument values and the result of function bodies, we need to transform argument expressions and function bodies into continuation-passing style, for example using the following transformation T :

$$\begin{aligned} T(\lambda x. e, k) &= k(\text{spwn } (\text{srv app } \langle x, k \rangle \triangleright T(e, k))) \\ T((f e), k) &= T(f, (\text{spwn } (\text{srv k}_1 \langle v_f \rangle \triangleright \text{where } v_f \text{ is fresh} \\ &\quad T(e, (\text{spwn } (\text{srv k}_2 \langle v_e \rangle \triangleright \text{where } v_e \text{ is fresh} \\ &\quad v_f \# \text{app}(v_e, k)))) \# \text{k}_1)) \\ T(e, k) &= k(e) \end{aligned}$$

For example, we can define and apply a function that instantiates a server-template argument:

$$(\lambda x. \text{spwn } x) \# \text{app} \langle \text{Fact}, k_0 \rangle \longrightarrow^* k_0 \langle \text{Fact}^0 \rangle$$

Our encoding of first-class functions is similar to the one proposed for the Join Calculus [12] and it also shows that our language is Turing-complete. Moreover, it enables Church-encodings of data types such as numbers or lists.

Thunks. A thunk is a first-class value that represents a packaged, delayed computation. Servers can force the com-

putation of a thunk and they can pass thunks to other servers. Thunks play a significant role in distributed systems, because they enable servers to distribute work over other servers dynamically.

Interestingly, lambdas as defined above do not give rise to a useful implementation of thunks, because a computation that is encoded as a lambda is already installed on a concrete spawned server: Every lambda expression gives rise to exactly one server instance that solely executes the body of this lambda. In contrast, we desire an implementation of thunks that allows us to dynamically allocate servers for executing a thunk. To this end, we represent thunks as server templates:

$\text{thunk } e \rightsquigarrow \text{srv force} \langle k \rangle \triangleright k \langle e \rangle$

Since server templates are first-class in CPL, thunks can be passed to other servers. A server can instantiate a thunk any number of times and call the **force** request with a continuation to get the result of the thunk.

Note that similarly to **let**, we substitute **this** in thunks and lambda abstractions by the enclosing server instance to make our encodings transparent.

Base operations. While we can use Church encodings to represent data types and their operations, it is more convenient (and more efficient in practice) to assume some built-in base operations. In particular, we can take the liberty of assuming that base operations are synchronous and in direct style, that is, base operations return values directly and do not require continuation-passing style. For the remainder of the paper, we assume built-in base operations on Booleans, integers, floating points, tuples and lists. We added these operations in our implementation and it is easy to add further base operations. To distinguish synchronous calls to base operations from asynchronous service requests, we use rounded parentheses for base operations, for example, $\text{max}(7, 11)$.

4. Type System

We designed and formalized a type system for CPL in the style of System F with subtyping and bounded quantification [24]. The type system ensures that all service requests in a well-typed program refer to valid service declarations with the correct number of arguments and the right argument types. Subtyping enables us to define public server interfaces, where the actual server implementation defines private services, for example, to manage internal state.

Figure 4 shows the syntax of types, typing contexts, location typings as well as extensions to expressions and values. Similar to lambda calculi with references, alongside standard typing contexts Γ we also record the type of server instances at each allocated address via location typings Σ . A type T is either the top type **Top**, the unit type **Unit**, a type variable α , a service type $\langle \bar{T} \rangle$ representing a service with arguments of type T_i , a server-template type $\text{srv } x : \bar{T}$ of a server with services x_i of type T_i , the special server-template type $\text{srv } \perp$ for inactive servers, a server-instance type $\text{inst } T$, a server-image type $\text{img } T$ or a universal

type $\forall \alpha <: T_1. T_2$. The syntax of typing contexts and the extensions of expressions and values is standard: We have type abstraction $\Lambda \alpha <: T. e$, and type application $e [T]$. Finally, we require type annotations in join patterns.

We define the typing judgment $\Gamma \mid \Sigma \vdash e : T$ by the rules depicted in figure 5. The rules are mostly straightforward. (T-VAR) looks up the type of a variable or **this** in the context. (T-PAR) requires that all expressions of a parallel expression have type Unit.

(T-SRV) is the most complicated type rule. Intuitively, the type of a server template is the set of all services that the server offers. r_i represents rule number i of the server template, where $p_{i,j}$ is pattern number j of rule number i . Patterns $p_{i,j}$ provide services $x_{i,j}$, which have service type $S_{i,j}$. The type T of the server template then consists of all provided services with their types. To make sure the server template is well-typed, we check that join patterns are linear (service parameters are distinct), services in different patterns have consistent types and that all free type variables are bound ($\text{ftv}(T) \subseteq \text{ftv}(\Gamma)$). Finally, we check the right-hand side e_i of each reaction rule, where we bind all service parameters $y_{i,j}$ as well as **this**.

Next, we define three introduction rules for server image types. The first is (T-O), which specifies that **0** is an image of an inert server. The second rule (T-IMG) types server image values $(\text{srv } \bar{r}, \bar{m})$, where we require that $\text{srv } \bar{r}$ is a well-typed server template and each service request value in the buffer \bar{m} is understood by this server template. That is, each value m_k in \bar{m} must correspond to a join pattern mentioned in \bar{r} and the arguments must have the types which are annotated in the join pattern. The last introduction rule for server image types is (T-SNAP) for snapshots, which requires that the argument to **snap** is actually a server instance in order to yield a corresponding server image.

Rule (T-REPL) types replacements $\text{repl } e_1 e_2$ as Unit and requires that replacements are preserving the interface of the server instance to be replaced. That is, the first argument must be an instance with interface type T and the second argument an image type for the same interface type.

There are two introduction rules for server instances. (T-SPWN) requires the argument of **spwn** to be a server image in order to yield a corresponding instance. Rule (T-INST) handles server addresses, which must be allocated in the location typing Σ to a server image.

(T-SVC) defines that a service reference is well-typed if the queried server provides a service of the required name. (T-REQ) requires that the target of a service request indeed is a service reference and that the request has the right number of arguments with the right types. The remaining four type rules are standard.

Figure 6 defines the subtyping relation $\Gamma \vdash T <: T$. We employ width subtyping and depth subtyping for server-template types, that is, the server subtype can provide more services than the server supertype and the server subtype

can provide specialized versions of services promised by the server supertype. A special case is rule (S-SRV $_{\perp}$), which specifies that the type $\text{srv } \perp$ for inert servers is a subtype of every other server template type. This ensures that **0** can be placed in every context requiring an image of type $\text{img } \text{srv } T$. The other subtyping rules are straightforward.

Preservation. We prove preservation for our type system using standard substitution lemmas [24]. The proofs appear in the appendix at the end of the paper.

Lemma 1 (Substitution Lemma). *If $\Gamma, x : T_1 \mid \Sigma \vdash e_2 : T_2$ and $\Gamma \mid \Sigma \vdash e_1 : T_1$ then $\Gamma \mid \Sigma \vdash e_2\{x := e_1\} : T_2$.*

Lemma 2 (Type Substitution Preserves Subtyping). *If $\Gamma, \alpha <: T', \Gamma' \vdash S <: T$ and $\Gamma \vdash S' <: T'$ then $\Gamma, \Gamma' \sigma \vdash S \sigma <: T \sigma$ where $\sigma = \{\alpha := S'\}$.*

Lemma 3 (Type Substitution Lemma). *If $\Gamma, \alpha <: S, \Gamma' \mid \Sigma \vdash e : T$ and $\Gamma \vdash S' <: S$ then $\Gamma, \Gamma' \sigma \mid \Sigma \sigma \vdash e \sigma : T \sigma$ where $\sigma = \{\alpha := S'\}$.*

Lemma 4 (Location Typing Extension Preserves Types). *If $\Gamma \mid \Sigma \vdash e : T$ and $\Sigma \subseteq \Sigma'$, then $\Gamma \mid \Sigma' \vdash e : T$.*

Lemma 5 (Replacement). *If \mathcal{D} is a derivation with root $\Gamma \vdash E[e] : U$, $\mathcal{D}' \preceq \mathcal{D}$ a derivation with root $\Gamma' \vdash e : U'$ and $\Gamma' \vdash e' : U'$, then $\Gamma \vdash E[e'] : U$.*

Theorem 1 (Preservation).

If $\Gamma \mid \Sigma \vdash e : T$ and $\Gamma \mid \Sigma \vdash \mu$ and $e \mid \mu \longrightarrow e' \mid \mu'$, then $\Gamma \mid \Sigma' \vdash e' : T$ for some Σ' , where $\Sigma \subseteq \Sigma'$ and $\Gamma \mid \Sigma' \vdash \mu'$.

Note that the proof of the preservation theorem requires the match soundness property from proposition 1, in order to verify that after the reduction step of rule (REACT) (fig. 2), consumption of service requests and instantiation of rule bodies preserves the type of the enclosing parallel composition.

Progress. Our type system does not satisfy progress. For example, the following program is well-typed and not a value but cannot reduce:

$$(\text{spwn } (\text{srv } \text{foo}\langle \rangle \ \& \ \text{bar}\langle \rangle \triangleright \text{par } \varepsilon)) \# \text{foo}\langle \rangle.$$

The service request **foo** resolves fine, but the server's rule cannot fire because it is lacking a request **bar** joined with **foo**. Since our type system does not correlate service requests, it cannot guarantee that join patterns must succeed eventually. The integration of such a property is an interesting direction of future work, but orthogonal to the main contributions of this work.

Auxiliary Notation We adopt the following conventions. We omit the type bound if it is Top, e.g., $\Lambda \alpha <: \text{Top}. e$ becomes $\Lambda \alpha. e$. It is sometimes convenient to abbreviate longer type expressions. We introduce an abbreviation syntax for faux type constructors, e.g, $\text{SVC}[\alpha_1, \dots, \alpha_n] := \langle \alpha_1, \langle \alpha_2, \dots, \alpha_n \rangle \rangle$ defines a type SVC with n free type variables. Writing $\text{SVC}[T_1, \dots, T_n]$ denotes the type obtained by substituting the free occurrences of $\alpha_1, \dots, \alpha_n$ with the provided types T_1, \dots, T_n . Instead of repeating type annotations of service

parameters in every join pattern, we declare the service types once at the beginning of server templates. For example, $\text{srv } (a(x: \text{Int}) \triangleright \text{foo}()) (a(x: \text{Int}) \& b(y: \langle \text{Bool} \rangle) \triangleright \text{bar}())$ becomes

$$\text{srv } a: \langle \text{Int} \rangle, b: \langle \langle \text{Bool} \rangle \rangle \\ (a(x) \triangleright \text{foo}()) (a(x) \& b(x) \triangleright \text{bar}()).$$

We define function types as

$$T_1, \dots, T_n \rightarrow T := \langle T_1, \dots, T_n, \langle T \rangle \rangle$$

following our function encoding in Section 3.5. We define the union $(\text{srv } x: T) \cup (\text{srv } y: U)$ of two server-template types as the server template that contains all services of both types. The union is only defined if service names that occur in both types have identical type annotations – they are merely for syntactic convenience and do not represent real union types, which we leave for future work.

5. CPL at Work

We present two case studies to demonstrate the adequacy of CPL for solving the deployment issues identified in Section 2. The case studies will also be subsequently used to answer research questions about CPL’s features.

Firstly, we developed a number of reusable *server combinators*, expressing deployment patterns found in cloud computing. Our examples focus on load balancing and fault tolerance, demonstrating that programmers can define their own cloud services as strongly-typed, composable modules and address nonfunctional requirements with CPL. Secondly, we use our language to model MapReduce [18] deployments for distributed batch computations. Finally, we apply our server combinators to MapReduce, effortlessly obtaining a type-safe composition of services.

5.1 Server Combinators

In Section 2, we identified extensibility issues with deployment languages, which prevents programmers from integrating their own service implementations. We show how to implement custom service functionality with *server combinators* in a type-safe and composable way. Our combinators are similar in spirit to higher-order functions in functional programming.

As the basis for our combinators, we introduce **workers**, i.e., servers providing computational resources. A worker accepts work packages as **thunks**. Concretely, a worker models a managed virtual machine in a cloud and thunks model application services.

Following our derived syntax for thunks (Section 3.5), given an expression e of type α , the type of **think** e is:

$$\text{TThink}[\alpha] := \text{srv } \text{force}: \langle \alpha \rangle.$$

Service **force** accepts a continuation and calls it with the result of evaluating e . A worker accepts a thunk and executes it. At the type level, workers are values of a polymorphic type

$$\text{TWorker}[\alpha] := \text{srv } \text{init}: \langle \rangle, \text{work}: \text{TThink}[\alpha] \rightarrow \alpha.$$

```

1 MkWorker[α] = srv {
2   make: () → TWorker[α]
3   make(k) ▷
4   let worker = spwn srv {
5     init: ⟨⟩, work: TThink[α] → α
6     init() ▷ par ε //stub, do nothing
7     work(thnk, k) ▷ (spwn thnk)#force(k)
8   } in k(worker)
9 }

```

Figure 7. Basic worker factory.

That is, to execute a thunk on a worker, clients request the **work** service which maps the thunk to a result value. In addition, we allow workers to provide initialization logic via a service **init**. Clients of a worker should request **init** before they issue **work** requests. Figure 7 defines a factory for creating basic workers, which have no initialization logic and execute thunks in their own instance scope. In the following, we define server combinators that enrich workers with more advanced features.

To model **locality** – a worker uses its own computational resources to execute thunks – the spawn of a thunk should in fact *not* yield a new remote server instance. As discussed in Section 3.4, to keep the core language minimal the operational semantics does not distinguish whether a server is local or remote to another server. However, in our concurrent implementation of CPL, we allow users to annotate spawns as being remote or local, which enables us to model worker-local execution of thunks.

The combinators follow a common design principle. (i) The combinator is a factory for server templates, which is a server instance with a single **make** service. The service accepts one or more server templates which implement the **TWorker** interface, among possibly other arguments. (ii) Our combinators produce *proxy* workers. That is, the resulting workers implement the worker interface but forward requests of the **work** service to an internal instance of the argument worker.

5.1.1 Load Balancing

A common feature of cloud computing is on-demand scalability of services by dynamically acquiring server instances and distributing load among them. CPL supports the encoding of on-demand scalability in form of a server combinator, that distributes load over multiple workers dynamically, given a user-defined decision algorithm.

Dynamically distributing load requires a means to approximate worker utilization. Our first combinator **MkLoadAware** enriches workers with the ability to answer **getLoad** requests, which sends the current number of pend-

```

1 MkLoadAware[α, ω <: TWorker[α]] = srv {
2   make: ω → TWorker[α]
3   make(worker, k) ▷
4   let lWorker = srv {
5     instnc: ⟨inst ω⟩, getLoad: () → Int, load: ⟨Int⟩
6     work: TThunk[α] → α, init: ⟨⟩
7     //... initialization logic omitted
8
9     //forwarding logic for work
10    work(thnk, k) & instnc⟨w⟩ & load⟨n⟩ ▷
11    this#load⟨n+1⟩ || this#instnc⟨w⟩
12    || letk res = w#work(thnk)
13    in (k(res) || this#done⟨⟩)
14
15    //callback logic for fulfilled requests
16    done⟨⟩ & load⟨n⟩ ▷ this#load⟨n-1⟩
17
18    getLoad⟨k⟩ & load⟨n⟩ ▷ k⟨n⟩ || this#load⟨n⟩
19  } in k(lWorker)
20 }

```

Figure 8. Combinator for producing load-aware workers.

ing requests of the work service, our measure for utilization. Therefore, the corresponding type⁹ for load aware workers is

$$\text{TLAWorker}[\alpha] := \text{TWorker}[\alpha] \cup \text{srv } \text{getLoad}:\langle\langle\text{Int}\rangle\rangle.$$

The make service of the combinator accepts a server template *worker* implementing the TWorker interface and returns its enhanced version (bound to *lWorker*) back on the given continuation *k*. Lines 10-13 implement the core idea of forwarding and counting the pending requests. Continuation passing style enables us to intercept and hook on to the responses of *worker* after finishing work requests, which we express in Line 13 by the *letk* construct.

By building upon load-aware workers, we can define a polymorphic combinator *MkBalanced* that transparently introduces load balancing over a list of load-aware workers. The combinator is flexible in that it abstracts over the scheduling algorithm, which is an impure polymorphic function of type

$$\text{Choose}[\omega] := \text{List}[\text{inst } \omega] \rightarrow \text{Pair}[\text{inst } \omega, \text{List}[\text{inst } \omega]].$$

Given a (church-encoded) list of possible worker instances, such a function returns a (church-encoded) pair consisting of the chosen worker and an updated list of workers, allowing for dynamic adjustment of the available worker pool (*elastic load balancing*).

Figure 9 shows the full definition of the *MkBalanced* combinator. Similarly to Figure 8, the combinator is a factory which produces a decorated worker. The only difference being that now there is a list of possible workers to forward requests to. Choosing a worker is just a matter of querying the scheduling algorithm *choose* (Lines 15-16). Note that this combinator is only applicable to server templates imple-

```

1 MkBalanced[α, ω <: TWorker[α]] = srv {
2   make: (List[ω], Choose[ω]) → TWorker[α]
3   make(workers, choose, k) ▷
4   let lbWorker = srv {
5     insts: ⟨List[inst ω]⟩,
6     work: TThunk[α] → α, init: ⟨⟩
7
8     init⟨⟩ ▷ //spawn and init all child workers
9     letk spawned = mapk(workers, λw:ω. spwn w)
10    in (this#insts(spawned)
11    || foreach(spawned, λinst:inst ω. inst#init⟨⟩))
12
13    //forward to the next child worker
14    work(thnk, k) & insts⟨l⟩ ▷
15    letk (w, l') = choose⟨l⟩
16    in (w#work(thnk, k) || this#insts(l'))
17  } in k(lbWorker)
18 }

```

Figure 9. Combinator for producing load-balanced workers.

menting the TLAWorker[α] interface (Line 1), since *choose* should be able to base its decision on the current load of the workers.

In summary, mapping a list of workers with *MkLoadAware* and passing the result to *MkBalanced* yields a composite, load-balancing worker. It is thus easy to define hierarchies of load balancers programmatically by repeated use of the two combinators. Continuation passing style and the type system enable flexible, type-safe compositions of workers.

5.1.2 Failure Recovery

Cloud platforms monitor virtual machine instances to ensure their continual availability. We model failure recovery for crash/omission, permanent, fail-silent failures [27], where a failure makes a virtual machine unresponsive and is recovered by a restart.

Following the same design principles of the previous section, we can define a failure recovery combinator *MkRecover*, that produces fault-tolerant workers. Its definition is in the appendix of this report.

Self-recovering workers follow a basic protocol. Each time a work request is processed, we store the given thunk and continuation in a list until the underlying worker confirms the request's completion. If the wait time exceeds a timeout, we replace the worker with a fresh new instance and replay all pending requests. Crucial to this combinator is the *repl* syntactic form, which swaps the running server instance at the worker's address: *repl w (worker, ε)*. Resetting the worker's state amounts to setting the empty buffer *ε* in the server image value we pass to *repl*.

5.2 MapReduce

In this section, we illustrate how to implement the MapReduce [7] programming model with typed combinators in CPL, taking fault tolerance and load balancing into account.

⁹The union \cup on server types is for notational convenience at the meta level and *not* part of the type language.

```

1 MapReduce[κ1,ν1,κ2,ν2,ν3] = spwn srv {
2   make: (TMap[κ1,ν1,κ2,ν2],
3         TReduce[κ2,ν2,ν3],
4         TPartition[κ2],
5         ∀α.() → TWorker[α],
6         Int) → TMR[κ1,ν1,κ2,ν3]
7
8   make⟨Map, Reduce, Partition, R, mkWorker, k⟩ ▷
9   let sv = srv {
10    app⟨data, k0⟩ ▷ let
11     mworker =
12     mapValues(data, λv. mkWorker[List[Pair[κ1, ν2]]])
13     rworker =
14     mkMap(map(range(1, R), λi. (i, mkWorker[ν3]))))
15     grouper =
16     MkGrouper⟨Partition, R, Reduce
17              size(mworker), rworker, k0⟩
18     in foreach(data, λkey, val. {
19       let thnk = thunk Map(key, val)
20       in get(mworker, key)#work(thnk, grouper#group)}}
21   } in k(sv)
22 }

```

Figure 10. MapReduce factory.

MapReduce facilitates parallel data processing – cloud platforms are a desirable deployment target. The main point we want to make with this example is that CPL programs do not exhibit the unsafe composition, non-extensibility and staging problems we found in Section 2. Our design is inspired by Lämmel’s formal presentation in Haskell [18].

Figure 10 shows the main combinator for creating a MapReduce deployment, which is a first-class server. Following Lämmel’s presentation, the combinator is generic in the key and value types. κ_i denotes type parameters for keys and ν_i denotes type parameters for values.

The combinator takes as parameters the *Map* function for decomposing input key-value pairs into an intermediate list of intermediate pairs, the *Reduce* function for transforming grouped intermediate values into a final result value, the *Partition* function which controls grouping and distribution among reducers and the number R of reducers to allocate (Line 8). Parameter *mkWorker* is a polymorphic factory of type $\forall\alpha.() \rightarrow TWorker[\alpha]$. It produces worker instances for both the map and reduce stage.

Invoking *make* creates a new server template that on invocation of its *app* service deploys and executes a distributed MapReduce computation for a given set of (church-encoded) key-value pairs *data* and returns the result on continuation k_0 (Lines 9-10).

Firstly, workers for mapping and reducing are allocated and stored in the local map data structures *mworker* and *rworker*, where we assume appropriate cps-encoded functions that create and transform maps and sequences (Lines 11-14). Each key in the input *data* is assigned a new mapping worker and each partition from 1 to R is assigned a

```

1 MkGrouper[κ2,ν2,ν3] = spwn srv {
2   make⟨Partition: (κ2, Int) → Int, R: Int,
3         Reduce: TReduce[κ2,ν2,ν3],
4         R: Int,
5         rworker: Map[Int, TWorker[ν3]],
6         kr: ⟨Pair[κ2,ν3⟩],
7         k: ⟨inst srv (group: ⟨List[Pair[κ2,ν2]]⟩)⟩ ▷
8   let grpr = spwn srv {
9     //accumulated per-partition values for reduce
10    state: ⟨Map[Int, Map[κ2,ν2]]⟩,
11
12    //result callback invoked by mappers
13    group: ⟨List[Pair[κ2,ν2]]⟩,
14
15    //waiting state for phase one
16    await: ⟨Int⟩,
17
18    //trigger for phase two
19    done: ⟨Map[Int, Map[κ2,ν2]]⟩
20
21    //phase one: wait for mapper results
22    state⟨m⟩ & group⟨kvs⟩ & await⟨n⟩ ▷
23    letk m' = foldk⟨kvs, m,
24            λm'', kv. letk i = Partition(fst(kv), R)
25                      in updateGroup(m'', i, kv)⟩
26
27    in if (n > 0)
28       then (this#await⟨n - 1⟩ || this#state⟨m'⟩)
29       else this#done⟨m'⟩
30
31    //phase two: distribute to reducers
32    done⟨m⟩ ▷
33    foreach⟨m, λi. data2.
34      foreach⟨data2, λkey, vals.
35        let thnk = thunk Reduce⟨key, vals⟩
36        in get(rworker, i)#work(thnk, kr)⟩⟩
37  } in grpr#state(emptyMap) || grpr#await⟨R⟩ || k⟨grpr⟩

```

Figure 11. Grouper factory.

reducing worker. Additionally, a component for grouping and distribution among reducers (*grouper*) is allocated.

Secondly, the *foreach* invocation (Lines 18-20) distributes key-value pairs in parallel among mapping workers. For each pair, the corresponding worker should invoke the *Map* function, which we express as a *thunk* (Line 19, cf. section 3.5). All resulting intermediate values are forwarded to the grouper’s *group* service.

The grouper (Figure 11) consolidates multiple intermediate values with the same key and forwards them to the reducer workers. It operates in phases: (1) wait for all mapper workers to finish, meanwhile grouping incoming results (Lines 22-28) and (2), assign grouped results to reducer workers with the *Partition* function and distribute as *thunks*, which invoke the *Reduce* function (Lines 31-35). All reduction results are forwarded to the continuation k_r . For brevity we omit the final merge of the results.

Thanks to our service combinators, we can easily address non-functional requirements and non-intrusively add new features. The choice of the *mkWorker* parameter determines which variants of MapReduce deployments we obtain: The default variant just employs workers without advanced features, i.e.,

```
let make =  $\Lambda\alpha$ .(spwn MkWorker[ $\alpha$ ])#make
in MapReduce[ $\kappa_1, \nu_1, \kappa_2, \nu_2, \nu_3$ ]#make⟨f, r, p, R, make, k⟩
```

for appropriate choices of the other MapReduce parameters.

In order to obtain a variant, where worker nodes are elastically load-balanced, one replaces *make* with *makeLB* below, which composes the combinators from the previous section:

```
let choose = ...//load balancing algorithm
makeLB =  $\Lambda\alpha$ . $\lambda k$ . {
  letk w = (spwn MkWorker[ $\alpha$ ])#make⟨⟩
  lw = (spwn MkLoadAware[ $\alpha$ , TWorker[ $\alpha$ ]])#make⟨w⟩
  in (spwn MkBalanced[ $\alpha$ , TLAWorker[ $\alpha$ ]])
    #make⟨mkList(lw), choose, k⟩
}
```

in *makeLB*

A similar composition with the fault tolerance combinator yields fault tolerant MapReduce, where crashed mapper and reducer workers are automatically recovered.

5.3 Discussion

We discuss how CPL performed in the case studies answering the following research questions:

- Q1 (Safety): *Does CPL improve safety of cloud deployments?*
- Q2 (Extensibility): *Does CPL enable custom and extensible service implementations?*
- Q3 (Dynamic self-adjustment): *Does CPL improve flexibility in dynamic reconfiguration of deployments?*

Safety CPL is a strongly-typed language. As such, it provides internal safety (Section 2). The issue of cross-language safety (Section 2) does not occur in CPL programs, because configuration and deployment code are part of the same application. In addition, the interconnection of components is well-typed. For example, in the MapReduce case study, it is guaranteed that worker invocations cannot go wrong due to wrongly typed arguments. It is also guaranteed that workers yield values of the required types. As a result, all mapper and reducer workers are guaranteed to be compatible with the grouper component. In a traditional deployment program, interconnecting components amounts to referring to each others attributes, but due to the plain syntactic expansion, there is no guarantee of compatibility.

Extensibility The possibility to define combinators in CPL supports extensible, custom service implementations. At the type system level, bounded polymorphism and subtyping ensure that service implementations implement the required interfaces. The load balancing example enables nested load

balancing trees, since the combinator implements the well-known Composite design pattern from object-oriented programming. At the operational level, continuation passing style enables flexible composition of components, e.g., for stacking multiple features.

Dynamic Self-Adjustment In the case studies, we encountered the need of dynamically adapting the deployment configuration of an application, which is also known as “elasticity”. For example, the load balancer combinator can easily support dynamic growth or shrinkage of the list of available workers: New workers need to be dynamically deployed in new VMs (growth) and certain VMs must be halted and removed from the cloud configuration when the respective workers are not needed (shrinkage). Dynamic reconfiguration is not directly expressible in configuration languages, due to the two-phase staging. For example, configurations can refer to external elastic load balancer services provided by the cloud platform, but such services only provide a fixed set of balancing strategies, which may not suit the application. The load balancer service can be regarded as a black box, which happens to implement elasticity features. Also, a configuration language can request load balancing services only to the fixed set of machines which is specified in a configuration, but it is not possible if the number of machines is unknown before execution, as in the MapReduce case study. In contrast, CPL users can specify their own load balancing strategies and apply them programmatically.

5.4 Interfacing with Cloud Platforms

A practical implementation of CPL requires (1) a mapping of its concepts to real-world cloud platforms and (2) integrate existing cloud APIs and middleware services written in other languages. In the following, we sketch a viable solution; we leave a detailed implementation for future work.

For (1), CPL programs can be compiled to bytecode and be interpreted by a distributed run time hosted on multiple virtual machines.

Concerning (2), we envision our structural server types as the interface of CPL’s run time with the external world, i.e., pre-existing cloud services and artifacts written in other languages. CPL developers must write wrapper libraries to implement typed language bindings. Indeed, CPL’s first-class servers resemble (remote) objects, where services are their methods and requests are asynchronous method invocations (returning results on continuations). CPL implementations hence can learn from work on language bindings in existing object-oriented language run times, e.g., the Java ecosystem. To ensure type safety, dynamic type checking is necessary at the boundary between our run time and components written in dynamically or weakly typed languages.

Note that the representation of external services and artifacts as servers requires *immutable* addresses. That is, the run time should forbid `snap` and `repl` on such objects,

because it is in general impossible to reify a state snapshot of the external world.

For the primitives `spwn`, `snap`, and `repl`, the run time must be able to orchestrate the virtualization facilities of the cloud provider via APIs. Following our annotation-based approach to placement (Section 3.4), these primitives either map to local objects or to fresh virtual machines. Thus, invoking `spwn v` may create a new virtual machine hosting the CPL run time, which allocates and runs `v`. For local servers, `v` is executed by the run time that invoked `spwn`. One could extend the primitives to allow greater control of infrastructure-level concerns, such as machine configuration and geographic distribution. From these requirements and CPL’s design targeting extensible services and distributed applications, it follows that CPL is cross-cutting the three abstraction layers in contemporary cloud platforms: Infrastructure as a Service (IaaS), Platform as a Service (PaaS) and Software as a Service (SaaS) [30].

6. Related Work

Programming Models for Cloud Computing. The popularity of cloud computing infrastructures [30] has encouraged the investigation of programming models that can benefit from on-demand, scalable computational power and feature location transparency. Examples of these languages, often employed in the context of big data analysis, are Dryad [16], PigLatin [23] and FlumeJava [6]. These languages are motivated by refinements and generalizations of the original MapReduce [7] model.

Unlike CPL, these models specifically target only certain kinds of cloud computations, i.e., massive parallel computations and derivations thereof. They deliberately restrict the programming model to enable automated deployment, and do not address deployment programmability in the same language setting as CPL does. In this paper, we showed that the server abstraction of CPL can perfectly well model MapReduce computations in a highly parametric way, but it covers at the same time a more generic application programming model as well as deployment programmability. Especially, due to its join-based synchronization, CPL is well suited to serve as a core language for modeling cloud-managed stream processing.

Some researchers have investigated by means of formal methods specific computational models or specific aspects of cloud computing. The foundations in functional programming of MapReduce have been studied by Lämmel [18]. In CPL it is possible to encode higher-order functions and hence we can model MapReduce’s functionality running on a cloud computing platform. Jarraya et al. [17] extend the Ambient calculus to account for firewall rules and permissions to verify security properties of cloud platforms. To the best of our knowledge, no attempts have been done in formalizing cloud infrastructures in their generality.

Formal Calculi for Concurrent and Distributed Services. Milner’s CCS [20], the π calculus [21] and Hoare’s CSP [15] have been studied as the foundation of parallel execution and process synchronization.

Fournet’s and Gonthier’s Join Calculus [12] introduced join patterns for expressing the interaction among a set of processes that communicate by asynchronous message passing over communication channels. The model of communication channels in this calculus more adequately reflects communication primitives in real world computing systems which allows for a simpler implementation. In contrast, the notion of channel in the previously mentioned process calculi would require expensive global consensus protocols in implementations.

The design of CPL borrows join patterns from the Join Calculus. Channels in the Join Calculus are similar to services in CPL, but the Join Calculus does not have first-class and higher-order servers with qualified names. Also, there is no support for deployment abstractions.

The Ambient calculus [5] has been developed by Cardelli and Gordon to model concurrent systems that include both mobile devices and mobile computation. Ambients are a notion of named, bounded places where computations occur and can be moved as a whole to other places. Nested ambients model administrative domains and capabilities control access to ambients. CPL, in contrast, is location-transparent, which is faithful to the abstraction of a singular entity offered by cloud applications.

Languages for Parallel Execution and Process Synchronization. Several languages have been successfully developed/extended to support features studied in formal calculi.

JoCaml is an ML-like implementation of Join Calculus which adopts state machines to efficiently support join patterns [11]. Polyphonic C# [1] extends C# with join-like concurrency abstractions for asynchronous programming that are compiler-checked and optimized. Scala Joins [14] uses Scala’s extensible pattern matching to express joins. The Join Concurrency Library [26] is a more portable implementation of Polyphonic C# features by using C# 2.0 generics. JEScala [29] combines concurrency abstraction in the style of the Join Calculus with implicit invocation.

Funnel [22] uses the Join Calculus as its foundations and supports object-oriented programming with classes and inheritance. Finally, JErLang [25] extends the Erlang actor-based concurrency model. Channels are messages exchanged by actors, and received patterns are extended to express matching of multiple subsequent messages. Turon and Russo [28] propose an efficient, lock-free implementation of the join matching algorithm demonstrating that declarative specifications with joins can scale to complex coordination problems with good performance – even outperforming specialized algorithms. Fournet et al. [13] provide an implementation of the Ambient calculus. The implementation is obtained through a formally-proved translation to JoCaml.

CPL shares some features with these languages, basically those built on the Join Calculus. In principle, the discussion about the relation of CPL to Join Calculus applies to these languages as well, since the Join Calculus is their shared foundation. Implementations of CPL can benefit from the techniques developed in this class of works, especially [26].

7. Conclusions and Future Work

We presented CPL, a statically typed core language for defining asynchronous cloud services and their deployment on cloud platforms. CPL improves over the state of the art for cloud deployment DSLs: It enables (1) statically safe service composition, (2) custom implementations of cloud services that are composable and extensible and (3) dynamic changes to a deployed application. In future work, we will implement and expand core CPL to a practical programming language for cloud applications and deployment.

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B. Case Studies

In the following, we give the full definition of the server combinators and actor supervision case studies, which we omitted in section 5 due to space limitations.

B.1 Server Combinators for Cloud Computing

B.1.1 Failure Recovery

The combinator for failure recovery:

```

1 MkRecover[α,ω<:TWorker[α]] = spwn srv {
2   make: (ω, Int) → TWorker[α]
3
4   make⟨worker, timeout, k⟩ ▷
5   let selfrecovering = srv {
6     init: ⟨⟩,
7     work: TThunk[α] → α,
8     instnc: ⟨inst ω⟩
9     pending: ⟨List[(Int, Int, TThunk[α], ⟨α⟩)]⟩,
10    done: ⟨Int⟩
11
12    //initialization
13    init⟨⟩ ▷ let w = (spwn worker) in
14      w#init⟨⟩ || this#inst⟨w⟩ || this#pending⟨Nil⟩
15
16    //store and forward work requests to instnc
17    work(thnk, k) & instnc⟨w⟩ & pending⟨xs⟩ ▷
18    let ID = freshID() in
19      let now = localTime() in
20      this#pending⟨(ID, now, thnk, k) :: xs⟩
21      || this#instnc⟨w⟩
22      || (letk r = w#work(thnk)
23         in (k⟨r⟩ || this#done⟨ID⟩))
24
25    //work completion by instnc
26    done⟨ID⟩ & pending⟨xs⟩ ▷
27    filterk(xs, λp. fst(p) ≠ ID, this#pending))
28
29    //check for timeouts, restart instnc if needed
30    pending⟨xs⟩ & instnc⟨w⟩ ▷
31    let now = localTime() in
32    letk late = exists(xs,
33                      λp. now - snd(p) > timeout) in
34    if (late) then
35      (repl w (worker,ε); (w#init⟨⟩ || this#instnc⟨w⟩))
36      || foreach(xs, λp. this#work(thrd(p), frth(p)))
37    else
38      (this#pending⟨xs⟩ || this#instnc⟨w⟩)
39  } in k⟨selfrecovering⟩
40 }

```

Service `make` accepts a stoppable worker and an integer timeout parameter. The first rule of the self-recovering worker initializes the list of pending requests to the empty list `Nil`. The second rule accepts work requests. It generates a fresh ID for the request and adds the request to the list of pending requests together with the ID, the local timestamp and the continuation (we assume functions `freshID` and `localTime`). The rule forwards the work request to the

underlying worker and installs a continuation that notifies the proxy that the request completed using service `done`. The third rule accepts this request and removes the corresponding request from the list of pending requests.

Finally, the last rule checks if any of the pending requests has a timeout. If this happens, the rule replaces the old worker instance by a new one via `repl`, effectively resetting the state of the worker and re-initializing it. In parallel, all pending requests are replayed.

C. Type System Proofs

Definition 1. *The typed language extends evaluation contexts with type applications:*

$$E ::= \dots \mid E [T].$$

The reduction relation is extended by an additional contraction rule:

$$(\Lambda\alpha<:U. e) [T] \longrightarrow e \{\alpha := T\}. \text{ (TAPPABS)}$$

Definition 2. *We write $\Sigma \subseteq \Sigma'$ if for all $(i: T) \in \Sigma$, $(i: T) \in \Sigma'$ holds.*

Definition 3. *A routing table μ is well typed with respect to Γ, Σ (written $\Gamma \mid \Sigma \vdash \mu$), if $\text{dom}(\mu) = \text{dom}(\Sigma)$ and for all $i \in \text{dom}(\mu)$, $\Gamma \mid \Sigma \vdash \mu(i) : \Sigma(i)$ holds.*

Note. In the proofs we use the standard variable convention. That is, bound variables are assumed to be distinct and can be renamed if necessary so that no variable capture can occur in substitutions.

Lemma 6 (Substitution Lemma). *If $\Gamma, x: T_1 \mid \Sigma \vdash e_2 : T_2$ and $\Gamma \mid \Sigma \vdash e_1 : T_1$ then $\Gamma \mid \Sigma \vdash e_2\{x := e_1\} : T_2$.*

Proof. By induction on the typing derivation \mathcal{D} of $\Gamma, x: T_1 \mid \Sigma \vdash e_2 : T_2$. In each case we assume $\Gamma \mid \Sigma \vdash e_1 : T_1$.

Basis:

(T-VAR): Therefore $e_2 = y$ for $y \in \mathcal{N} \cup \{\mathbf{this}\}$ and $(\Gamma, x: T_1)(y) = T_2$. Case distinction:

$x = y$: Therefore $e_2 = x, T_2 = T_1$ and $e_2\{x := e_1\} = e_1$. From this and $\Gamma \mid \Sigma \vdash e_1 : T_1$ we obtain a derivation of $\Gamma \mid \Sigma \vdash e_2\{x := e_1\} : T_2$.

$x \neq y$: Therefore $e_2\{x := e_1\} = y\{x := e_1\} = y = e_2$, hence $\Gamma \mid \Sigma \vdash e_2\{x := e_1\} : T_2$, since the assumption $x: T_1$ can be dropped.

(T-INST): Immediate, since the context Γ is not considered in the premise.

(T-0): Immediate.

Inductive step:

Induction hypothesis (IH): The property holds for all proper subderivations of the derivation \mathcal{D} of $\Gamma, x: T_1 \mid \Sigma \vdash e_2 : T_2$.

(T-PAR): From the conclusion of the rule it holds that $e_2 = \mathbf{par} \overline{e_2}, T_2 = \mathbf{Unit}$ and from its premises $\Gamma, x: T_1 \mid$

$\Sigma \vdash e'_{2,i} : T_2$ for each $e'_{2,i}$ in the sequence $\overline{e'_2}$. Applying (IH) to each of the derivations in the premise yields $\Gamma \mid \Sigma \vdash e'_{2,i} \{x := e_1\} : T_2$ for each i . Together with rule (T-PAR) we obtain a derivation for $\Gamma \mid \Sigma \vdash \mathbf{par} \overline{e'_2} \{x := e_1\} : T_2$, which is also a derivation for $\Gamma \mid \Sigma \vdash e_2 \{x := e_1\} : T_2$ as desired, since $\mathbf{par} \overline{e'_2} \{x := e_1\} = (\mathbf{par} e'_2) \{x := e_1\} = e_2 \{x := e_1\}$.

(T-SRV): It holds that $e_2 = \mathbf{srv} \bar{r}$, $r_i = \bar{p}_i \triangleright e'_i$, $T_2 = \mathbf{srv} \overline{x_{i,j} : S_{i,j}}$, $\text{ftv}(T_2) \subseteq \text{ftv}(\Gamma, x : T_1)$, $p_{i,j} = \overline{x_{i,j} \langle y_{i,j} : T_{i,j} \rangle}$, $S_{i,j} = \langle T_{i,j} \rangle$ and $\Gamma, x : T_1, y_{i,j} : T_{i,j}, \mathbf{this} : T \mid \Sigma \vdash e'_i : \text{Unit}$ for each r_i in \bar{r} . Note that $\text{ftv}(T_2) \subseteq \text{ftv}(\Gamma)$, since $\text{ftv}(\Gamma, x : T_1) = \text{ftv}(\Gamma)$ by definition of ftv .

Case distinction:

$x = \mathbf{this}$: From the derivations of $\Gamma, x : T_1, \overline{y_{i,j} : T_{i,j}}$, $\mathbf{this} : T \mid \Sigma \vdash e'_i : \text{Unit}$ we obtain derivations for $\Gamma, \overline{y_{i,j} : T_{i,j}}, \mathbf{this} : T \mid \Sigma \vdash e'_i : \text{Unit}$ since the assumption $\mathbf{this} : T$ shadows $x : T_1$. Since server templates bind \mathbf{this} , it follows that $e_2 \{x := e_1\} = e_2$. Together with the other assumptions from the original derivation we obtain a derivation for $\Gamma \mid \Sigma \vdash e_2 \{x := e_1\} : T_2$ with rule (T-SRV) as desired.

$v \neq \mathbf{this}$: For each r_i in \bar{r} it holds that x distinct from $\overline{y_{i,j}}$ by the variable convention. From the derivation of $\Gamma, x : T_1, \overline{y_{i,j} : T_{i,j}}, \mathbf{this} : T \mid \Sigma \vdash e'_i : \text{Unit}$ we obtain by permutation a derivation of $\Gamma, \overline{y_{i,j} : T_{i,j}}, \mathbf{this} : T, x : T_1 \mid \Sigma \vdash e'_i : \text{Unit}$. With (IH) we obtain a derivation for $\Gamma, \overline{y_{i,j} : T_{i,j}}, \mathbf{this} : T \mid \Sigma \vdash e'_i \{x := e_1\} : \text{Unit}$.

From these intermediate derivations and the assumptions from the original (T-SRV) derivation we obtain by (T-SRV) a derivation of $\Gamma \mid \Sigma \vdash \mathbf{srv} (\bar{p}_i \triangleright e'_i \{x := e_1\}) : T_2$, which is also a derivation of $\Gamma \mid \Sigma \vdash e_2 \{x := e_1\} : T_2$ as desired.

(T-SPWN): Therefore $e_2 = \mathbf{spwn} e'_2$, $T_2 = \mathbf{inst} T'_2$, and $\Gamma, x : T_1 \mid \Sigma \vdash e'_2 : \mathbf{img} T'_2$. Applying (IH) yields $\Gamma \mid \Sigma \vdash e'_2 \{x := e_1\} : \mathbf{img} T'_2$. Together with rule (T-SPWN) we obtain a derivation of $\Gamma \mid \Sigma \vdash \mathbf{spwn} (e'_2 \{x := e_1\}) : T_2$, which is also a derivation of $\Gamma \mid \Sigma \vdash e_2 \{x := e_1\} : T_2$ as desired, since $\mathbf{spwn} (e'_2 \{x := e_1\}) = (\mathbf{spwn} e'_2) \{x := e_1\} = e_2 \{x := e_1\}$.

(T-SVC), (T-REQ), (T-IMG), (T-SNAP), (T-REPL) : Straightforward application of (IH) and substitution.

(T-TABS): Therefore $e_2 = \Lambda \alpha < : T'_2. e'_2$, $T_2 = \forall \alpha < : T'_2. T''_2$ and $\Gamma, x : T_1, \alpha < : T'_2 \mid \Sigma \vdash e'_2 : T''_2$. By the variable convention, it holds that α is not free in T_1 . Therefore, by permutation we obtain a derivation of $\Gamma, \alpha < : T'_2, x : T_1 \mid \Sigma \vdash e'_2 : T''_2$. Together with (IH) we obtain a derivation for $\Gamma, \alpha < : T'_2 \mid \Sigma \vdash e'_2 \{x := T_1\} : T''_2$. Extending this derivation with rule (T-TABS), we obtain a derivation of $\Gamma \mid \Sigma \vdash \Lambda \alpha < : T'_2. e'_2 \{x := T_1\} : T_2$, which is also a derivation of $\Gamma \mid \Sigma \vdash e_2 \{x := e_1\} : T_2$ as desired.

(T-TAPP): Therefore $e_2 = e'_2 [T'_2]$, $T_2 = T''_2 \{\alpha := T'_2\}$, $\text{ftv}(T'_2) \subseteq \text{ftv}(\Gamma)$, $\Gamma, x : T_1 \vdash T'_2 < : T''_2$ and $\Gamma, x : T_1 \mid \Sigma \vdash e'_2 : \forall \alpha < : T''_2. T''_2$. Applying (IH) yields a derivation of $\Gamma \mid \Sigma \vdash e'_2 \{x := e_1\} : \forall \alpha < : T''_2. T''_2$. Note that $\Gamma, x : T_1 \vdash T'_2 < : T''_2$ implies $\Gamma \vdash T'_2 < : T''_2$, since assumptions on variables do not play a role in subtyping rules. With rule (T-TAPP) we obtain a derivation of $\Gamma \mid \Sigma \vdash e'_2 \{x := e_1\} [T'_2] : T_2$, which is also a derivation of $\Gamma \mid \Sigma \vdash e_2 \{x := e_1\} : T_2$ as desired.

(T-SUB): By premise of the rule, $\Gamma, x : T_1 \mid \Sigma \vdash e_2 : T'_2$ and $\Gamma, x : T_1 \vdash T'_2 < : T_2$. The latter implies $\Gamma \vdash T'_2 < : T_2$, since assumptions on variables are not required in subtyping rules. Apply (IH) to obtain a derivation of $\Gamma \mid \Sigma \vdash e_2 \{x := e_1\} : T'_2$. Together with the previously established facts and (T-SUB) we obtain a derivation of $\Gamma \mid \Sigma \vdash e_2 \{x := e_1\} : T_2$ as desired. \square

Lemma 7 (Type Substitution Preserves Subtyping).

If $\Gamma, \alpha < : T', \Gamma' \vdash S < : T$ and $\Gamma \vdash S' < : T'$ then $\Gamma, \Gamma' \sigma \vdash S \sigma < : T \sigma$ where $\sigma = \{\alpha := S'\}$.

Proof. By induction on the typing derivation \mathcal{D} of $\Gamma, \alpha < : T', \Gamma' \vdash S < : T$. In each case we assume $\Gamma \vdash S' < : T'$ and $\sigma = \{\alpha := S'\}$.

Basis:

(S-TOP): Therefore $T = \text{Top}$. By rule (S-TOP) it holds that $\Gamma, \Gamma' \sigma \vdash S \sigma < : \text{Top}$, i.e., $\Gamma, \Gamma' \sigma \vdash S \sigma < : T \sigma$ as desired.

(S-REFL): Therefore $S = T$. $\Gamma, \Gamma' \sigma \vdash S \sigma < : T \sigma$ holds by rule (S-REFL).

(S-TVAR): Therefore $S = \alpha', \alpha' < : T \in (\Gamma, \alpha < : T', \Gamma')$. Case distinction:

$\alpha' \neq \alpha$: Immediate by rule (S-TVAR).

$\alpha' = \alpha$: Therefore $S = \alpha$, $S \sigma = T'$, and $T' = T = T \sigma$.

Apply rule (S-REFL).

(S-SRV \perp): Immediate by rule (S-SRV \perp).

Inductive step:

Induction hypothesis (IH): The property holds for all proper subderivations of the derivation $\Gamma, \alpha < : T', \Gamma' \vdash S < : T$.

(S-SRV): Therefore $S = \mathbf{srv} \overline{x : S_2}$, $T = \mathbf{srv} \overline{y : T_2}$ and for each j there is i such that $y_j = x_i$ and $\Gamma, \alpha < : T', \Gamma' \vdash S_{2,i} < : T_{2,j}$. Applying (IH) yields $\Gamma, \Gamma' \sigma \vdash S_{2,i} \sigma < : T_{2,j} \sigma$. Together with rule (S-SRV) we obtain a derivation of $\Gamma, \Gamma' \sigma \vdash \mathbf{srv} \overline{x : S_2 \sigma} < : \mathbf{srv} \overline{y : T_2 \sigma}$, i.e., $\Gamma, \Gamma' \sigma \vdash S \sigma < : T \sigma$ as desired.

(S-INST), (S-IMG), (S-SVC), (S-TRANS): Straightforward application of (IH).

(S-UNIV): Therefore $S = \forall \alpha_1 < : U. S_2$, $T = \forall \alpha_2 < : U. T_2$ and $\Gamma, \alpha < : T', \Gamma', \alpha_1 < : U \vdash S_2 < : T_2 \{\alpha_2 := \alpha_1\}$. Together with (IH) we obtain a derivation of $\Gamma, \Gamma' \sigma, \alpha_1 < : U \sigma \vdash S_2 \sigma < : T_2 \{\alpha_2 := \alpha_1\} \sigma$, i.e., $\Gamma, \Gamma' \sigma, \alpha_1 < : U \sigma \vdash S_2 \sigma < :$

$T_2\sigma \{\alpha_2 := \alpha_1\}$ (since by our variable convention, we may assume $\alpha_2 \neq \alpha$ and $\alpha_1 \neq \alpha$). Together with rule (S-UNIV) we obtain a derivation of $\Gamma, \Gamma'\sigma \vdash \forall \alpha_1 <: U\sigma. S_2\sigma <: \forall \alpha_2 <: U\sigma. T_2\sigma$, i.e., $\Gamma, \Gamma'\sigma \vdash S\sigma <: T\sigma$ as desired. \square

Lemma 8 (Type Substitution Lemma). *If $\Gamma, \alpha <: S, \Gamma' \mid \Sigma \vdash e : T$ and $\Gamma \vdash S' <: S$ then $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash e\sigma : T\sigma$ where $\sigma = \{\alpha := S'\}$.*

Proof. By induction on the typing derivation \mathcal{D} of $\Gamma, \alpha <: S, \Gamma' \mid \Sigma \vdash e : T$. In each case we assume $\Gamma \vdash S' <: S$ and $\sigma = \{\alpha := S'\}$.

Basis:

(T-VAR): Therefore $e = x$ for $x \in \mathcal{N} \cup \{\mathbf{this}\}$ and $(\Gamma, \alpha <: S, \Gamma')(x) = T$. By the variable convention, it holds that α is not bound in Γ , therefore $(\Gamma, \Gamma'\sigma)(x) = T\sigma$, which holds by a structural induction on Γ' . Together with (T-VAR) we obtain $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash e\sigma : T\sigma$ as desired.

(T-INST): Immediate, since Γ is not considered in the premises.

(T-0): Immediate by rule (T-0).

Inductive step:

Induction hypothesis (IH): The property holds for all proper subderivations of the derivation \mathcal{D} of $\Gamma, \alpha <: S, \Gamma' \mid \Sigma \vdash e : T$.

(T-PAR): From the conclusion of the rule it holds that $e = \mathbf{par} \bar{e}_2, T = \mathbf{Unit}$ and from its premises $\Gamma, \alpha <: S, \Gamma' \mid \Sigma \vdash e_{2,i} : T$ for each $e_{2,i}$ in the sequence \bar{e}_2 . Applying (IH) yields $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash e_{2,i}\sigma : T\sigma$ for each i . Together with $T\sigma = \mathbf{Unit}\sigma = \mathbf{Unit} = T$, by (T-PAR) we obtain a derivation of $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash \mathbf{par} \bar{e}_2\sigma : T$, which is also a derivation of $\Gamma, \Gamma' \mid \Sigma\sigma \vdash e\sigma : T\sigma$ as desired.

(T-SRV): Therefore $e = \mathbf{srv} \bar{r}, r_i = \bar{p}_i \triangleright e_i, T = \mathbf{srv} \overline{x_{i,j} \langle y_{i,j} : \bar{T}_{i,j} \rangle}, \text{ftv}(T) \subseteq \text{ftv}(\Gamma, \alpha <: S, \Gamma'), p_{i,j} = x_{i,j} \langle y_{i,j} : \bar{T}_{i,j} \rangle, S_{i,j} = \langle \bar{T}_{i,j} \rangle$ and $\Gamma, \alpha <: S, \Gamma', y_{i,j} : \bar{T}_{i,j}, \mathbf{this} : T \mid \Sigma \vdash e_i : \mathbf{Unit}$ for each r_i in \bar{r} . Applying (IH) yields $\Gamma, \Gamma'\sigma, y_{i,j} : \bar{T}_{i,j}\sigma, \mathbf{this} : T\sigma \mid \Sigma\sigma \vdash e_i\sigma : \mathbf{Unit}$ for each i in \bar{r} . Note that $\text{ftv}(T) \subseteq \text{ftv}(\Gamma, \alpha <: S, \Gamma'), \Gamma \vdash S' <: S$ and $\sigma = \{\alpha := S'\}$ imply $\text{ftv}(T\sigma) \subseteq \text{ftv}(\Gamma, \Gamma'\sigma)$. By applying σ to the types in the assumptions of the original derivation \mathcal{D} , we obtain with the previously established facts a derivation of $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash e\sigma : T\sigma$ as desired by rule (T-SRV).

(T-IMG), (T-SNAP), (T-REPL), (T-SPWN), (T-SVC), (T-REQ), (T-TABS): obtain a derivation of $\Gamma \mid \Sigma' \vdash \mathbf{par} \bar{e}_{11} \bar{e}_{12} \bar{e}_{13} : T$ by straightforward application of (IH) and substitution.

(T-TAPP): Therefore $e = e_2 [T_2], T = T' \{\alpha' := T_2\}$, $\text{ftv}(T_2) \subseteq \text{ftv}(\Gamma \alpha <: S, \Gamma'), \Gamma, \alpha <: S, \Gamma' \vdash T_2 <: T_3$ and $\Gamma, \alpha <: S, \Gamma' \mid \Sigma \vdash e_2 : \forall \alpha' <: T_3. T'$. Applying (IH) yields $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash e_2\sigma : (\forall \alpha' <: T_3. T')\sigma$, hence $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash e_2\sigma : \forall \alpha' <: T_3\sigma. T'\sigma$. By lemma 7 and

$\Gamma, \alpha <: S, \Gamma' \vdash T_2 <: T_3$ it holds that $\Gamma, \Gamma'\sigma \vdash T_2\sigma <: T_3\sigma$. From $\text{ftv}(T_2) \subseteq \text{ftv}(\Gamma, \alpha <: S, \Gamma')$ it holds that $\text{ftv}(T_2\sigma) \subseteq \text{ftv}(\Gamma, \Gamma'\sigma)$. Applying rule (T-TAPP) yields a derivation of $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash e_2\sigma [T_2\sigma] : (T'\sigma) \{\alpha' := T_2\sigma\}$, i.e., $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash (e_2 [T_2])\sigma : (T' \{\alpha' := T_2\})\sigma$, i.e., $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash e\sigma : T\sigma$ as desired.

(T-SUB): By premise of the rule, $\Gamma, \alpha <: S, \Gamma' \mid \Sigma \vdash e : T'$ and $\Gamma, \alpha <: S, \Gamma' \vdash T' <: T$. Applying (IH) to the former yields $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash e\sigma : T'\sigma$. By lemma 7 and $\Gamma, \alpha <: S, \Gamma' \vdash T' <: T$ it holds that $\Gamma, \Gamma'\sigma \vdash T'\sigma <: T\sigma$. Thus, by rule (T-SUB) we obtain a derivation of $\Gamma, \Gamma'\sigma \mid \Sigma\sigma \vdash e\sigma : T\sigma$ as desired. \square

Lemma 9 (Location Typing Extension Preserves Types).

If $\Gamma \mid \Sigma \vdash e : T$ and $\Sigma \subseteq \Sigma'$, then $\Gamma \mid \Sigma' \vdash e : T$.

Proof. Straightforward induction on the derivation for $\Gamma \mid \Sigma \vdash e : T$. \square

Theorem 2 (Preservation). *If $\Gamma \mid \Sigma \vdash e : T$ and $\Gamma \mid \Sigma \vdash \mu$ and $e \mid \mu \longrightarrow e' \mid \mu'$, then $\Gamma \mid \Sigma' \vdash e' : T$ for some Σ' , where $\Sigma \subseteq \Sigma'$ and $\Gamma \mid \Sigma' \vdash \mu'$.*

Proof. By induction on the typing derivation \mathcal{D} of $\Gamma \mid \Sigma \vdash e : T$. We always assume $e \mid \mu \longrightarrow e' \mid \mu'$ for some e', μ' . Otherwise, e is stuck ($e \not\rightarrow$) and the property trivially holds. We also assume $\Gamma \mid \Sigma \vdash \mu$ in each case.

Basis:

(T-VAR), (T-INST), (T-0): Immediate, since e is stuck.

Inductive step:

Induction hypothesis (IH): The property holds for all proper subderivations of the derivation \mathcal{D} of $\Gamma \mid \Sigma \vdash e : T$.

(T-SRV), (T-IMG), (T-TABS): The property trivially holds, since in each case, e is stuck.

(T-PAR): From the conclusion of the rule it holds that $e = \mathbf{par} \bar{e}_1, T = \mathbf{Unit}$ and from its premises $\Gamma \mid \Sigma \vdash e_{1,i} : \mathbf{Unit}$ for each $e_{1,i}$ in the sequence \bar{e}_1 . By the structure of e , there are three possible rules which can be at the root of the derivation for $e \mid \mu \longrightarrow e' \mid \mu'$:

(PAR): Therefore $e = \mathbf{par} \bar{e}_{11} (\mathbf{par} \bar{e}_{12}) \bar{e}_{13}$ and $e' = \mathbf{par} \bar{e}_{11} \bar{e}_{12} \bar{e}_{13}$ and $\mu' = \mu$. From the premises of (T-PAR) it holds that $\Gamma \mid \Sigma \vdash e_{12,k} : \mathbf{Unit}$ for each $e_{12,k}$ in the sequence \bar{e}_{12} . Choose $\Sigma' = \Sigma$. Together with the previously established facts we obtain a derivation of $\Gamma \mid \Sigma' \vdash \mathbf{par} \bar{e}_{11} \bar{e}_{12} \bar{e}_{13} : T$ by (T-PAR), $\Sigma \subseteq \Sigma'$ and $\Gamma \mid \Sigma' \vdash \mu'$ as desired.

(REACT): Therefore $e = \mathbf{par} e'', e' = \mathbf{par} e'' \sigma_b(e_b)$, $\mu' = \mu[i \mapsto (s, \bar{m}')]]$. By the premises of (REACT), $\mu(i) = (\mathbf{srv} \bar{r}_1 (\bar{p} \triangleright e_b) \bar{r}_2, \bar{m})$, $\text{match}(\bar{p}, \bar{m}) \Downarrow (\bar{m}', \sigma)$ and $\sigma_b = \sigma \cup \{\mathbf{this} := i\}$. Choose $\Sigma' = \Sigma$. Since $\Gamma \mid \Sigma \vdash \mu, \mu(i)$ is well typed. From its shape

it is typed by rule (T-IMG) as some $\mathbf{img} T'$. Thus, by the premises of (T-IMG), s is typed as $\mathbf{srv} T'$ and each element in the buffer \overline{m} is a valid request value for the server template s . By the match soundness and completeness lemma from the paper, each request value in the buffer \overline{m}' occurs in \overline{m} . Hence, we obtain a derivation for $\Gamma, \Sigma \vdash (s, \overline{m}') : \mathbf{img} T'$. Thus, $\Gamma \mid \Sigma \vdash \mu'$ and also $\Gamma \mid \Sigma' \vdash \mu'$.

From the shape of server template s , it must be typed by rule (T-SRV) as the last step in a derivation D_s . Therefore, from the premises of this rule, we obtain a derivation for $\Gamma, \overline{y_{l,k}} : \overline{T_{l,k}}, \mathbf{this} : \mathbf{srv} T' \mid \Sigma \vdash e_b : \mathbf{Unit}$, i.e., $\Gamma, \overline{y_{l,k}} : \overline{T_{l,k}}, \mathbf{this} : \mathbf{srv} T' \mid \Sigma' \vdash e_b : \mathbf{Unit}$, since $\Sigma' = \Sigma$. The $\overline{y_{l,k}} : \overline{T_{l,k}}$ are the arguments in the join pattern \overline{p} . Applying the substitution lemma 6 multiple times to the latter derivation yields a derivation of $\Gamma \mid \Sigma' \vdash e_b : \mathbf{Unit}$. $\Gamma \mid \Sigma' \vdash \sigma_b(e_b) : \mathbf{Unit}$. The first application of the lemma to eliminate $\mathbf{this} : \mathbf{srv} T'$ is justified by the derivation D_s . The other applications to eliminate $\overline{y_{l,k}} : \overline{T_{l,k}}$ are justified by the match soundness and completeness lemma, which guarantees that the selection of argument values in the substitution σ are from matching service request values in \overline{m} , which is well-typed under $\Gamma \mid \Sigma'$. Hence $\Gamma \mid \Sigma' \vdash \sigma(y_{l,k}) : \overline{T_{l,k}}$ holds.

Finally, since $\Sigma' = \Sigma$, $\Gamma \mid \Sigma \vdash e'' : \mathbf{Unit}$, we have $\Gamma \mid \Sigma' \vdash e'' : \mathbf{Unit}$. Together with $\Gamma \mid \Sigma' \vdash e_b : \mathbf{Unit}$ by rule (T-PAR), we obtain $\Gamma \mid \Sigma' \vdash \mathbf{par} e'' e_b : \mathbf{Unit}$, i.e., $\Gamma \mid \Sigma' \vdash e' : T$. This together with the previously established $\Gamma \mid \Sigma' \vdash \mu'$ is the property we wanted to show.

(CONG): Therefore $e = E[e_{1,j}]$ for an expression $e_{1,j}$ in the sequence $\overline{e_1}$ and $e' = E[e'_{1,j}]$ for some $e'_{1,j}$, where $e_{1,j} \mid \mu \longrightarrow e'_{1,j} \mid \mu'$. Since $\Gamma \mid \Sigma \vdash e_{1,j} : \mathbf{Unit}$, it follows from (IH) that there is a derivation of $\Gamma \mid \Sigma' \vdash e'_{1,j} : \mathbf{Unit}$ with $\Sigma \subseteq \Sigma'$ and $\Gamma \mid \Sigma \vdash \mu'$. From $\Gamma \mid \Sigma \vdash e_{1,i} : \mathbf{Unit}$ for each $i \neq j$, $\Sigma \subseteq \Sigma'$ and lemma 9, we obtain derivations $\Gamma \mid \Sigma' \vdash e_{1,i} : \mathbf{Unit}$. Together with $\Gamma \mid \Sigma' \vdash e'_{1,j} : \mathbf{Unit}$ we obtain a derivation for $\Gamma \mid \Sigma' \vdash e' : T$ by rule (T-PAR) as desired.

(T-SNAP): Therefore $e = \mathbf{snap} e_1$, $T = \mathbf{img} T'$ and $\Gamma \mid \Sigma \vdash e_1 : \mathbf{inst} T'$. By the structure of e , there are two possible rules which can be at the root of the derivation for $e \mid \mu \longrightarrow e' \mid \mu'$:

(SNAP): Therefore, $e_1 = i \in \mathbb{N}$, $e' = (\mathbf{srv} \overline{r}, \overline{m})$ or $e' = \mathbf{0}$, and $\mu' = \mu$. From $\Gamma \mid \Sigma \vdash e_1 : \mathbf{inst} T'$ and the shape of e_1 , rule (T-INST) is the root of the corresponding derivation. Thus, $i \in \Sigma$ and $\Sigma(i) = \mathbf{inst} T'$ by the premises of this rule. Choose $\Sigma' = \Sigma$. Since $\Gamma \mid \Sigma \vdash \mu$, $\mu = \mu'$ and $\Sigma' = \Sigma$, it also holds that $\Gamma \mid \Sigma' \vdash \mu'$. Hence $\Gamma \mid \Sigma' \vdash e' : T$ as desired.

(CONG): Therefore, $e' = \mathbf{snap} e_2$ for some e_2 , and $e_1 \mid \mu \longrightarrow e_2 \mid \mu'$ holds. Applying this together with $\Gamma \mid \Sigma \vdash e_1 : \mathbf{inst} T'$ to (IH) yields $\Gamma \mid \Sigma'' \vdash e_2 : \mathbf{inst} T'$ for Σ'' , where $\Sigma \subseteq \Sigma''$ and $\Gamma \mid \Sigma'' \vdash \mu'$. Together with rule (T-SNAP) we obtain $\Gamma \mid \Sigma'' \vdash \mathbf{snap} e_2 : \mathbf{img} T'$, i.e., $\Gamma \mid \Sigma'' \vdash e' : T$. Choose $\Sigma' = \Sigma''$.

(T-REPL): Therefore $e = \mathbf{repl} e_1 e_2$, $T = \mathbf{Unit}$, $\Gamma \mid \Sigma \vdash e_1 : \mathbf{inst} T'$ and $\Gamma \mid \Sigma \vdash e_2 : \mathbf{img} T'$. By the structure of e , there are two possible rules which can be at the root of the derivation for $e \mid \mu \longrightarrow e' \mid \mu'$:

(REPL): Therefore, $e_1 = i \in \mathbb{N}$, $i \in \text{dom}(\mu)$, $e_2 = (\mathbf{srv} \overline{r}, \overline{m})$ or $e_2 = \mathbf{0}$, $e' = \mathbf{par} \varepsilon$ and $\mu' = mu[i \mapsto s]$. Hence i is well typed under $\Gamma \mid \Sigma$ as $\mathbf{inst} T'$. Together with $\Gamma \mid \Sigma \vdash e_2 : \mathbf{img} T'$ and $\Gamma \mid \Sigma \vdash \mu$ it holds that $\Gamma \mid \Sigma \vdash \mu'$. Choose $\Sigma' = \Sigma$. Apply rule (T-PAR) to obtain $\Gamma \mid \Sigma' \vdash e' : T$ as desired.

(CONG): Therefore, $e' = \mathbf{repl} e'_1 e'_2$ for some e'_1, e'_2 and either $e_1 \mid \mu \longrightarrow e'_1 \mid \mu'$, $e_2 = e'_2$ or $e_2 \mid \mu \longrightarrow e'_2 \mid \mu'$, $e_1 = e'_1$ holds. We only show the first case, the other is similar. Apply (IH) to $\Gamma \mid \Sigma \vdash e_1 : \mathbf{inst} T'$ and $e_1 \mid \mu \longrightarrow e'_1 \mid \mu'$ to obtain Σ'' with $\Sigma \subseteq \Sigma''$ and $\Gamma \mid \Sigma'' \vdash \mu'$ and $\Gamma \mid \Sigma'' \vdash e'_1 : \mathbf{inst} T'$. Choose $\Sigma' = \Sigma''$. Apply lemma 9 to $\Gamma \mid \Sigma \vdash e_2 : \mathbf{img} T'$ in order to obtain $\Gamma \mid \Sigma' \vdash e_2 : \mathbf{img} T'$. Finally, apply rule (T-REPL) to obtain $\Gamma \mid \Sigma' \vdash e' : T$ as desired.

(T-SPWN): Therefore $e = \mathbf{spwn} e''$, $T = \mathbf{inst} T'$ and $\Gamma \mid \Sigma \vdash e'' : \mathbf{img} T'$. By the structure of e , there are two possible rules which can be at the root of the derivation of $e \longrightarrow e'$:

(SPWN): Therefore $e'' = (\mathbf{srv} \overline{r}, \overline{m})$ or $e'' = \mathbf{0}$, $e' = i \in \mathbb{N}$, $i \notin \text{dom}(\mu)$ and $\mu' = \mu[i \mapsto e'']$. With $\Gamma \mid \Sigma \vdash \mu$ and definition 3 it follows that $i \notin \text{dom}(\Sigma)$. Choose $\Sigma' = \Sigma[i \mapsto \mathbf{img} T']$. By rule (T-INST) and definition of Σ' , it holds that $\Gamma \mid \Sigma' \vdash i : \mathbf{inst} T'$, i.e., $\Gamma \mid \Sigma' \vdash e' : T$. By construction, $\Sigma \subseteq \Sigma'$.

What is left to show is $\Gamma \mid \Sigma' \vdash \mu'$:

First note $\text{dom}(\mu) = \text{dom}(\Sigma)$ and for all $j \in \text{dom}(\Sigma)$, $\Sigma(j) = \Sigma'(j)$, $\mu(j) = \mu'(j)$ and $\Gamma \mid \Sigma \vdash \mu(j) : \Sigma(j)$. Hence $\Gamma \mid \Sigma \vdash \mu'(j) : \Sigma'(j)$ and by lemma 9, $\Gamma \mid \Sigma' \vdash \mu'(j) : \Sigma'(j)$ for each $j \in \text{dom}(\Sigma)$.

From $\Gamma \mid \Sigma \vdash e'' : \mathbf{img} T'$, $\Sigma \subseteq \Sigma'$ and lemma 9 we obtain $\Gamma \mid \Sigma' \vdash e'' : \mathbf{img} T'$. Together with the definitions of Σ' , μ' , this is a derivation for $\Gamma \mid \Sigma' \vdash \mu'(i) : \Sigma'(i)$.

In summary, we have established that $\Gamma \mid \Sigma' \vdash \mu'(j) : \Sigma'(j)$ for all $j \in \text{dom}(\mu) \cup \{i\} = \text{dom}(\mu') = \text{dom}(\Sigma')$. By definition 3, this means $\Gamma \mid \Sigma' \vdash \mu'$, what was left to show.

(CONG): Therefore, by the structure of e , it holds that $e = E[e'']$ for the context $E[\cdot] = \mathbf{spwn} [\cdot]$. By the premise of (CONG) we obtain $e'' \mid \mu \longrightarrow e''' \mid \mu'$, hence $e' = E[e'''] = \mathbf{spwn} e'''$. Applying the (IH) to $\Gamma \mid \Sigma \vdash e'' : \mathbf{img} T'$ and $e'' \mid \mu \longrightarrow e''' \mid \mu'$ yields a derivation of $\Gamma \mid \Sigma' \vdash e''' : \mathbf{img} T'$ for some Σ' with

$\Sigma \subseteq \Sigma'$ and $\Gamma \mid \Sigma' \vdash \mu'$. From the previous typing derivation and rule (T-SPWN) we obtain a derivation for $\Gamma \mid \Sigma' \vdash \text{spwn } e''' : \text{inst } T'$, i.e., $\Gamma \mid \Sigma' \vdash e' : T$ as desired.

(T-SVC): Therefore $e = \overline{e'' \sharp x_i}$, $T = T_{1,i}$, $\Gamma \mid \Sigma \vdash e'' : \text{inst srv } \overline{x : T_1}$, where $x_i : T_{1,i}$ occurs in the sequence $\overline{x : T_1}$. Since $e \mid \mu \longrightarrow e' \mid \mu'$ by assumption, e'' cannot be a value, otherwise e too would be a value and hence stuck. Together with the structure of e , reduction rule (CONG) is the only possible root of the derivation of $e \mid \mu \longrightarrow e' \mid \mu'$, where $e = E[e'']$. Hence $e' = E[e'''] = \overline{e''' \sharp x_i}$ and $e'' \mid \mu \longrightarrow e''' \mid \mu'$ by the premise of (CONG). Applying the (IH) to $e'' \mid \mu \longrightarrow e''' \mid \mu'$ and $\Gamma \mid \Sigma \vdash e'' : \text{inst srv } \overline{x : T_1}$ yields a derivation of $\Gamma \mid \Sigma' \vdash e''' : \text{inst srv } \overline{x : T_1}$, for some Σ' with $\Sigma \subseteq \Sigma'$ and $\Gamma \mid \Sigma' \vdash \mu'$. By rule (T-SVC) and the previously established facts, we obtain a derivation for $\Gamma \mid \Sigma' \vdash \overline{e''' \sharp x_i} : T_{1,i}$, which is also a derivation of $\Gamma \mid \Sigma' \vdash e' : T$ as desired.

(T-REQ): Therefore $e = e'' \langle e_1 \dots e_n \rangle$, $T = \text{Unit}$, $\Gamma \vdash e'' : \langle T_1 \dots T_n \rangle$ and $(\Gamma \vdash e_i : T_i)_{i \in 1 \dots n}$. Since $e \longrightarrow e'$ by assumption, there is an expression in the set $\{e'', e_1, \dots, e_n\}$ which is not a value, otherwise e is a value and stuck. Together with the structure of e , reduction rule (CONG) is the only possible root of the derivation of $e \longrightarrow e'$. Therefore $e = E[e''']$, where

$E[\cdot] = [\cdot] \langle e_1 \dots e_n \rangle$ or $E[\cdot] = e'' \langle \overline{e_{11}} [\cdot] \overline{e_{22}} \rangle$. For any of the possible shapes of E , we can straightforwardly apply the (IH) to obtain a derivation of $\Gamma \vdash e' : T$ as desired.

(T-TAPP): Therefore $e = e'' [T_1]$, $T = T' \{\alpha := T_1\}$, $\text{ftv}(T_1) \subseteq \text{ftv}(\Gamma)$, $\Gamma \vdash T_1 < T_2$ and $\Gamma \mid \Sigma \vdash e'' : \forall \alpha < T_2. T'$. By the structure of e , there are two possible rules which can be at the root of the derivation of $e \mid \mu \longrightarrow e' \mid \mu'$:

(TAPPABS): Therefore $e'' = \Lambda \alpha < T_2. e'''$ and hence $e' = e''' \{\alpha := T_1\}$. From the structure of e , e'' and the available rules, there is a proper subderivation in \mathcal{D} of $\Gamma, \alpha < T_2 \mid \Sigma \vdash e''' : T'$. Together with $\Gamma \vdash T_1 < T_2$ and the type substitution lemma 8, we obtain a derivation for $\Gamma \mid \Sigma \vdash e''' \{\alpha := T_1\} : T' \{\alpha := T_1\}$. Choose $\Sigma' = \Sigma$, then the previous derivation also is a derivation of $\Gamma \mid \Sigma' \vdash e' : T$ as desired.

(CONG): Straightforward application of the (IH) similar to the previous cases.

(T-SUB): By premise of the rule, $\Gamma \mid \Sigma \vdash e : T'$ and $\Gamma \vdash T' < T$. Apply (IH) to the former and then (T-SUB) to obtain a derivation of $\Gamma \mid \Sigma' \vdash e' : T$ and an appropriate Σ' .

□